Robust Constant-Time Cryptography

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Cryptographic *library* developers take care to ensure their library does not leak secrets even when there are (inevitably) exploitable vulnerabilities in the *applications* the library is linked against. To do so, they choose some class of application vulnerabilities to defend against and hardcode protections against those vulnerabilities in the library code. A single set of choices is a poor fit for all contexts: a chosen protection could impose unnecessary overheads in contexts where those attacks are impossible, and an ignored protection could render the library insecure in contexts where the attack is feasible.

We introduce RoboCop, a new methodology and toolchain for building secure *and* efficient applications from cryptographic libraries, via four contributions. First, we present an operational semantics that describes the behavior of a (cryptographic) library executing in the context of a potentially vulnerable application so that we can precisely specify what different attackers can observe. Second, we use our semantics to define a novel security property, *Robust Constant Time* (RCT), that defines when a cryptographic library is secure *in the context of* a vulnerable application. Crucially, our definition is parameterized by an attacker model, allowing us to factor out the classes of attackers that a library may wish to secure against. This refactoring yields our third contribution: a compiler that can synthesize bespoke cryptographic libraries with security tailored to the specific application context against which the library will be linked, guaranteeing that the library is RCT in that context. Finally, we present an empirical evaluation that shows the RoboCop compiler can automatically generate code to efficiently protect a wide range (over 540) of cryptographic library primitives against three classes of attacks: read gadgets (due to application memory safety vulnerabilities), speculative read gadgets (due to application speculative execution vulnerabilities), and concurrent observations (due to application threads), with performance overhead generally under 2% for protections from read gadgets and under 4% for protections from speculative read gadgets, thus freeing library developers from making one-size-fits-all choices between security and performance.

1 Introduction

"Don't roll your own crypto" is a well known adage directed at application developers when they are considering using cryptography in their code. Instead developers are exhorted to use cryptographic libraries written by experts who have hopefully learned the hard-won lessons of decades of cryptographic and software security research and practice. For such cryptographic library developers, as well as algorithm designers, one of those hard-won lessons is that cryptographic code must be *constant time* [9]. More recently (due to microarchitectural attacks like Spectre [31]) this lesson has been expanded to the requirement that, under certain circumstances, it is also crucial for cryptographic code to be *speculatively constant time* [14]. Together these ensure that the code in the library does not leak secrets (such as cryptographic keys) through either timing channels or speculative execution, respectively, and significant work has gone into developing theory, tools, and rules of thumb to ensure that cryptographic code is (speculatively) constant time [6, 15].

Attacks via application vulnerabilities. Sadly, this is not enough. While a constant time cryptographic library will not *itself* leak secrets, it is but one component executing within the context of a larger *application*. Security vulnerabilities in application code could themselves lead to inadvertent disclosure of secrets, no matter how careful library authors were to avoid vulnerabilities.

```
1 int stream(u8 *c, u64 clen, u8 *n, u8 *k) {
2     ... u8 kcopy[32]; ...
3     for (i = 0; i < 32; i++) { kcopy[i] = k[i]; }
4     ...
5     while (clen >= 64) { crypto_core_salsa20(c, in, kcopy, NULL); ... }
6     ...
7     sodium_memzero(kcopy, sizeof kcopy);
8     return 0;
9 }
```

Fig. 1. An excerpt from the reference implementation of Salsa20 in LibSodium.

Library authors are keenly aware of this problem and routinely add protections to harden their code against application vulnerabilities. For example, consider the excerpt of the implementation of the `Salsa20` stream cipher [10] from `LibSodium` [18] shown in Figure 1. We might hope that the (elided) body being constant time suffices to ensure that the secrets in `kcopy` are not leaked. However the *classic* constant time guarantee only ensures that `stream` does not leak the secrets: it makes no guarantees about what happens if there is a memory safety vulnerability in the application linked against the library. Such a vulnerability can lead to a *read-gadget* that may be used to exfiltrate the secrets in `kcopy`! To defend against such gadgets, `LibSodium`'s developers take care to zero the intermediate memory used in `stream` (Line 7) to ensure those secrets cannot be leaked through memory vulnerabilities that reside in the application. Unfortunately, optimizations like dead-store elimination can remove secret scrubbing and nullify the efforts of `LibSodium`'s developers [61].

Security vs. performance. Read gadgets are but one of several possible attacks that library developers must defend against. For each attack, the library developer must either manually add relevant defenses to their code, or make an explicit choice *not* to do so (e.g. due to prohibitive overheads). Consequently, library code "bakes in" a subset of possible protections against application vulnerabilities: these may be unnecessary or insufficient depending upon the application context. For example, `LibSodium`'s zeroization protection against read gadgets is unnecessary when linked against a memory safe Rust application. On the other hand, `LibSodium`'s protections are insufficient against Spectre attacks [54]. The library's authors chose to elide an appropriate fence needed to protecting against speculative read gadgets as *all* the clients of the library would have to suffer the corresponding performance degradation, not just the ones where Spectre was a legitimate concern.

In a nutshell, cryptographic library developers are currently in a difficult position: they must make one-size-fits-all security-performance trade offs by manually inserting fragile protections orthogonal to the cryptographic algorithms and protocols they are implementing. Luckily, secure compilers offer a promising solution to escape this dilemma [41]. Library users are in a better position to make security-performance trade offs and can provide a compiler with security policies to be automatically enforced through inserted protections. To realize this vision, we introduce RoboCop, a new methodology and toolchain for building secure *and* efficient applications from cryptographic libraries. We develop RoboCop via four concrete contributions:

1. Abstraction: libraries and attackers (§3, §4.3). Our first contribution is a formal operational semantics describing the behavior of a library executing within a potentially vulnerable application. We further define a semantics capturing a high-level, abstract model of speculative execution, based on the notion of a speculation oracle that "guesses" the result of evaluating an expression and then later rolling back or committing if the guesses were correct. These semantics provide a unified setting for precisely stating different attackers' observations, guiding the design of our protections.

2. Specification: robust constant time (§4.2). Our second contribution is a novel security property, *robust constant time* (RCT), using our model to precisely define security for a cryptographic library running *in the context of* a potentially vulnerable application. Crucially, our definition is parameterized by an attacker model, capturing the set of attackers that a library is supposed to be secure against. We further define a speculative version, *robust speculative constant time*, capturing constant time in the presence of Spectre.

3. Implementation: the RoboCop compiler (§5). By factoring out assumptions about the context, our notion of RCT enables our third contribution: a compiler that takes a cryptographic library and synthesizes a bespoke binary tailored to the application against which the library will be linked. To do so, we show how to map each kind of attacker to a concrete code transform that provably, with respect to our definition of RCT, protects against that attacker. Thus, our RoboCop compiler lets library developers focus on implementing constant-time cryptographic algorithms, without having to worry about baking in potentially inefficient or insecure protections against application vulnerabilities. Instead, protections can be automatically inserted based on the application context, thereby ensuring the same library code can be securely *and* efficiently reused in all contexts.

4. Evaluation: SUPERCOP (§6). Finally, our fourth contribution is an empirical evaluation that shows that our RoboCop compiler can automatically generate protections for a wide range of cryptographic library code defending against a variety of attacks. Here we modify the `SUPERCOP` [56] cryptographic benchmarking suite. We instrument `SUPERCOP` to generate and measure the overhead of protecting against three classes of attacks: read gadgets (due to memory safety vulnerabilities in the application), speculative read gadgets (due to speculative execution in the application), and concurrent observations (due to threads in the application). In our test suite of 542 different implementations of cryptographic operations, we show that our RoboCop compiler can automatically generate code that is secure against application vulnerabilities with the majority of overheads under 2% when protecting against read gadgets and under 4% when protecting against speculative read gadgets, thereby demonstrating that RCT reconciles the tension between security and efficiency when reusing cryptographic libraries in different application contexts.

2 Overview

Every application that works with sensitive or personal user data uses cryptography to ensure the confidentiality or integrity of the data. These cryptographic operations typically rely upon sophisticated mathematics and are notoriously difficult to get right: bugs or security vulnerabilities within cryptographic code risk leaking critical secret keys or data, which could compromise the security of the whole application. Thus, cryptographic operations are typically implemented and encapsulated within *libraries* that are carefully authored and audited by cryptographic experts. These libraries provide trusted implementations of cryptographic operations (e.g. `LibSodium` [18]) or protocols (e.g. `OpenSSL` [5]), that can then be widely reused by developers—without requiring cryptographic expertise—to build secure applications.

In addition to correctly implementing cryptographic algorithms, the developers of cryptographic libraries must carefully ensure their code meets certain *generic* security requirements. For example, they must ensure that their libraries are *constant time*, meaning that the execution time of the library must be independent of the values of the secret data that the library operates over. Otherwise, an attacker can measure the timing variations to learn whether a secret conditioned branch or operation went one way or the other, and from that, eventually recover the data itself. There has been significant work [14, 16] on characterizing when a cryptographic library is constant-time and designing recipes to write constant time code, and this work guides guides both the design of cryptographic algorithms and their implementation in cryptographic libraries.

2.1 Application (attacker) assumptions

Sadly, timing leaks are not the only attacker capability cryptographic library developers need worry about. Ultimately, libraries are not executed in isolation: they are linked against *applications* written by non-expert developers: another source of vulnerabilities through which secrets can be leaked.

Defending against application vulnerabilities. Consider an application written in C that uses the `LibSodium` library. If this application has a *buffer overflow* leading to a memory read then an attacker targeting the whole program now has the ability to read any cryptographic secrets that are left accessible in memory, completely bypassing the need for timing channel based attacks.

In fact, the developers of `LibSodium` are keenly aware of the need to defend against these potential application vulnerabilities. Figure 1 shows an excerpt of `LibSodium`'s reference implementation of the `Salsa20` stream cipher. `stream` takes as input a key `k` and a nonce `n` and outputs a length `clen` stream of pseudo random bytes in the buffer `c`. The developers make a copy, `kcopy`, of the key buffer, `k`. This copied buffer contains secret data that must not be left on the stack, as otherwise a buffer overflow in the application would let an attacker read the secret key off the stack. Thus, in anticipation of `LibSodium` being used in the context of a vulnerable application, on Line 7 the `LibSodium` developers invoke `sodium_memzero` to *zero* out the contents of `kcopy`, thereby ensuring that its value is inaccessible regardless of memory vulnerabilities from the application.

Attacker (application) assumptions. In general, cryptographic libraries must defend against vulnerabilities in their host applications. We capture these threats (attacker capabilities) as *assumptions* about application behavior. We specifically adopt the threat models assumed by real world cryptographic libraries: applications may be vulnerable to (speculative) read gadgets, but those vulnerable to control flow hijacking (e.g., because of write gadgets) are out of scope as most cryptographic libraries assume the application is not *completely* under attacker control.[1]

1. Attacks via memory unsafety. The first class of assumptions, illustrated by the code in Figure 1, is that owing to memory unsafety, there exist *memory read gadgets* within the application. Zeroing buffers (like `kcopy`) is one key defense against such gadgets, but does nothing to prevent the attacker from reading the original version of the key `k` which remains unzeroed. To protect `k` from read gadgets, libraries like `Libsodium` offer memory protection APIs which must be manually inserted and toggled on and off by application developers, and hence, are prone to incorrect usage.

2. Attacks via speculation. If the application is written in a memory safe language like Rust, then the library developer need not fret about read gadgets, and can avoid the overhead of zeroing out secrets. However, the library developer must still contend with the spectre of hardware speculation and its attendant vulnerabilities [25, 31, 32, 35, 50, 60]. There is work on extending constant time protections to Spectre vulnerabilities, but it focuses on protecting cryptographic code *itself* from speculatively leaking secrets. Owing to the overheads imposed by such protections, it is currently unreasonable to apply the same protections to the entirety of application code. As such, library developers may have to contend with attacks based on Spectre vulnerabilities *in the application* [36].

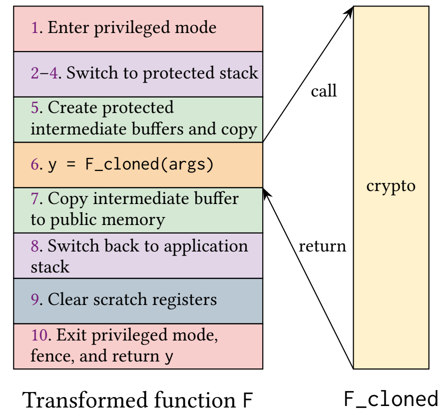
Indeed, in the case of `stream`, Spectre leads to a potential security issue with the clearing of `kcopy`: The `LibSodium` developers forgo an appropriate memory fence in `sodium_memzero`, leading to the possibility of the zeroed memory being read by speculatively executing application code *before* it is zeroed [54].[2] The fence was eschewed due to its performance overhead: *all* clients of `Libsodium` would suffer the performance degradation for protecting against Spectre. By manually adding or changing protections, `Libsodium`'s developers implicitly restrict their defenses to certain classes of attackers: i.e. they assume that the only application vulnerabilities are *non-speculative* read gadgets.

[1] Applications are also starting to employ software and hardware CFI techniques such as Clang's CFI [53] and Intel's CET [1].

[2] Well-documented issues with zeroing functions in high-level code make the zeroing best effort regardless of Spectre [44].

```c
int stream(u8 *c, u64 clen, u8 *n, u8 *k) {
  mpk_allow_access();
  switch_to_protected_stack();
  u8 *c_internal = mpk_malloc(clen);
  int y = stream_cloned(c_internal, clen, n, k);
  memcpy(c_internal, c, clen);
  switch_to_unprotected_stack();
  clear_scratch_registers();
  mpk_disable_access();
  fence;
  return y;
}
```

(a) Protections applied to LibSodium's Salsa20.



(b) Wrapping of cryptographic function.

Fig. 2. RoboCop protections.

3. Attacks via concurrency. The last category of application assumptions that we consider is whether the application is *concurrent*. In a concurrent context, work like Spectre-Declassified [51] has shown that an attacker can recover secret information if they can observe intermediate results, thus enhancing the reach of (speculative) read gadgets. In fact, the possibility of such observations are cited by the `LibSodium` developers as a reason to forgo the memory fence in `sodium_memzero` [54].

2.2 Robust constant time

Unlike with the property of constant time, there does not exist a security property capturing when a cryptographic library is secure when running *in the context of* a potentially vulnerable application. To address this and to capture the different application assumptions that a cryptographic library developer needs to consider we introduce the notion of *robust (speculative) constant time* (RCT). Like other robustness properties [2, 20, 21, 42, 43, 48, 52], a cryptographic library being robustly constant time captures that the library does not leak secrets *when linked against a context (application)*.

RCT serves as a formal security model for how cryptographic code is actually developed and used: protections are applied to the libraries with the goal that their use within an as yet unknown application will remain secure. Attacker capabilities can then be directly expressed as assumptions about the contexts in which the library code will run. For instance, `LibSodium`'s implicit protections can be explicitly specified as providing RCT assuming only the presence of single-threaded attacks based on memory unsafety (i.e. the presence of gadgets that can perform out-of-bounds reads).

2.3 A robust constant time compiler

By factoring out security assumptions about the context, our notion of RCT lets us design and develop a compiler that takes as input: (1) an implementation of a cryptographic library[3] and (2) an explicit set of *assumptions* about application/attacker capabilities and then synthesizes a protected library that is guaranteed to be robustly constant time with respect to the given attacker. Library developers can then focus on implementing cryptographic algorithms and their library can be used securely and efficiently in a variety of contexts.

[3]Our compiler assumes that the library is already (speculatively) constant-time. In §4.2 we discuss how RoboCop's robust protections are orthogonal to existing (speculative) constant time protections, thus allowing library developers to use existing automated tools or manual technique to meet this assumption.

Bespoke protection. Cryptographic libraries like `LibSodium` are designed to be used in a broad range of applications. As we saw with `libsodium_memzero`, the current state of manually baking protections into library code means that a single decision is made trading off between performance and which contexts the library is robust against. Instead, by factoring the protections out of the library and placing them in the compiler, RoboCop, developers can tailor protections to exactly the level required by the particular application context. For example, when used in a memory safe Rust application we can omit the unnecessary zeroing from `LibSodium`.

Efficient protection via wrapping and MPK. Our notion of RCT lets us use modern hardware memory protections, in particular Intel™ Memory Protection Keys (MPK), to protect library code incredibly efficiently. Our key insight in RoboCop is to allocate secret keys and carry out secret computation within a protected memory region. To do so, RoboCop wraps the library's external API functions, as illustrated by the protected version of `stream` in Figure 2a. We first enable access to the protected memory region and switch to a stack within protected memory. This protected stack ensures all intermediate computation remains protected. Line 4 shows the parameterized nature of RoboCop where we allocate a copy of the output buffer within protected memory. This allocation is *only* needed when (1) the underlying implementation uses the buffer for intermediate computation, and (2) the application context is *concurrent*. If these conditions are not met, RoboCop does not generate the extra allocation (the intermediate results *cannot* be observed by an attacker and hence, do not need to be protected). The wrapper function then calls the original library implementation, `stream_cloned`. After encryption, we copy the internal buffer back to the publicly visible `c` buffer (again, only if the attacker assumptions require.) The wrapper then switches back to the unprotected stack, clears scratch registers (if in Spectre protection mode where these registers may contain secret values that could be speculatively read), and finally disables access to the protected memory region before returning to the application (additionally fencing if in Spectre protection mode).

3 Security semantics

To formally ground robust constant time and the attacker models we develop a high-level, stateful calculus, λ_{spec}, whose syntax is shown in Figure 3. Syntactically, λ_{spec} is relatively standard, following λ_{rust}, CompCert, and others [29, 34] in employing a block-based memory model: memory is structured as "disjoint", fixed width blocks addressed via a block label and offset ($z_b[z_o]$). New blocks are allocated with $\texttt{new}_p\ e$ with e the size of the block and the protection label p determines whether the allocation is in a protected or unprotected memory "page". The size, set of values, and memory page are tracked in the second component of the non-speculative states (S). Correspondingly there is a protection operation, $\texttt{protect}_p$, which models hardware memory protection. $\texttt{protect}_p$ sets the memory access policy in the first component of the state: `public` only allows access to the public memory page whereas `protected` allows access to all memory. Dereferences are written $!\,e$ and assignments are written $e_{ptr} := e_{val}$. Functions are also stored (immutably) in memory.

We equip λ_{spec} with these semantics: a non-speculative semantics capturing a trace of events that we use to define our attacker models (§3.1) a novel, high-level speculative semantics (§3.2), and (speculative) concurrent semantics capturing a passive observer. Throughout this work we use $\bullet$ for an empty list, $\diamond$ for list concatenation, and an overline (e.g. $\overline{e}$) for a list of elements.

3.1 Non-speculative trace semantics

We first describe the structure of the traces the non-speculative semantics is designed to capture.

Traces. We are interested in capturing three aspects of the execution: (1) *who* (which party, `app` or `lib`) is executing code at a given time as well as the transfer of control between the parties, (2) *what memory* each party accesses, and (3) the internal branching which will be used to define the various

$$
\begin{array}{rcll}
\text{labels} & \ell & ::= & \texttt{app} \mid \texttt{lib} \\
\text{protection} & p & ::= & \texttt{public} \mid \texttt{protected} \\
\text{values} & v & ::= & z \mid z[z] \\
\text{expressions} & e & ::= & v \mid x \mid x\{v\} \mid \mathrm{op}(\overline{e}) \mid e[e] \mid \texttt{!}\,e \mid \texttt{new}_p\ e \mid e \texttt{:=} e \mid \texttt{protect}_p \mid e;e \\
& & \mid & \texttt{get-block}\ e \mid \texttt{get-offset}\ e \mid e(\overline{e}) \mid \texttt{fence} \mid \texttt{if}\ e\ \texttt{then}\ e\ \texttt{else}\ e \\
\text{states} & S & : & p \times (\mathbb{Z} \rightharpoonup m) \\
\text{memory slots} & m & : & \{size : \mathbb{Z}, p : p, v : [size] \rightharpoonup v\} \mid \lambda_\ell \overline{x}.e \\
\\
\text{events} & \epsilon & ::= & \delta^\ell \mid \tau^{\ell\to\ell} \\
\text{transition events} & \tau & ::= & \mathsf{call}\ z_f \mid \mathsf{ret}\ v \mid \mathsf{begin} \mid \mathsf{end}\ v \\
\text{domain events} & \delta & ::= & \mu \mid \mathsf{call}\ z \mid \mathsf{ret}\ v \mid \mathsf{branch}\ v \mid \mathsf{fence} \mid 0 \\
\text{memory label} & b & ::= & \mathsf{ib} \mid \mathsf{oob} \\
\text{memory events} & \mu & ::= & \mathsf{new}_p\ z@z \mid \mathsf{read}_b\ v \hookleftarrow z[z] \mid \mathsf{write}_b\ v \mapsto z[z] \mid \mathsf{protect}_p \\
\text{observable events} & c & ::= & 0 \mid \mathsf{branch}\ v \mid \mathsf{read} \hookleftarrow z[z] \\
& & \mid & \mathsf{write} \mapsto z[z] \mid \mathsf{new}_p\ z@z \mid \mathsf{call}\ z \mid \mathsf{end}\ v \\
\\
\text{speculation frame} & \Xi & : & (\widehat{S,\ \overline{K} :: e,\ \overline{\delta}}) \mid (S,\ \overline{K} :: e,\ \overline{\delta},\ \mu) \\
\text{speculative states} & \Phi & : & \{S : S, a : A, \Xi : \overline{\Xi}\} \\
\text{speculation directives} & d & ::= & \mathsf{nonspec} \mid \mathsf{spec}\ v \mid \mathsf{fence}
\end{array}
$$

Fig. 3. λ_{spec} syntax

constant time properties. To this end each reduction is labeled with an event, ϵ, whose syntax is shown in Figure 3. Labels are also attached to every function ($\lambda_\ell \overline{x}.e$) to distinguish application and library code. An event is either a labeled *domain event*, δ^ℓ, which captures an event executed by the party ℓ or a *transition event*, $\tau^{\ell\to\ell}$, which captures the transfer of control between the two parties.

The main transition events are call z_f and ret v which represent a call to the function at address z_f and returning from a function with return value v. Beyond the call and return events, there are begin and end v events that capture the (implicit) beginning and end of a trace. Domain events are either a *memory event* (μ), a call z_f and its corresponding ret v events (capturing a function call that stays within a single domain), a branch v event (capturing branching on the value v), or the empty event 0. Memory events are one of an allocation, a protection operation, a read, or a write and track the data associated with each operation (e.g. the value read/written, the location it was read/written from/to, and whether the memory access was in bounds or out of bounds).

Transition operational semantics. Our non-speculative semantics are split between two labeled reduction judgments: top-level transition reductions and domain reductions (Figure 4). Transition reductions are of the form $\langle S \mid \overline{K^\ell} :: e^\ell \rangle \overset{\epsilon}{\Rightarrow} \langle S \mid \overline{K^\ell} :: e^\ell \rangle$ where K are the standard call-by-value evaluation contexts. The state S tracks the memory and the current access level. To explain the control stack $\overline{K^\ell} :: e^\ell$ we examine the rule RED-CALL. The top of our stack (e^ℓ) is the mid-reduction body of the currently executing function (with label ℓ). For RED-CALL we are reducing a function call in the evaluation context K. We push this continuation (where the called function will return to) onto the stack of labeled continuations ($\overline{K'^{\ell'}}$) and then use the substituted body as the new execution frame. If the current label, ℓ, and the label of the function we are calling, ℓ_f, differ then we emit a transition event with label $\ell\to\ell_f$ otherwise we omit a domain event for the call.

Dually, rule RED-RET handles returning from a function. This is a transfer of control flow from ℓ, the label of the currently executing function, back to ℓ_K, the function caller. The return value is plugged into the top continuation on the stack and a corresponding return event is generated (we

$$\boxed{\langle S \mid \overline{K^{\ell}} :: e^{\ell}\rangle \overset{\epsilon}{\Longrightarrow} \langle S \mid \overline{K^{\ell}} :: e^{\ell}\rangle}$$

RED-CALL

$$\frac{S(z) = \lambda_{\ell_f}\overline{x}.e \qquad \epsilon = \begin{cases}(\text{call } z)^{\ell} & \text{when } \ell_f = \ell\\ (\text{call } z)^{\ell\to\ell_f} & \text{otherwise}\end{cases}}{\langle S \mid \overline{K'^{\ell'}} :: K[z(\overline{v})]^{\ell}\rangle \overset{\epsilon}{\Longrightarrow} \langle S \mid \overline{K'^{\ell'}} :: K^{\ell} :: e\overline{[v/x]}^{\ell_f}\rangle}$$

RED-RET

$$\frac{\epsilon = \begin{cases}(\text{ret } v)^{\ell} & \text{when } \ell_K = \ell\\ (\text{ret } v)^{\ell\to\ell_K} & \text{otherwise}\end{cases}}{\langle S \mid \overline{K'^{\ell'}} :: K^{\ell_K} :: v^{\ell}\rangle \overset{\epsilon}{\Longrightarrow} \langle S \mid \overline{K'^{\ell'}} :: K[v]^{\ell_K}\rangle}$$

RED-β

$$\frac{\langle S \mid e\rangle \overset{\delta}{\longrightarrow} \langle S' \mid e'\rangle}{\langle S \mid \overline{K'^{\ell'}} :: K[e]^{\ell}\rangle \overset{\delta^{\ell}}{\Longrightarrow} \langle S' \mid \overline{K'^{\ell'}} :: K[e']^{\ell}\rangle}$$

$$\boxed{\langle S \mid e\rangle \overset{\delta}{\longrightarrow} \langle S \mid e\rangle}$$

β-DEREF

$$\frac{\begin{array}{c}\text{accessible}(S, z_b)\\ z_o \in [S(z_b).size)] \qquad v = S(z_b).v(z_o)\end{array}}{\langle S \mid !\,(z_b[z_o])\rangle \xrightarrow{\text{read}_{\text{ib}}\ v \hookleftarrow z_b[z_o]} \langle S \mid v\rangle}$$

β-NEW

$$\frac{\begin{array}{c}z > 0 \qquad z_b = \text{fresh}(S) \qquad S.p \sqsubseteq p\\ S' = S[z_b := \{size = z, v = \bot, p = p\}]\end{array}}{\langle S \mid \text{new}_p\ z\rangle \xrightarrow{\text{new}_p\ z@z_b} \langle S' \mid z_b[0]\rangle}$$

β-DEREF-OOB

$$\frac{\begin{array}{c}z_b \notin \text{dom}(S) \vee z_o \notin [S(z_b).size] \qquad z'_b \in \text{dom}(S)\\ z'_o \in [S(z'_b).size)] \qquad v = S(z'_b).v(z'_o) \qquad \text{accessible}(S, z'_b)\end{array}}{\langle S \mid !\,(z_b[z_o])\rangle \xrightarrow{\text{read}_{\text{oob}}\ v \hookleftarrow z'_b[z'_o]} \langle S \mid v\rangle}$$

β-SUBST

$$\frac{}{\langle S \mid x\{v\}\rangle \overset{0}{\longrightarrow} \langle S \mid v\rangle}$$

Fig. 4. Non-speculative trace semantics excerpts

ignore same domain returns). The last "transition" reduction rule, RED-β, dispatches to the domain reduction relation ($\langle S \mid e\rangle \overset{\delta}{\longrightarrow} \langle S \mid e\rangle$), and labels the domain event δ with the current label.

Domain operational semantics. Most of the domain rules are standard and produce an empty trace event. Conditionals are also standard, but produce a branch event based on the condition. The rule β-NEW takes a protection domain p in which to allocate a new block of size z, checking that our current access level allows writing to the domain p using the "can-access" judgment $S.p \sqsubseteq p$.

The most notable of the reduction rules are those related to dereferencing and writing through pointers. The rules mirror each other so we will focus on dereferencing as it is simpler. In the rule β-DEREF we are dereferencing the pointer $z_b[z_o]$ with block label z_b and offset into that block z_o. To actually obtain the value at that location the following must hold: (1) the block must be accessible (accessible(S, z_b)), (2) the offset must be within the allocated size of the block ($z_o \in [S(z_b).size]$), and (3) a value must have been written to the block at the offset z_o ($v = S(z_b).v(z_o)$). When these conditions are met the value is read and a corresponding *in bounds* read trace event is generated. On the other hand, if either of the latter two conditions are not met the dereference is considered out-of-bounds and the rule β-DEREF-OOB applies instead. Here, we model the out-of-bounds read as a nondeterministic read from an arbitrary, valid, and accessible location $z'_o[z'_b]$.

$$\boxed{\langle \Phi \mid \overline{K} :: e \rangle \overset{\delta}{\hookrightarrow\!\!\!\!\twoheadrightarrow} \langle \Phi \mid \overline{K} :: e \rangle}$$

SPEC-NONSPEC

$$\frac{(a', \text{nonspec}) = \mathit{spec}(\Phi.a, e) \qquad \langle \Phi \mid \overline{K} :: e \rangle \overset{\delta}{\hookrightarrow} \langle \Phi' \mid \overline{K'} :: e' \rangle}{\langle \Phi \mid \overline{K} :: e \rangle \overset{\delta}{\hookrightarrow\!\!\!\!\twoheadrightarrow} \langle \Phi'[a := a'] \mid \overline{K'} :: e' \rangle}$$

SPEC-TRY-COMMIT

$$\frac{(a', \text{fence}) = \mathit{spec}(\Phi.a, e) \qquad \text{fence } \langle \Phi \mid \overline{K} :: e \rangle \text{ to } \langle \Phi' \mid \overline{K'} :: e' \rangle_{\overline{\delta}}}{\langle \Phi \mid \overline{K} :: e \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\!\twoheadrightarrow} \langle \Phi'[a := a'] \mid \overline{K'} :: e' \rangle}$$

SPEC-SPEC

$$\frac{(a', \text{spec } v) = \mathit{spec}(\Phi.a, K[e]) \qquad \langle \Phi \mid \bullet :: e \rangle \overset{\overline{\delta}}{\hookrightarrow}{}^{*} \langle \Phi' \mid \bullet :: v' \rangle \qquad \overline{\Xi'} = \text{makeFrame}_{v=v'}(\Phi.S, \overline{K'} :: K[e], \overline{\delta}) :: \Phi.\Xi \qquad e \neq \texttt{fence}}{\langle \Phi \mid \overline{K'} :: K[e] \rangle \overset{0}{\hookrightarrow\!\!\!\!\twoheadrightarrow} \langle \Phi[\Xi := \overline{\Xi'}, a := a'] \mid \overline{K'} :: K[v] \rangle}$$

$$\boxed{\langle \Phi \mid \overline{K} :: e \rangle \overset{\delta}{\hookrightarrow} \langle \Phi \mid \overline{K} :: e \rangle}$$

SPEC-β

$$\frac{\langle \Phi.S \mid \overline{K^{\ell}} :: e^{\ell} \rangle \overset{\epsilon}{\Longrightarrow} \langle S' \mid \overline{K'^{\ell'}} :: e'^{\ell'} \rangle \qquad \neg\text{stalled}(\Phi.\Xi, \Phi.S, \text{unlabel}(\epsilon)) \qquad \overline{\Xi} = \text{addEvent}(\Phi.\Xi, \text{unlabel}(\epsilon))}{\langle \Phi \mid \overline{K} :: e \rangle \xhookrightarrow{\text{unlabel}(\epsilon)} \langle \Phi[S := S', \Xi := \overline{\Xi}] \mid \overline{K'} :: e' \rangle}$$

Fig. 5. Small step speculative semantics

$$\boxed{\text{stalled}(\overline{\Xi}, S, \delta) : \overline{\Xi} \times S \times \delta \longrightarrow 2}$$

$$\begin{aligned}
\text{stalled}(\bullet, S, \text{fence}) &= \top \\
\text{stalled}(\Xi :: \overline{\Xi}, S, \text{read}_b\ v \leftarrow z_b[z_o]) &= (\text{protect}_p \in \Xi.\overline{\delta} \diamond \Xi.\overline{\mu} \land S(z_b).p = \text{protected}) \\
&\quad \lor (\text{stalled}(\overline{\Xi}, S, \text{read}_b\ v \leftarrow z_b[z_o]))
\end{aligned}$$

$$\boxed{\text{addEvent}(\Xi, \delta) : \Xi \times \delta \longrightarrow \Xi}$$

$$\begin{aligned}
\text{addEvent}((S, \overline{K} :: e, \overline{\delta}, \overline{\mu}), \text{read}_b\ v \leftarrow v_r) &= (S, \widehat{\overline{K} :: e, \overline{\delta}} \diamond \overline{\mu}) \quad \text{when } v_r \in \text{writeLocs}(\overline{\delta}) \\
\text{addEvent}((S, \overline{K} :: e, \overline{\delta}, \overline{\mu}), \text{write}_b\ v \mapsto v_w) &= (S, \overline{K} :: e, \overline{\delta}, \overline{\mu} \diamond \text{write}_b\ v \mapsto v_w)
\end{aligned}$$

$$\boxed{\text{fence } \langle \Phi \mid \overline{K} :: e \rangle \text{ to } \langle \Phi \mid \overline{K} :: e \rangle_{\overline{\delta}}}$$

FENCE-ROLLBACK

$$\frac{\Phi.\Xi = (S, \widehat{\overline{K'} :: e', \overline{\delta}}) :: \overline{\Xi} \qquad \Phi' = \Phi[S := S, \Xi := \overline{\Xi}]}{\text{fence } \langle \Phi \mid \overline{K} :: e \rangle \text{ to } \langle \Phi' \mid \overline{K'} :: e' \rangle_{\bullet}}$$

FENCE-COMMIT

$$\frac{\Phi.\Xi = (S, \overline{K'} :: e', \overline{\delta}, \overline{\mu}) :: \overline{\Xi} \qquad S' = \text{commit}(S, \overline{\delta} \diamond \overline{\mu}) \qquad \overline{\Xi'} = \text{addEvents}(\overline{\Xi}, \overline{\delta} \diamond \overline{\mu}) \qquad \Phi' = \Phi[S := S', \Xi := \overline{\Xi'}]}{\text{fence } \langle \Phi \mid \overline{K} :: e \rangle \text{ to } \langle \Phi' \mid \overline{K} :: e \rangle_{\overline{\delta}}}$$

Fig. 6. Speculative semantics auxiliary definitions

3.2 Speculative semantics

To model Spectre and speculative execution broadly we define a second operational semantics for λ_{spec}. Instead of modeling a specific version of Spectre we seek to capture a high-level essence of speculation, namely that speculation is the combination of guessing how an expression might evaluate and then either rolling back or committing if the guess was correct. That is, different types of speculation can be captured as the following sequence: First, instead of evaluating an expression, make up the value you think it would evaluate to and continue running using that value instead. While running "under speculation" check if any operation invalidates the speculative guess. Then, once you hit a "fence", rollback to the speculation point if the guess was invalid, or commit the effects of the skipped expression and continue evaluating.

We capture (excerpts of) these notions in Figures 5 and 6. Figure 5 defines the top-level reduction relation $\langle \Phi \mid e \rangle \xhookrightarrow{\overline{\delta}} \langle \Phi \mid e \rangle$. Speculative states, Φ, extend the non-speculative state (S) with a microarchitectural state (a) and a stack of speculation frames (Ξ). Speculation frames come in two forms, a "mispeculation" frame, $\widetilde{(S,\ e,\ \overline{\delta})}$, capturing that the current speculation is invalid and will be rolled back to the state S and expression e and "in progress" frames, $(S,\ e,\ \overline{\delta},\ \overline{\mu})$, capturing that the current speculation is potentially valid. The $\overline{\delta}$ and $\overline{\mu}$ components capture the trace of events of the skipped expression and the subsequent memory events, respectively. The skipped events are used if we commit a speculative frame, "replaying" the events that were speculatively passed over, and the memory events are used to determine whether speculation is invalid (discussed below).

The decision of whether and how to speculate is handled by a microarchitectural state "speculation" function ($spec : A \times e \longrightarrow A \times d$). The function takes the current microarchitectural state and the current state of the executing function, updates the microarchitectural state, and returns a speculation directive, d. This directive says whether we will be speculating with the guessed value v (spec v), not speculating (nonspec), or performing a fence operation (fence).

Speculation. To see how speculation plays out we will go over the three corresponding rules: SPEC-NONSPEC, SPEC-SPEC, and SPEC-TRY-COMMIT and how they capture speculation in the classic Spectre exploit of bypassing a bounds check when evaluating `if` $i\{100\} < 10$ `then` $!\,b[i\{100\}]$ `else 0` (where the size of the block b is 10). The term $i\{100\}$ captures a delayed substitution: earlier in the execution the value 100 was substituted for i.[4] We first evaluate the term $i\{100\}$ and apply the rule SPEC-NONSPEC: the speculator returns that it does not want to speculate on this expression and evaluation proceeds normally (via the SPEC-β rule) so our branch condition is now $100 < 10$. Here the speculator consults its microarchitectural state and sees that every other time we have done this check the result has been true. The speculator thus returns spec 1 (and a new microarchitectural state) and the rule SPEC-SPEC applies. SPEC-SPEC runs the skipped expression capturing any memory events (writes, reads, etc.) but does not commit them, instead saving them in a new speculation frame via makeFrame. makeFrame checks if the speculated value matches the real value: in this case it does not and therefore our new frame is a mispeculation frame saving the current state and continuation on our stack Ξ. From here evaluation continues with the speculated value 1 and another two SPEC-NONSPEC steps evaluate $!\,b[100]$.

In these SPEC-NONSPEC steps the rule SPEC-β handles two additional aspects beyond the evaluation. First, it checks that the instruction is not stalled ($\neg\mathsf{stalled}(\Phi.\Xi, \Phi.S, \mathsf{unlabel}(\epsilon))$).[5] This captures that fence instructions will not execute speculatively and that, with MPK, writes and reads to `protected` memory will not execute if the protection level is speculatively uncertain (the rules for fences and reads are shown in Figure 6). SPEC-β also adds the new event to the speculative stack ($\overline{\Xi} = \mathsf{addEvent}(\Phi.\Xi, \mathsf{unlabel}(\epsilon))$), updating whether the current speculation is invalid or not. Excerpts of addEvent are shown in Figure 6. For reads it checks if the read is to the location of a speculatively skipped write: if so the current speculation is invalid and will be rolled back (but

[4]We use delayed substitutions (β-SUBST) to capture the fact that, when executing on hardware, argument substitution will be compiled to a register access or memory lookup and as such should not be treated as an immediate value.

[5]The unlabel function removes labels and is defined in Appendix A.

$$\begin{array}{rrcl}
\text{API contexts} & \Gamma & ::= & \overline{(f : z)} \\
\text{libraries} & L & ::= & \overline{f \mapsto (z_f, \lambda \overline{x}.e)} \\
\text{secret contexts} & \Delta & ::= & \overline{x \mapsto (z_{loc}, z_{len})} \\
\text{heaplets} & H & : & \mathbb{Z} \rightharpoonup m
\end{array}
\qquad
\begin{array}{rrcl}
\text{app traces} & A & ::= & \tau^{\texttt{lib}\to\mathsf{app}} \diamond \overline{\delta^{\mathsf{app}}} \diamond \tau^{\mathsf{app}\to\texttt{lib}} \\
\text{lib traces} & L & ::= & \tau^{\mathsf{app}\to\texttt{lib}} \diamond \overline{\delta^{\texttt{lib}}} \diamond \tau^{\texttt{lib}\to\mathsf{app}} \\
\text{traces} & T & ::= & A \circ L \circ T \\
 & & \mid & \tau^{\texttt{lib}\to\mathsf{app}} \diamond \overline{\delta^{\mathsf{app}}} \diamond \mathsf{end}\ v^{\mathsf{app}\to\texttt{lib}}
\end{array}$$

Fig. 7. Syntax of programs and traces

continues executing until a fence). The displayed rule for writes shows how other events are added to an in progress frame to be used to check the validity of previous speculation.

Back in our example, were we to continue evaluating we would reach the classic Spectre out of bounds read, however we will instead assume that the speculator decides to stop speculating and returns the fence directive. The rule SPEC-FENCE thus applies and we turn to the judgment fence $\langle \Phi \mid\ !b[100] \rangle$ to $\langle \Phi_1 \mid e_1 \rangle_{\overline{\delta}}$. This judgment, defined in Figure 6, returns the state and expression with which we will continue evaluation as well as the trace of events from evaluating the speculatively skipped expression. If the speculation was invalidated (FENCE-ROLLBACK) then we will return to the continuation where speculation began (this applies to our example and we return to the saved continuation $\texttt{if}\ 100 < 10\ \texttt{then}\ !b[i\{100\}]\ \texttt{else}\ 0$ and the respective state at the time of speculation). If the speculation had been valid then FENCE-COMMIT would apply. This commits any memory events that were speculatively skipped and checks whether the new, now committed events invalidate previous speculation (we allow nested speculation thus the speculative "stack").

3.3 Concurrent observer semantics

To capture concurrent observer capabilities we layer another semantics on top of both the speculative and non-speculative semantics. We show the new judgment for the concurrent speculative semantics below. It consists of a singular rule that, before any step, adds a read event for every memory location visible to a concurrent thread. As MPK guarantees thread local protection this consists of every location that is not in the `protected` memory region, even if our "main" thread currently has access to protected memory. The non-speculative version is defined similarly.

$$\frac{\langle \Phi \mid \overline{K} :: e \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\!\rightarrow} \langle \Phi' \mid \overline{K'} :: e' \rangle \qquad \overline{\delta'} = [\mathsf{branch}\ v \mid \langle \Phi[S.p := \texttt{public}] \mid\ !z_b[z_o] \rangle \overset{\mu}{\hookrightarrow} \langle \Phi' \mid v \rangle]}{\langle \Phi \mid \overline{K} :: e \rangle \xrightarrow{\overline{\delta'} \diamond \overline{\delta}}\!\!\!\!\rightarrow_C \langle \Phi' \mid \overline{K'} :: e' \rangle}$$

4 Robust constant time

Next we formalize the security semantics of cryptographic libraries when used within an application.

4.1 Programs and Traces

Figure 7 shows the syntax for defining libraries and applications.

Libraries. An *API context*, Γ is a map from function names to the number of arguments that function takes and defines the external API for a library. Given an API context Γ a *library* L is a set of functions and their heap locations for each of the external names in Γ. We capture this with the well-formedness judgment $\Gamma \vDash L$, which additionally allows L to contain internal functions (defined in Figure 17 in Appendix B).

Applications. We assume that secrets are stored in memory. A *secret context* Δ is a map from secret variable names x (used to refer to that secret block in application code) to a pair of integer values denoting the location (address) and length of the corresponding block. Given an API context Γ and a secret context Δ, an *application* is a pair of a *heaplet* H (a partial heap containing initial state and application functions), and a "main" expression e with free variables from Γ and Δ. We define a well-formedness judgment for applications $\Gamma, \Delta \vdash (H,\ e)$ (detailed in Figure 17 in Appendix B).

Programs. A *whole program* is then composed of a library L plugged into an application $(H,\ e)$. To define this we build a judgment $L \mid \Delta \mid H \vDash S$ (defined in Figure 17 in Appendix B). This captures that the initial state S (1) contains all of the functions defined in L, (2) contains locations for all of the secrets in Δ, and (3) contains the application heaplet H. We then define substitution operations $e[\Delta][L]$ and $S[\Delta][L]$ replacing the API variable names from Γ with the function locations in L and the secret variables in Δ with their actual block locations (defined in Figure 18 in Appendix B).

Program Traces. Our operational semantics captures a sequence of events, but for whole programs we factor this sequence into a *program trace* (T) with additional structure (shown in Figure 7). This trace captures the decomposition of execution into alternating sequences of application and library domain events, with transition events the boundaries between them. Application traces A are thus a transition event from library to application, followed by a sequence of application domain events, and then a transition back to the library. Library traces are defined similarly and the trace of an entire program is then alternating sequences of these application and library traces. For program traces we define a gluing concatenation operator $\overline{\epsilon_1}\epsilon \circ \epsilon\overline{\epsilon_2}$ which matches the sequence $\overline{\epsilon_1}\epsilon\overline{\epsilon_2}$. We use this to capture that the transition event ending an application trace and starting the subsequent library trace are in fact the same transition event. We then define trace and speculative trace metafunctions capturing the set of all program traces for a given program (the corresponding concurrent versions are as expected). These use the respective non-speculative and speculative termination judgments $\downarrow^{T}$ and $\downarrow_{S}^{\overline{\delta}}$ (which capture the terminating trace and are defined in Appendix B):

$$\begin{aligned} \mathsf{traces}(\langle S \mid e \rangle) &\triangleq \{\mathsf{begin} \diamond T \diamond \mathsf{end}\ v \mid \langle S \mid e \rangle \downarrow^{T} \langle S' \mid v \rangle\} \\ \mathsf{specTraces}(\langle \Phi \mid e \rangle) &\triangleq \{\mathsf{begin} \diamond \overline{\delta} \diamond \mathsf{end}\ v \mid \langle \Phi \mid e \rangle \downarrow_{S}^{\overline{\delta}} \langle \Phi' \mid v \rangle\} \end{aligned}$$

4.2 Robust constant time

Our core security property, robust constant time, much like classic constant time, comes in two flavors: speculative and non-speculative (we also separate the associated concurrent versions).

Attacker predicates. Intuitively, robust constant time captures that a library behaves correctly when plugged into an "unknown" context (in this case an application) representing an attacker. For cryptographic libraries the attacker attempts to exploit vulnerabilities in the application to extract secrets, so we parameterize our definition by an *attacker*: a predicate on applications, $\mathsf{pred}_{\Gamma,\Delta} : \wp(H,\ e)$ that captures a vulnerability as a set of applications with that vulnerability. We instantiate $\mathsf{pred}_{\Gamma,\Delta}$ in Section 4.3 to capture read-only and memory-safe attackers.

Robust constant time. Robust constant time, then, is a parameterized, robust version of classic constant time definitions. Let $\mathsf{ct} : \overline{\epsilon} \to \overline{c}$ be a meta-function that erases the components of the trace that are not visible from timing-based attacks (defined by the syntactic class of observable events in Figure 3). Robust constant time can thus be defined as follows:

Definition 1 (Robust constant time). *We say a library* $\Gamma \vDash L$ *is* robustly constant time *for an attacker class* $\mathsf{pred}_{\Gamma,\Delta}$ *if, for all secret contexts* Δ*, applications* $\Gamma, \Delta \vdash (H,\ e)$ *such that* $\mathsf{pred}_{\Gamma,\Delta}(H,\ e)$*, and initial states* S_0, S_0' *such that* $L \mid \Delta \mid H \vDash S_0 = S_0'$ *we have that* $\mathsf{ct}(\mathsf{traces}(\langle S_0[\Delta][L] \mid e[\Delta][L]\rangle)) = \mathsf{ct}(\mathsf{traces}(\langle S_0'[\Delta][L] \mid e[\Delta][L]\rangle))$.

The judgment $L \mid \Delta \mid H \vDash S_0 = S_0'$ is a variant of our well-formedness judgment for initial states capturing that S_0 and S_0' *only* vary at the secret locations contained in Δ. Robust speculative constant time is defined similarly, however we consider the speculative traces and a speculative attacker oracle $\mathsf{spec} : A \times e \to A \times d$ (capturing varying speculative attacks). The concurrent versions simply use the set of concurrent traces. These definitions can be found in Appendix B.1.

$$\boxed{\Gamma \vdash \text{read-only } T}$$

READ-ONLY-REC

$$\frac{\begin{array}{c} A = \tau_1{}^{\texttt{lib}\to\texttt{app}} \diamond \overline{\delta_A}^{\,\texttt{app}} \diamond \tau_2{}^{\texttt{app}\to\texttt{lib}} \qquad L = \tau_2{}^{\texttt{app}\to\texttt{lib}} \diamond \overline{\delta_L}^{\,\texttt{lib}} \diamond \tau_3{}^{\texttt{lib}\to\texttt{app}} \\ \tau_2 = \text{call } z_f \Rightarrow z_f \in \text{dom}(\Gamma) \qquad \text{wf-read-only } \overline{\delta_A} \qquad \text{wf-read-only } \overline{\delta_L} \implies \Gamma \vdash \text{read-only } T \end{array}}{\Gamma \vdash \text{read-only } A \circ L \circ T}$$

Fig. 8. Excerpt of read-only attacker model definition

4.3 Attackers

As discussed in Section 2.1 we wish to protect libraries against different kinds of application vulnerabilities: read-only vulnerabilities (corresponding to the attacker model libraries like `LibSodium` assume), speculative vulnerabilities, and concurrent observers. Speculative vulnerabilities and concurrent observers are captured by our speculative and concurrent semantics leaving us with non-speculative read-only attackers (which we contrast with memory-safe attackers such as those written in memory-safe languages). We will use the fact that our traces T capture the back and forth of actions of the application and library and the hand-offs between them to instantiate the predicate of Definition 1. We then define read-only and memory-safe attackers by restricting the set of unsafe behaviors during the application subtraces: Read-only attackers can read memory they were not given access to but cannot write to it (and thus cannot carry out *active* attacks). In contrast, memory-safe attackers may neither read nor write memory they were not given access to. We define both read-only and memory-safe attackers as excluding control-flow exploits.

We wish to capture sets of attackers by their trace properties, however the attacker predicate is defined as a property of applications which are only a partial program and cannot be run. As such we must first link with a library before we can assess the application's (mis)behavior. But we cannot simply take an arbitrary library: an ill-formed library can break application invariants. To untangle this knot we simultaneously define the relevant *restrictions* on the application we are classifying with the *assumptions* that it may make about the library that it is running against.

Definition 2 (Read-only attackers). *We say an application* $\Gamma, \Delta \vdash (H,\ e)$ *is a* read-only attacker *if, for all libraries* $\Gamma \vDash L$*, initial states* S_0 *such that* $L \mid \Delta \mid H \vDash S_0$*, and traces* $T \in \text{traces}(\langle S_0[\Delta][L] \mid e[\Delta][L]\rangle)$, $\Gamma \vdash \text{read-only } T$.

Figure 8 shows an excerpt of the judgment $\Gamma \vdash \text{read-only } T$ which handles the restrictions on the application and assumptions on the library components of the trace T. The rule READ-ONLY-REC captures the back and forth of assumptions and restrictions: it decomposes the next application and library trace sequences, imposes the read-only restrictions on the application events (wf-read-only $\overline{\delta_A}$), requires that the call into the library is an API function, and then, under the assumption that the library events do not write out of bounds, inductively requires that the rest of the trace is read-only. The definition of wf-read-only can be found in Figure 19 in Appendix B: it captures that the application trace can only contain in bounds write events, but may contain arbitrary read events. Memory-safe attackers are defined similarly, with the predicate adjusted to also preclude reads from unexposed blocks.

5 A robust compiler

Guided by our formal models we develop RoboCop, a compiler providing robust constant time protections for cryptographic libraries against different attackers. RoboCop is built on top of the LLVM framework [33] and uses Intel™ Memory Protection Keys (MPK) to guarantee that secret data (cryptographic secrets and the data derived from them) is only accessible while executing trusted cryptographic library code. While RoboCop mainly operates on the library, there are two components that involve the application. First, cryptographic secrets often originate and are managed by application code, and it is therefore necessary to allocate these secrets in protected memory. To do so we provide manual MPK allocation APIs and also adapt techniques from CryptoMPK [28] to provide an alternative, automatic application transformation that securely allocates secrets. Further, for the efficiency of our library protections, we reuse a single stack allocated in protected memory. To enable this we develop a simple LLVM pass that allocates this stack on program entry.

5.1 Making libraries robust

Cryptographic code operates directly on secret data and, to prevent timing-based leaks, is required to be constant time. With this baseline security requirement we operate under the assumption that cryptographic code is *trusted*. The task of RoboCop then is to ensure that the secret data remains inaccessible even if there are vulnerabilities within the client application code. These protections are provided in three steps: (1) Cryptographic developers label the external API functions. (2) RoboCop wraps these API functions to handle the memory isolation. (3) RoboCop replaces all dynamic memory functions (`malloc` and similar) with custom MPK compatible versions, ensuring that all memory allocated within the cryptographic library is kept within the protected domain.

Figure 2b shows our wrapping of cryptographic API functions. For every exported function `F` in the library, a clone, `F_cloned`, is generated containing the original implementation of `F`. *Internal* calls to `F` are replaced with calls to `F_cloned`: `F` becomes the external API wrapper for use by the client. `F` is responsible for switching into and out of the protected memory region.

The new `F` takes the following domain switching steps: (1) We enable access to the protected memory region with a `wrpkru` instruction. The specification for MPK [27] ensures this is speculatively secure: `wrpkru` will not execute speculatively and protected memory cannot be accessed until `wrpkru` is committed. (2) We get the address of the protected stack and save the current stack pointer. (3) We copy any stack arguments from the unprotected frame to the new protected stack. (4) We switch the stack pointer to the protected stack frame. (5) If concurrent protections are enabled, we allocate an internal copy buffer for the external buffers (discussed below). (6) We call `F_cloned`. (7) After the cryptographic function returns, we copy any internal buffers back to the original, public buffers. (8) We restore the previous stack pointer. (9) We clear all scratch registers which may potentially contain transient secret computation. (10) We disable access to protected memory, fence if speculative protections are enabled, and return to the application. Together these ensure that all data produced and used by the cryptographic code is within the protected memory region and the region is inaccessible to application code.

Concurrent protections. Rather than allocate extra memory, cryptographic algorithms sometimes carry out intermediate computations within output (often ciphertext) buffers. In a single-threaded context, this is safe as there is no way for client code to access these buffers before they contain their final, declassified (cryptographically secure) value. In a concurrent context, work like Spectre-Declassified [51] has shown that an attacker can recover secret information if they can observe intermediate results. To defend against these attacks we add a concurrent protection option to RoboCop. Here library developers additionally annotate API arguments that are used for intermediate computation, and RoboCop allocates memory for the arguments within protected memory, performs the intermediate work within the protected domain, and then copies the declassified result back out to the unprotected memory.

5.2 Proving RoboCop secure

On its own (concurrent) robust (speculative) constant time serves as a useful security property for understanding when library protections are secure against different attackers. However we are interested in automatically providing robust protections via RoboCop. To do so we need to understand our "source language": the assumptions we make about the source of our compilation. For RoboCop we assume that the source is *classically* (speculative) constant time and will prove that RoboCop then guarantees robust (speculative) constant time. This gives library developers flexibility: they can safely use any tool or handwritten technique to guarantee that their library implementation is constant time and then RoboCop lifts this to the *robust* counterpart.

In defining a "source language" that captures classic constant time, note that the classic notion of constant time is that executing a library function produces invariant traces. As such classic constant time is in fact a restricted form of robust constant time, where, for a given API context Γ, the main function is defined by the grammar $e_\Gamma ::= f(\overline{v})$ for $f \in \Gamma$. That is, "source applications" are simply individual function calls into the library. There is a slight caveat: classical constant time also requires that secrets have not leaked into the application state. With this in mind we define classical constant time as the following variation on robust constant time:

Definition 3 (Classical constant time). *We say a library* $\Gamma \vDash L$ *is* classically constant time *if, for all secret contexts* Δ*, classical "applications"* $\Gamma, \Delta \vdash (H, e_\Gamma)$*, and initial states* S_0, S_0' *such that* $L \mid \Delta \mid H \vDash S_0 = S_0'$*, we have that for all traces* $\langle S_0[\Delta][L] \mid e_\Gamma[\Delta][L]\rangle \downarrow^{\overline{\delta}} \langle S_1 \mid v\rangle$ *there exists a trace* $\langle S_0'[\Delta][L] \mid e_\Gamma[\Delta][L]\rangle \downarrow^{\overline{\delta'}} \langle S_1' \mid v\rangle$ *such that* $\mathrm{ct}(\overline{\delta}) = \mathrm{ct}(\overline{\delta'})$ *and* $S_1(\mathrm{dom}(H)) = S_1'(\mathrm{dom}(H))$.

The speculative version is defined similarly and can be found in Appendix B.1. Our compiler correctness properties are then that the compiler guarantees (speculative) (concurrent) robust constant under the assumption that the library is (speculatively) classically constant time.[6]

A formal model of RoboCop. Formally, we represent RoboCop as a parameterized compiler $\mathbb{C}$ which transforms source libraries ($\Gamma \vDash L$) into protected targets L. We have four compilers: (1) $\mathbb{C}_{\mathrm{ro}}$, which protects libraries from read-only attackers, (2) $\mathbb{C}_{\mathrm{spec}}$ which protects against speculative attackers, and (3) $\mathbb{C}_{\text{ro-co}}$, and (4) $\mathbb{C}_{\text{spec-co}}$ which also protect against concurrent observers. $\mathbb{C}_{\mathrm{ro}}$ transforms all internal uses of $\mathtt{new}_p\ e$ into $\mathtt{new}_{\mathtt{protected}}\ e$, and wraps each external API function with $\mathtt{protect}_{\mathtt{protected}}$ and $\mathtt{protect}_{\mathtt{public}}$. $\mathbb{C}_{\mathrm{spec}}$ is the same as $\mathbb{C}_{\mathrm{ro}}$ but also inserts a fence before returns. The concurrent compilers additionally reallocate all external buffers used by the library within protected memory, and insert memory copying to and from the external and protected buffers.

To capture that the application manages the secret buffers and must allocate them in protected memory we slightly modify the initial state well-formedness judgment to capture that each secret block in Δ is allocated in the `protected` memory region. We additionally assume that applications and libraries do not contain any $\mathtt{protect}_p$ expressions prior to being run through our compiler. We prove that each compiler is secure and guarantees its corresponding robust constant time property. The theorem statement that $\mathbb{C}_{\mathrm{ro}}$ guarantees read-only robust constant time is shown below (the remaining theorems are in Appendix D):

Theorem 1 ($\mathbb{C}_{\mathrm{ro}}$ guarantees read-only robust constant time). *If* $\Gamma \vDash L$ *is classically constant time and does not contain any* $\mathtt{protect}_p$ *subterms, then* $\mathbb{C}_{\mathrm{ro}}(\Gamma \vDash L)$ *is robustly constant time for read-only attackers (that do not contain* $\mathtt{protect}_p$*).*

Proof sketches. To prove Theorem 1 we must show that, for *any* read-only attackers and any two initial states that vary only in the values of secrets, when we plug our compiled library into

[6] This property can be equivalently viewed as preservation of robust constant time from our source language of classical "applications" to the target language containing the contexts of our attacker model.

the application the execution under both states does not vary (up to the observable leakage). The difficulty arises in two places: firstly, we must reason about a (mostly) arbitrary application. Secondly, the inherent nondeterminism of our semantics (due to allocation and out of bounds memory semantics), means that we are dealing with sets of traces rather than just a single trace.

To handle these challenges we rely on the fact that our attacker model is defined as properties on the structured traces T. Our proof begins with the two initial states S and S' with the same whole program e and some trace T for $\langle S \mid e \rangle$. We must then show that there is a corresponding trace for $\langle S' \mid e \rangle$. To do so we inductively split the *read-only* trace T into its attacker and library subsequences and define a semantic interpretation that captures that there is a corresponding trace T' for $\langle S' \mid e \rangle$. The interpretation enforces the following invariant during application subtraces: (1) the access level in both states is `public`, (2) all memory that varies between the two states (contains secrets) is within protected memory, and (3) the same memory locations have been allocated in both states. These conditions and our assumption that our application is read-only allow us to prove that, for all of the application subsequences A of T, there is a T' that contains the *exact same* subsequences A. The first two conditions ensure that the memory that the application can access (even if it reads out of bounds) does not contain secrets and is therefore identical, and the third condition aligns the non-determinism between the two traces. For each non-deterministic choice in one application subtrace we can then always show that we can make the exact same non-deterministic choice in our other trace. Our interpretation of library subtraces requires that there exists a corresponding library trace with the same observable events, which follows from assuming the library is classically constant time. Put together these let us prove Theorem 1 (details are in Appendix D.1).

Our proof of the corresponding theorem for $\mathbb{C}_{\text{spec}}$ uses a similar technique of semantically interpreting traces as relations between the two states. Here we instead split the speculative traces into subsequences of non-speculative events (which correspond to an underlying non-speculative trace T), mispeculated events (where the speculation was rolled back), and "correctly speculated" events (where the speculative guess was eventually committed). We then show that the non-speculative and "correctly" speculated events are related to an underlying non-speculative trace and thus constant time by the proof of Theorem 1. To show the latter, we rely on the inserted `fence` instruction ensuring that we cannot speculatively return into the application. For mispeculated events we rely on the invariant that the application cannot change the protection level and a lemma that speculative execution cannot read protected memory unless the access level was changed non-speculatively (details are in Appendix D.2). This proof relies on a slightly stronger assumption than its non-speculative counterpart: namely that the library is classically speculative constant time *under an extended function context*. That is, adding new, unused functions should not break the classical speculatively constant time protections. To illustrate where this might fail, consider protections that rely on Intel's hardware Indirect Branch Tracking for Spectre BTB protections. This hardware feature adds an `ENDBRANCH` instruction labeling the set of legal jump targets. If a protected library was linked with an application containing `ENDBRANCH` instructions, the classical SCT protections might no longer hold. The assumption that classical SCT holds under an extended context is thus akin to the assumption that the application does not contain $\texttt{protect}_p$ subterms.

The corresponding proofs for concurrent observers strengthen the invariant to capture that unprotected buffers never contain secrets, even during library subtraces. Beyond this the proofs follow the same (non-)speculative reasoning as for $\mathbb{C}_{\text{ro}}$ and $\mathbb{C}_{\text{spec}}$.

6 Evaluation

To evaluate the cost of guaranteeing robust constant time we ask the following questions:

Q1: What is the overhead of *robust* constant time against read-only/speculative attackers? (§6.1)

Q2: What is the overhead of *robust* constant time against concurrent observers? (§6.2)

Benchmarks. To study the performance of RoboCop on a wide range of cryptographic code we modify the SUPERCOP [56] cryptographic benchmarking tool. SUPERCOP's benchmarking suite is broken down into operations: we focus on its collections of implementations of authenticated encryption (**aead**), Diffie-Hellman key exchange (**dh**), public key encryption (**encrypt**), key encapsulation mechanism (**kem**), public key signatures (**sign**), and stream cipher (**stream**) algorithms. Within each of these operations SUPERCOP collects multiple implementations of each algorithm: e.g. the **stream** data set contains several implementations of both Salsa20 and ChaCha20.

The design intent of SUPERCOP is to find the fastest each cryptographic algorithm can run on a given machine. To this end its benchmarking tries every implementation of an algorithm with each compiler (in our case Clang with the `-O3`, `-O2`, `-Os`, and `-O` flags). These implementations are benchmarked multiple times across a range of input data sizes. Due to the nature of our implementation of RoboCop we restrict one of these axes by removing all non-C/C++ implementations.

With its broad suite of algorithms and its find-the-fastest methodology, SUPERCOP is a more *robust* means of benchmarking cryptographic software security techniques and we encourage future designers to use it for benchmarking. We found its selecting from multiple compilation levels particularly beneficial as, due to the cascading effects of optimizations, comparing at the same optimization level is not truly a head-to-head comparison. Indeed we found that sometimes the fastest optimization level would differ between RoboCop and the baseline, with several instances of `-O3` producing significantly slower code in combination with the protections than `-O`.

Machine and software setup. We run all benchmarks on a 13th Gen Intel® Core™ i9-13900KS, with 125GB RAM, and running Linux kernel version 6.3.0. We run SUPERCOP configured to collect data only on cores with the same frequency and our data is collected from 5.6 GHz cores. RoboCop adds new passes to LLVM 16.0.2 and is split into two passes: the library pass adds the protections and the application pass allocates a stack in protected memory on program entry. SUPERCOP defines API functions for each operation: these are annotated as the external API for RoboCop. We manually label secret key buffers and insert protected allocations for them in the protected versions. For the concurrent protection benchmarks we label the ciphertext arguments on the **dh** and **stream** APIs as being used for internal computation. The speculative protections differ from the read-only protections solely in the insertion of a single fence before returning from the library. We use a modified version of jemalloc 5.2.0 [28] patched to provide MPK allocation functions. Our baseline replaces the libc malloc implementation with an unpatched version of jemalloc.

Summary of results. We find that robust constant time protections can generally be guaranteed with minimal overhead (less than 5% in almost all cases for both read-only and speculative protections), with a small number of outliers with a peak of 40% overhead. At small data sizes highly optimized stream ciphers also carry a large overhead (with a median around 33% for read-only and 37% for speculative protections), however these workloads take on the order of a few hundred cycles. We also find that concurrent protections add additional overhead but the resulting costs remain minimal (the majority have overheads under 6%).

6.1 Read-only and speculative attackers

We measure the cost of ensuring robust constant time against read-only and speculative attackers across six data sets (shown in Table 1 and Figure 9). For the largest data size for each benchmark we find that the median overhead across all benchmarked algorithms is below 1% for both read-only and speculative protections and that 75% of all implementations for each primitive have overheads under 2% for read-only protections and 4% for speculative protections. For algorithms with varying input lengths, SUPERCOP measures performance across a wide range of lengths. We show the

			READ-ONLY			SPECULATIVE		
Benchmark	N	Size	Q_1	**Median overhead**	Q_3	Q_1	**Median overhead**	Q_3
aead	385	2048	-0.26%	**0.16%**	0.65%	-0.15%	**0.15%**	0.71%
dh	9	fixed	-0.11%	**0.32%**	0.56%	0.36%	**0.46%**	1.50%
encrypt	15	4237	-0.57%	**-0.09%**	0.41%	-3.61%	**0.48%**	3.38%
kem	47	fixed	-0.59%	**0.04%**	0.55%	-0.82%	**0.00%**	0.49%
sign	40	4237	-1.00%	**-0.11%**	0.67%	-1.00%	**-0.22%**	0.97%
stream	11	4096	0.24%	**0.88%**	1.45%	0.51%	**0.92%**	1.90%

Table 1. Overheads for read-only and speculative protections vs. unprotected baseline. N is the number of algorithms in the dataset, Size is the size of the operation's input in bytes, Q_1 and Q_3 are the first and third quartile overheads, and the median overhead is of the overheads of the mean runtimes.

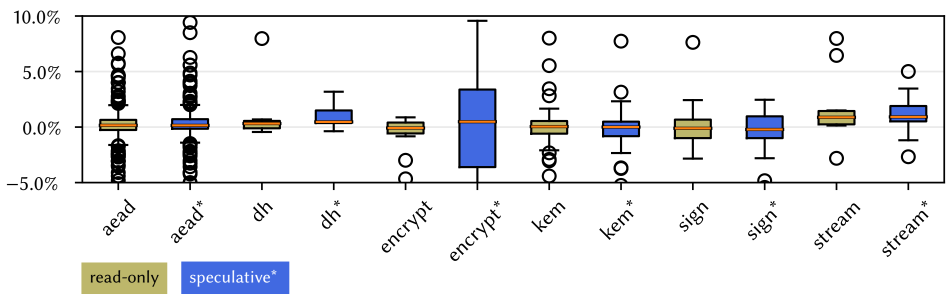


Fig. 9. Box plot of overheads for read-only and speculative protections* compared to unprotected baseline. Cutoff at 10% overhead, outliers beyond this point are discussed directly.[7]

	READ-ONLY			SPECULATIVE			
Size	Q_1	**Median overhead**	Q_3	Q_1	**Median overhead**	Q_3	Baseline cycles
1	29.07%	**33.40%**	55.56%	27.63%	**37.49%**	55.43%	4.29e+02
128	15.08%	**21.66%**	24.78%	15.63%	**21.25%**	25.35%	7.69e+02
256	8.23%	**12.01%**	17.52%	9.50%	**13.48%**	15.79%	1.44e+03
512	5.04%	**6.71%**	9.88%	5.03%	**6.80%**	8.81%	2.11e+03
1024	3.39%	**5.27%**	7.28%	3.48%	**4.94%**	6.60%	4.11e+03
2048	1.54%	**2.03%**	2.73%	1.60%	**2.45%**	3.82%	7.72e+03
4096	0.24%	**0.88%**	1.45%	0.51%	**0.92%**	1.90%	1.54e+04

Table 2. Read-only and speculative protection overheads for stream ciphers with varying plaintext sizes. Baseline cycles is the mean number of cycles for the unprotected baseline.

varying overheads across these sizes for stream ciphers in Table 2. The overhead increases as data sizes get smaller, with a median overhead of 33% for encrypting a single byte with read-only protections and 37% for speculative protections. Fortunately, at this data size encryption only takes a few hundred cycles so this high relative overhead remains a minimal raw cost.

Outliers. For 75% of implementations, read-only protections have overheads below 2% and speculative protections below 4%, there are some outliers above 10%: one **kem** implementation has a read-only overhead of 17% and a speculative overhead of 26%; two **sign** implementations have read-only overheads of 15% and 16% and speculative overheads of 16% and 35%; and four **aead** implementations have read-only overheads between 26% and 37% and speculative overheads between 28% and 42%. We do not have explanations for these higher overheads, however we observe that

[7]For clarity: 21/385 read-only and 23/385 speculative **aead** implementations lie above the whiskers and below the cutoff.

Base protections	NON-CONCURRENT Q_1	Median overhead	Q_3	CONCURRENT Q_1	Median overhead	Q_3
read-only	-0.11%	**0.32%**	0.56%	0.24%	**0.74%**	1.26%
speculative	0.36%	**0.46%**	1.50%	0.10%	**0.69%**	1.07%

(a) Diffie-Hellman key exchange (**dh**) algorithms

Base protections	Size	NON-CONCURRENT Q_1	Median overhead	Q_3	CONCURRENT Q_1	Median overhead	Q_3
read-only	4096	0.24%	**0.88%**	1.45%	1.77%	**2.39%**	5.67%
speculative	4096	0.51%	**0.92%**	1.90%	1.61%	**2.14%**	2.60%

(b) Stream (**stream**) ciphers

Table 3. Overhead of read-only and speculative protections vs. overhead with concurrent protections.

they call OpenSSL's cryptographic implementations and use internal randomness and dynamic allocation, though they are not the only algorithms that do so. We hypothesize the observed speedups are due to (1) noise as the baseline number of cycles is often small and (2) the separate protected allocator. We've previously observed better locality properties with separate library allocators.

6.2 Concurrent attackers

To protect against concurrent attackers it is necessary for buffers containing intermediate values derived from secrets to remain within the MPK protected memory. In its read-only attacker protections mode RoboCop ensures all memory originating from cryptographic code meets this requirement, however some cryptographic implementations use external, public buffers as internal, intermediate (private) buffers. To measure the cost of protecting these intermediate buffers we benchmark the **dh** and **stream** data sets with RoboCop's concurrent protections mode. We annotate the top-level SUPERCOP API functions `crypto_dh` and `crypto_stream` to mark the ciphertext argument as being used for intermediate computation. In the case of **stream** the size of this buffer is dynamically determined so we mark the plaintext length argument as the size for allocating a secure intermediate buffer. Table 3 shows the median overheads for the read-only protections compared to the median overheads for the concurrent protections. For our **dh** data set protecting these intermediate buffers increases the median overhead from 0.32% to 0.74% for read-only attackers and from 0.46% to 0.69% for speculative attackers. For our stream cipher data set at the largest data size the increase is greater, with the median increasing from 0.88% to 2.39% for read-only attacker and 0.92% to 2.14% for speculative attackers. We hypothesize that the higher overhead for stream ciphers is due to the dynamic allocation whereas the statically sized buffer of the Diffie-Hellman API allows optimizations in both allocation and copying back to the external buffer.

Our benchmarking treats every algorithm within each data set as if they use public buffers for intermediate computations. In practice RoboCop lets library designers label specifically which/if buffers are used for intermediate computations. This ensures that code that does not use the external buffers never has to pay the price for the extra allocation and copying and that the same library code base can be used in all contexts: in single-threaded mode RoboCop can omit the allocation and copy but in concurrent mode it handles the creation of a protected intermediate buffer.

7 Limitations

Our implementation of RoboCop has a few limitations: (1) While RoboCop handles observers in concurrent threads it does not handle running concurrent *cryptographic library* code. To handle this we could allocate a single protected stack per thread. It would be safe to reuse a single MPK protection key across all threads as concurrent threads would only have access while running cryptographic library code. (2) RoboCop assumes that no other code is using MPK. (3) RoboCop does not handle ensuring that protected memory does not get dumped on crashes, written to disk, or that it is cleared at the end of execution. These could be handled in one place by operating on MPK protected pages. (4) RoboCop assumes the speculative behavior of MPK follows Intel's specification and `wrpkru` will never execute speculatively and protected memory will never be accessed speculatively until protections have been fully committed [27]. Hardware bugs such as meltdown-pk [13] violate these assumptions, and we rely on hardware fixes for such bugs (for instance official patches have already been provided for the machine used for our evaluation).

8 Related Work

Memory isolation. MPK has been used to provide in-process memory isolation [26, 55], including between safe and unsafe Rust code [24, 30, 46]. MPK protections can be modified by unprivileged instructions, as such RoboCop assumes that applications do not have access to these instructions and including through control-flow hijacks. Countermeasures like binary rewriting [55], system call filtering [49, 58], and Cfi schemes [11, 12] could be used to lift these assumptions.

Similar to RoboCop, CryptoMPK [28] leverages MPK to protect the confidentiality of secret cryptographic data from memory disclosure vulnerabilities. We believe that it can guarantee robust constant time against read-only attackers, though their work is not framed in this manner. CryptoMPK differs from RoboCop in several fundamental ways: First, it is instead a whole-program analysis and transformation, identifying "crypto buffers" throughout the program and toggling MPK protection around use sites. As such it inserts significantly more context switches than RoboCop, and dynamically allocates and frees secret stack buffers. RoboCop's trusted library model avoids these costs, allowing a single context switch on library entry and exit, completely avoiding the need to dynamically allocate stack buffers and leading to significant performance gains. Second, by avoiding a whole program analysis RoboCop's model allows a much simpler implementation. This allows RoboCop to be applied to more complicated code and even hard to analyze assembly code (though we have not implemented this). Lastly, CryptoMPK does not handle robust speculative protections nor robust protections against concurrent attackers. As an optimization, CryptoMPK chooses not to securely allocate small secret stack buffers and instead inserts zeroing code. This zeroing code has the same trade offs between protecting against Spectre and performance as in `LibSodium`, and the buffers are also visible to concurrent attackers. CryptoMPK provides a `mxor` annotation to ensure that eventually declassified buffers are not marked as tainted when they are mixed with secret key data. This has the result of exposing these buffers to concurrent attackers.

Secure zeroization is often deployed as a manual protection in cryptographic libraries. Unfortunately, implementing secure memory zeroing in a high-level language is essentially impossible [44, 47, 61]. Recent work [40] shows how to develop a compiler pass to implement secure zeroization. RoboCop avoids these issues by restricting all secret data to a protected memory region, thus avoiding the need for zeroization (apart from register zeroing).

(Speculative) constant time. Many domain-specific languages and compilers have been developed to produce high-assurance cryptographic code [4, 7, 16, 45, 59]. Spectre attacks [31] significantly reduced the guarantees of these tools and prompted defenses against speculative leaks [15]. Jasmin implements Selective Speculative Load Hardening to protect against Spectre-PHT [51], performing stack zeroization and register clearing [40]. Blade inserts a minimum number of protections to prevent leaks via Spectre-PHT gadgets [57]. Swivel hardens WebAssembly sandboxes against speculative sandbox breakout and sandbox poisoning attacks [39, 62]. Serberus mitigates all currently known Spectre variants in code that follows the *static* constant-time discipline, which additionally prohibits secret function arguments and return values [38]. These tools serve as complements to

RoboCop: in combination they can be used to guarantee end-to-end protections against speculative attackers as discussed in §4.2. Notably, there is an exception to this statement in the case of Serberus: In handling Spectre-RSB [32], Serberus makes an implicit assumption that the return stack buffer is empty when entering cryptographic code. Much like the issues with prior work on constant time protections, Serberus is assuming that the cryptographic code represents the entire program. This can be remedied by RSB filling on entry to cryptographic code.

Foundations for cryptographic software security. Researchers have developed trace-based leakage models to reason about timing leaks in cryptographic code [8, 37]. These models have then been extended with prediction oracles [23], speculative semantics and directives [14, 17, 22] to capture leaks via (combinations of) different Spectre gadgets [19]. Our formal approach builds on these well-established practices. Our notion of RCT is inspired by previous work on secure compilers [3], which are formally defined as compilers that preserve classes of (hyper)-properties in adversarial contexts. Patrignani and Guarnieri [43] develop secure robust compilation criteria to formally examine the security guarantees of protections inserted by major compilers against Spectre-PHT, our approach to a secure compiler property follows this line of work.

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A Language

$$\boxed{\langle S \mid \overline{K^{\ell}} :: e^{\ell} \rangle \overset{\epsilon}{\Rightarrow} \langle S \mid \overline{K^{\ell}} :: e^{\ell} \rangle}$$

RED-CALL

$$\frac{S(z) = \lambda_{\ell_f} \overline{x}.e \qquad \epsilon = \begin{cases} (\text{call } z)^{\ell} & \text{when } \ell_f = \ell \\ (\text{call } z)^{\ell \to \ell_f} & \text{otherwise} \end{cases}}{\langle S \mid \overline{K'^{\ell'}} :: K[z(\overline{v})]^{\ell} \rangle \overset{\epsilon}{\Rightarrow} \langle S \mid \overline{K'^{\ell'}} :: K^{\ell} :: e\overline{[v/x]}^{\ell_f} \rangle}$$

RED-RET

$$\frac{\epsilon = \begin{cases} (\text{ret } v)^{\ell} & \text{when } \ell_K = \ell \\ (\text{ret } v)^{\ell \to \ell_K} & \text{otherwise} \end{cases}}{\langle S \mid \overline{K'^{\ell'}} :: K^{\ell_K} :: v^{\ell} \rangle \overset{\epsilon}{\Rightarrow} \langle S \mid \overline{K'^{\ell'}} :: K[v]^{\ell_K} \rangle}$$

RED-β

$$\frac{\langle S \mid e \rangle \overset{\delta}{\to} \langle S' \mid e' \rangle}{\langle S \mid \overline{K'^{\ell'}} :: K[e]^{\ell} \rangle \overset{\delta^{\ell}}{\Longrightarrow} \langle S' \mid \overline{K'^{\ell'}} :: K[e']^{\ell} \rangle}$$

$$\boxed{\langle S \mid e\rangle \xrightarrow{\delta} \langle S \mid e\rangle}$$

β-SUBST
$$\frac{}{\langle S \mid x\{v\}\rangle \xrightarrow{0} \langle S \mid v\rangle}$$

β-OP
$$\frac{v' = \delta(op)(\overline{v})}{\langle S \mid \mathsf{op}(\overline{v})\rangle \xrightarrow{0} \langle S \mid v'\rangle}$$

β-DEREF
$$\frac{\mathsf{accessible}(S, z_b) \qquad z_o \in [S(z_b).size)] \qquad v = S(z_b).v(z_o)}{\langle S \mid !\,(z_b[z_o])\rangle \xrightarrow{\mathsf{read}_{\mathsf{ib}}\ v \hookleftarrow z_b[z_o]} \langle S \mid v\rangle}$$

β-DEREF-OOB
$$\frac{z_b \notin \mathsf{dom}(S) \vee z_o \notin [S(z_b).size] \qquad z'_b \in \mathsf{dom}(S) \qquad z'_o \in [S(z'_b).size)] \qquad v = S(z'_b).v(z'_o) \qquad \mathsf{accessible}(S, z'_b)}{\langle S \mid !\,(z_b[z_o])\rangle \xrightarrow{\mathsf{read}_{\mathsf{oob}}\ v \hookleftarrow z'_b[z'_o]} \langle S \mid v\rangle}$$

β-WRITE
$$\frac{\mathsf{accessible}(S, z_b) \qquad z_o \in [S(z_b).size] \qquad S' = S(z_b).v[z_o := v]}{\langle S \mid z_b[z_o] := v\rangle \xrightarrow{\mathsf{write}_{\mathsf{ib}}\ v \mapsto z_b[z_o]} \langle S' \mid 0\rangle}$$

β-WRITE-OOB
$$\frac{z_b \notin \mathsf{dom}(S) \vee z_o \notin [S(z_b).size] \qquad z'_b \in \mathsf{dom}(S) \qquad z'_o \in [S(z'_b).size)] \qquad S' = S(z'_b).v[z'_o := v] \qquad \mathsf{accessible}(S, z'_b)}{\langle S \mid z_b[z_o] := v\rangle \xrightarrow{\mathsf{write}_{\mathsf{oob}}\ v \mapsto z'_b[z'_o]} \langle S' \mid 0\rangle}$$

β-NEW
$$\frac{z > 0 \qquad z_b = \mathsf{fresh}(S) \qquad S.p \sqsubseteq p \qquad S' = S[z_b := \{size = z, v = \bot, p = p\}]}{\langle S \mid \mathsf{new}_p\ z\rangle \xrightarrow{\mathsf{new}_p\ z@z_b} \langle S' \mid z_b[0]\rangle}$$

β-GET-BLOCK
$$\frac{}{\langle S \mid \texttt{get-block}\ (z_b[z_o])\rangle \xrightarrow{0} \langle S \mid z_b\rangle}$$

β-GET-OFFSET
$$\frac{}{\langle S \mid \texttt{get-offset}\ (z_b[z_o])\rangle \xrightarrow{0} \langle S \mid z_o\rangle}$$

β-IF-FALSE
$$\frac{}{\langle S \mid \texttt{if}\ 0\ \texttt{then}\ e\ \texttt{else}\ e'\rangle \xrightarrow{\mathsf{branch}\ 0} \langle S \mid e'\rangle}$$

β-IF-TRUE
$$\frac{v \neq 0}{\langle S \mid \texttt{if}\ v\ \texttt{then}\ e\ \texttt{else}\ e'\rangle \xrightarrow{\mathsf{branch}\ v} \langle S \mid e\rangle}$$

β-PROTECT
$$\frac{}{\langle S \mid \mathsf{protect}_p\rangle \xrightarrow{\mathsf{protect}_p} \langle S[p := p] \mid 0\rangle}$$

β-SEQ
$$\frac{}{\langle S \mid v; e\rangle \xrightarrow{0} \langle S \mid e\rangle}$$

β-FENCE
$$\frac{}{\langle S \mid \texttt{fence}\rangle \xrightarrow{\mathsf{fence}} \langle S \mid 0\rangle}$$

Fig. 10. Non-speculative trace semantics

$$
\begin{array}{rcl}
x[v/x] & \triangleq & x\{v\} \\
y[v/x] & \triangleq & y \\
(x\{v'\})[v/x] & \triangleq & x\{v'[v/x]\} \\
(y\{v'\})[v/x] & \triangleq & y\{v'[v/x]\} \\
& \cdots &
\end{array}
$$

$$
\frac{}{p \sqsubseteq p} \qquad \frac{}{\texttt{protected} \sqsubseteq \texttt{public}} \qquad \frac{}{\texttt{public} \not\sqsubseteq \texttt{protected}} \qquad \frac{S.p \sqsubseteq S(z_b).p}{\mathsf{accessible}(S, z_b)}
$$

Fig. 11. Non-speculative semantics auxiliary definitions

The following definitions assume a fixed microarchitectural state type A, and a fixed speculation oracle $spec : A \times e \longrightarrow A \times d$.

$$\boxed{\langle \Phi \mid \overline{K} :: e\rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow} \langle \Phi \mid \overline{K} :: e\rangle}$$

SPEC-NONSPEC

$$
\frac{(a', \mathsf{nonspec}) = spec(\Phi.a, e) \qquad \langle \Phi \mid \overline{K} :: e\rangle \overset{\delta}{\hookrightarrow} \langle \Phi' \mid \overline{K'} :: e'\rangle}{\langle \Phi \mid \overline{K} :: e\rangle \overset{\delta}{\hookrightarrow\!\!\!\rightarrow} \langle \Phi'[a := a'] \mid \overline{K'} :: e'\rangle}
$$

SPEC-SPEC

$$
\frac{(a', \mathsf{spec}\ v) = spec(\Phi.a, K[e]) \qquad \langle \Phi \mid \bullet :: e\rangle \overset{\overline{\delta}}{\hookrightarrow}{}^{*} \langle \Phi' \mid \bullet :: v'\rangle \qquad \overline{\Xi'} = \mathsf{makeFrame}_{v=v'}(\Phi.S, \overline{K'} :: K[e], \overline{\delta}) :: \Phi.\Xi \qquad e \neq \mathsf{fence}}{\langle \Phi \mid \overline{K'} :: K[e]\rangle \overset{0}{\hookrightarrow\!\!\!\rightarrow} \langle \Phi[\Xi := \overline{\Xi'}, a := a'] \mid \overline{K'} :: K[v]\rangle}
$$

SPEC-TRY-COMMIT

$$
\frac{(a', \mathsf{fence}) = spec(\Phi.a, e) \qquad \mathsf{fence}\ \langle \Phi \mid \overline{K} :: e\rangle\ \mathsf{to}\ \langle \Phi' \mid \overline{K'} :: e'\rangle_{\overline{\delta}}}{\langle \Phi \mid \overline{K} :: e\rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow} \langle \Phi'[a := a'] \mid \overline{K'} :: e'\rangle}
$$

$$\boxed{\langle \Phi \mid \overline{K} :: e\rangle \overset{\delta}{\hookrightarrow} \langle \Phi \mid \overline{K} :: e\rangle}$$

SPEC-β

$$
\frac{\langle \Phi.S \mid \overline{K^{\ell}} :: e^{\ell}\rangle \overset{\epsilon}{\Longrightarrow} \langle S' \mid \overline{K'^{\ell'}} :: e'^{\ell'}\rangle \qquad \neg\mathsf{stalled}(\Phi.\Xi, \Phi.S, \mathsf{unlabel}(\epsilon)) \qquad \overline{\Xi} = \mathsf{addEvent}(\Phi.\Xi, \mathsf{unlabel}(\epsilon))}{\langle \Phi \mid \overline{K} :: e\rangle \xhookrightarrow{\mathsf{unlabel}(\epsilon)} \langle \Phi[S := S', \Xi := \overline{\Xi}] \mid \overline{K'} :: e'\rangle}
$$

$$\boxed{\mathsf{stalled}(\overline{\Xi}, S, \delta) : \overline{\Xi} \times S \times \delta \longrightarrow \mathbb{2}}$$

$$
\begin{array}{rcl}
\mathsf{stalled}(\bullet, S, \mathsf{fence}) & = & \top \\
\mathsf{stalled}(\Xi :: \overline{\Xi}, S, \mathsf{read}_b\ v \hookleftarrow z_b[z_o]) & = & (\mathsf{protect}_p \in \Xi.\overline{\delta} \diamond \Xi.\overline{\mu} \wedge S(z_b).p = \mathsf{protected}) \\
& & \vee\ (\mathsf{stalled}(\overline{\Xi}, S, \mathsf{read}_b\ v \hookleftarrow z_b[z_o]))
\end{array}
$$

$$\begin{aligned}
\text{stalled}(\Xi :: \overline{\Xi}, S, \text{write}_b\ v \mapsto z_b[z_o]) &= (\text{protect}_p \in \Xi.\overline{\delta} \diamond \Xi.\overline{\mu} \wedge S(z_b).p = \text{protected}) \\
&\quad \vee (\text{stalled}(\overline{\Xi}, S, \text{write}_b\ v \mapsto z_b[z_o])) \\
\text{stalled}(\Phi, S, \delta) &= \bot
\end{aligned}$$

$$\boxed{\text{addEvent}(\overline{\Xi}, \delta) : \overline{\Xi} \times \delta \to \overline{\Xi}}$$
$$\boxed{\text{addEvent}(\Xi, \delta) : \Xi \times \delta \to \Xi}$$

$$\begin{aligned}
\text{addEvent}(\bullet, \delta) &= \bullet \\
\text{addEvent}(\Xi :: \overline{\Xi}, \delta) &= \text{addEvent}(\Xi, \delta) :: \overline{\Xi}
\end{aligned}$$

$$\begin{aligned}
\text{addEvent}((\widehat{S,\ \overline{K} :: e,\ \overline{\delta}}), \delta') &= (\widehat{S,\ \overline{K} :: e,\ \overline{\delta} \diamond \delta'}) \\
\text{addEvent}((S,\ \overline{K} :: e,\ \overline{\delta},\ \overline{\mu}), \delta') &= (S,\ \overline{K} :: e,\ \overline{\delta},\ \overline{\mu}) \text{ when } \delta' \neq \mu' \\
\text{addEvent}((S,\ \overline{K} :: e,\ \overline{\delta},\ \overline{\mu}), \text{read}_b\ v \hookleftarrow v_r) &= (\widehat{S,\ \overline{K} :: e,\ \overline{\delta} \diamond \overline{\mu}}) \text{ when } v_r \in \text{writeLocs}(\overline{\delta}) \\
\text{addEvent}((S,\ \overline{K} :: e,\ \overline{\delta},\ \overline{\mu}), \text{read}_b\ v \hookleftarrow v_r) &= (S,\ \overline{K} :: e,\ \overline{\delta},\ \overline{\mu}) \text{ when } v_r \notin \text{writeLocs}(\overline{\delta}) \\
\text{addEvent}((S,\ \overline{K} :: e,\ \overline{\delta},\ \overline{\mu}), \text{write}_b\ v \mapsto v_w) &= (S,\ \overline{K} :: e,\ \overline{\delta},\ \overline{\mu} \diamond \text{write}_b\ v \mapsto v_w) \\
\text{addEvent}((S,\ \overline{K} :: e,\ \overline{\delta},\ \overline{\mu}), \text{new}_p\ z@z_b) &= (S,\ \overline{K} :: e,\ \overline{\delta},\ \overline{\mu} \diamond \text{new}_p\ z@z_b) \\
\text{addEvent}((S,\ \overline{K} :: e,\ \overline{\delta},\ \overline{\mu}), \text{protect}_p) &= (S,\ \overline{K} :: e,\ \overline{\delta},\ \overline{\mu} \diamond \text{protect}_p)
\end{aligned}$$

Fig. 12. Small step speculative semantics

$$\boxed{\text{fence } \langle \Phi \mid \overline{K} :: e \rangle \text{ to } \langle \Phi \mid \overline{K} :: e \rangle_{\overline{\delta}}}$$

FENCE-NO-SPEC

$$\frac{\Phi.\Xi = \bullet}{\text{fence } \langle \Phi \mid \overline{K} :: e \rangle \text{ to } \langle \Phi \mid \overline{K} :: e \rangle_{\bullet}}$$

FENCE-ROLLBACK

$$\frac{\Phi.\Xi = (\widehat{S,\ \overline{K'} :: e'},\ \overline{\delta}) :: \overline{\Xi}}{\text{fence } \langle \Phi \mid \overline{K} :: e \rangle \text{ to } \langle \Phi[S := S, \Xi := \overline{\Xi}] \mid \overline{K'} :: e' \rangle_{\bullet}}$$

FENCE-COMMIT

$$\frac{\Phi.\Xi = (S,\ \overline{K'} :: e',\ \overline{\delta},\ \overline{\mu}) :: \overline{\Xi} \qquad S' = \text{commit}(S, \overline{\delta} \diamond \overline{\mu}) \qquad \overline{\Xi'} = \text{addEvents}(\overline{\Xi}, \overline{\delta} \diamond \overline{\mu})}{\text{fence } \langle \Phi \mid \overline{K} :: e \rangle \text{ to } \langle \Phi[S := S', \Xi := \overline{\Xi'}] \mid \overline{K} :: e \rangle_{\overline{\delta}}}$$

$$\begin{aligned}
\text{unlabel}(\delta^{\ell}) &\triangleq \delta \\
\text{unlabel}(\tau^{\ell \to \ell'}) &\triangleq \tau
\end{aligned}$$

$$\text{writeLocs}(\overline{c}) \triangleq \{v_w \mid \text{write}_{\in}\ v \mapsto v_w \overline{c}\}$$

$$\begin{aligned}
\text{commit}(S, \bullet) &\triangleq S \\
\text{commit}(S, \delta \diamond \overline{\delta'}) &\triangleq \text{commit}(\text{commit}(S, \delta), \overline{\delta'}) \\
\text{commit}(S, \text{write}_b\ v \mapsto z_b[z_o]) &\triangleq S(z_b).[z_o := v] \\
\text{commit}(S, \text{protect}_p) &\triangleq S[p := p] \\
\text{commit}(S, \text{new}_p\ z@z_b) &\triangleq S[z_b := \{size = z, v = \bot, p = p\}] \\
\text{commit}(S, \delta) &\triangleq S
\end{aligned}$$

$$\text{makeFrame}_{\top}(S, \overline{K} :: e, \overline{\delta}) \triangleq (S,\ \overline{K} :: e,\ \overline{\delta},\ \bullet)$$

$$\mathsf{makeFrame}_{\bot}(S, \overline{K} :: e, \overline{\delta}) \quad \triangleq \quad (S, \widehat{\overline{K} :: e}, \overline{\delta})$$

Fig. 13. Speculative semantics auxiliary functions

$$\boxed{\langle S \mid \overline{K^{\ell}} :: e^{\ell} \rangle \overset{\overline{\epsilon}}{\Rightarrow}_C \langle S \mid \overline{K^{\ell}} :: e^{\ell} \rangle}$$

$$\frac{\langle S \mid \overline{K^{\ell}} :: e^{\ell'} \rangle \overset{\epsilon}{\Rightarrow} \langle S \mid \overline{K_2{}^{\ell_2}} :: e_2{}^{\ell_2'} \rangle \qquad \overline{\delta} = [\mathsf{branch}\ v \mid \mathsf{public} \sqsubseteq S(z_b) \land z_o \in [S(z_b).size) \land v = S(z_b).v(z_o)]}{\langle S \mid \overline{K^{\ell}} :: e^{\ell'} \rangle \xRightarrow{\overline{\delta}^{\mathsf{app}} \diamond \epsilon}_C \langle S \mid \overline{K_2{}^{\ell_2}} :: e_2{}^{\ell_2'} \rangle}$$

$$\boxed{\langle \Phi \mid \overline{K} :: e \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}_C \langle \Phi \mid \overline{K} :: e \rangle}$$

$$\frac{\langle \Phi \mid \overline{K} :: e \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow} \langle \Phi' \mid \overline{K'} :: e' \rangle \qquad \overline{\delta'} = [\mathsf{branch}\ v \mid \langle \Phi[S.p := \mathtt{public}] \mid !\, z_b[z_o] \rangle \overset{\mu}{\hookrightarrow} \langle \Phi' \mid v \rangle]}{\langle \Phi \mid \overline{K} :: e \rangle \overset{\overline{\delta'} \diamond \overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}_C \langle \Phi' \mid \overline{K'} :: e' \rangle}$$

Fig. 14. Concurrent observer semantics

B Attacker models

$$\begin{array}{lcl} A & ::= & \tau^{\mathtt{lib} \to \mathsf{app}} \diamond \overline{\delta^{\mathsf{app}}} \diamond \tau^{\mathsf{app} \to \mathtt{lib}} \\ L & ::= & \tau^{\mathsf{app} \to \mathtt{lib}} \diamond \overline{\delta^{\mathtt{lib}}} \diamond \tau^{\mathtt{lib} \to \mathsf{app}} \\ T & ::= & \tau^{\mathtt{lib} \to \mathsf{app}} \diamond \overline{\delta^{\mathsf{app}}} \diamond (\mathsf{end}\ v)^{\mathsf{app} \to \mathtt{lib}} \mid A \circ L \circ T \end{array}$$

Fig. 15. Traces

$$\begin{array}{rccl} \text{API contexts} & \Gamma & ::= & \bullet \mid (f : z) :: \Gamma \\ \text{secret contexts} & \Delta & ::= & \bullet \mid x \mapsto (z_{loc}, z_{len}) :: \Delta \\ \text{libraries} & L & ::= & \bullet \mid f \mapsto (z_f, \lambda \overline{x}.e) :: L \\ \text{heaplets} & H & : & \mathbb{Z} \rightharpoonup m \end{array}$$

Fig. 16. Syntax of programs

$$\boxed{\Gamma \vDash L}$$

$$\frac{}{\bullet \vDash \bullet} \qquad \frac{\mathsf{length}(\overline{x}) = z \qquad \Gamma \vDash L \qquad z_{loc} \notin \mathsf{locs}(L)}{(f : z) :: \Gamma \vDash f \mapsto (z_f, \lambda \overline{x}.e) :: L} \qquad \frac{(f : z) :: \Gamma \vDash L \qquad z_g \notin \mathsf{locs}(L)}{(f : z) :: \Gamma \vDash g \mapsto (z_g, \lambda \overline{x}.e) :: L}$$

$$\boxed{\Gamma, \Delta \vdash (H,\ e)}$$

$$\frac{\mathsf{fv}(H) \cup \mathsf{fv}(e) \subseteq \mathsf{dom}(\Gamma) \uplus \mathsf{dom}(\Delta) \qquad \forall \lambda_{\ell} \overline{x}.e \in \mathsf{cod}(H).\ell = \mathsf{app}}{\Gamma, \Delta \vdash (H, e)}$$

$$\boxed{L \vDash S}$$

$$\frac{\forall \lambda_{\ell} \overline{x}.e \in \mathsf{cod}(S.H).\ \ell = \mathsf{app}}{\bullet \vDash S} \qquad \frac{S(z_f) = \lambda_{\mathtt{lib}} \overline{x}.e \qquad L \vDash S[H := S.H - \{z_f\}]}{f \mapsto (z_f, \lambda \overline{x}.e) :: L \vDash S}$$

$$\boxed{\Delta \mid H \vDash S = S}$$

$$\frac{S.H = H \qquad S.p = \mathsf{app}}{\bullet \mid H \vDash S = S} \qquad \frac{\begin{array}{c} S(z_{loc}).size = S'(z_{loc}).size = z_{len} \\ \Delta \mid H \vDash S[H \coloneqq S.H - \{z_{loc}\}] = S'[H \coloneqq S'.H - \{z_{loc}\}] \end{array}}{x \mapsto (z_{loc}, z_{len}) :: \Delta \mid H \vDash S = S'}$$

$$\boxed{L \mid \Delta \mid H \vDash S = S} \quad \boxed{L \mid \Delta \mid H \vDash S}$$

$$\frac{\Delta \mid H \vDash S = S'}{\bullet \mid \Delta \mid H \vDash S = S'} \qquad \frac{\begin{array}{c} S.H(z_f) = S.H'(z_f) = \lambda_{\texttt{lib}}\overline{x}.e \\ L \mid \Delta \mid H \vDash S[H \coloneqq S.H - \{z_f\}] = S'[H \coloneqq S'.H - \{z_f\}] \end{array}}{f \mapsto (z_f, \lambda\overline{x}.e) :: L \mid \Delta \mid H \vDash S = S'}$$

$$\frac{L \mid \Delta \mid H \vDash S = S}{L \mid \Delta \mid H \vdash S}$$

Fig. 17. Program well-formedness judgments

$$\frac{\langle S \mid \bullet :: e^{\mathsf{app}}\rangle \overset{T}{\Longrightarrow}{}^{*} \langle S' \mid \bullet :: v^{\mathsf{app}}\rangle}{\langle S \mid e\rangle \downarrow^{T} \langle S' \mid v\rangle} \qquad \frac{\langle S \mid \bullet :: e^{\mathsf{app}}\rangle \overset{\overline{\epsilon}}{\Longrightarrow}{}^{*}_{C} \langle S' \mid \bullet :: v^{\mathsf{app}}\rangle}{\langle S \mid e\rangle \downarrow^{\overline{\epsilon}}_{C} \langle S' \mid v\rangle}$$

$$\frac{\langle \Phi \mid \bullet :: e\rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\!\rightarrow}{}^{*} \langle \Phi' \mid \bullet :: v\rangle \qquad \Phi'.\Xi = \bullet}{\langle \Phi \mid e\rangle \downarrow^{\overline{\delta}}_{\mathcal{S}} \langle \Phi' \mid v\rangle} \qquad \frac{\langle \Phi \mid \bullet :: e\rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\!\rightarrow}{}^{*}_{C} \langle \Phi' \mid \bullet :: v\rangle \qquad \Phi'.\Xi = \bullet}{\langle \Phi \mid e\rangle \downarrow^{\overline{\delta}}_{\mathcal{S}C} \langle \Phi' \mid v\rangle}$$

$$\begin{array}{rcl}
\mathsf{traces}(\langle S \mid e\rangle) & \triangleq & \{\mathsf{begin} \diamond T \diamond \mathsf{end}\ v \mid \langle S \mid e\rangle \downarrow^{T} \langle S' \mid v\rangle\} \\
\mathsf{concurrentTraces}(\langle S \mid e\rangle) & \triangleq & \{\mathsf{begin} \diamond \overline{\epsilon} \diamond \mathsf{end}\ v \mid \langle S \mid e\rangle \downarrow^{\overline{\epsilon}}_{C} \langle S' \mid v\rangle\} \\
\mathsf{specTraces}(\langle \Phi \mid e\rangle) & \triangleq & \{\mathsf{begin} \diamond \overline{\delta} \diamond \mathsf{end}\ v \mid \langle \Phi \mid e\rangle \downarrow^{\overline{\delta}}_{\mathcal{S}} \langle \Phi' \mid v\rangle\} \\
\mathsf{concurrentSpecTraces}(\langle \Phi \mid e\rangle) & \triangleq & \{\mathsf{begin} \diamond \overline{\delta} \diamond \mathsf{end}\ v \mid \langle \Phi \mid e\rangle \downarrow^{\overline{\delta}}_{\mathcal{S}C} \langle \Phi' \mid v\rangle\}
\end{array}$$

Fig. 18. Traces

$$\begin{array}{rcl}
e[\bullet] & \triangleq & e \\
e[x \mapsto (z_b, z) :: \Delta] & \triangleq & (e[z_b/x])[\Delta] \\
H[\Delta] & \triangleq & z \mapsto H(z)[\Delta] \\
\{size, p, v\}[\Delta] & \triangleq & \{size, p, v[\Delta]\} \\
(\lambda_\ell \overline{x}.e)[\Delta] & \triangleq & \lambda_\ell \overline{x}.(e[\Delta])
\end{array}$$

$$\begin{array}{rcl}
e[\bullet] & \triangleq & e \\
e[f \mapsto (z_f, \lambda_{\texttt{lib}}\overline{x}.e) :: L] & \triangleq & (e[z_f/f])[L] \\
H[L] & \triangleq & z \mapsto H(z)[L] \\
\{size, p, v\}[L] & \triangleq & \{size, p, v[L]\} \\
(\lambda_\ell \overline{x}.e)[L] & \triangleq & \lambda_\ell \overline{x}.(e[L])
\end{array}$$

We define four attacker models: unrestricted attackers, read-only attackers, memory-safe attackers, and speculative attackers.

Definition 4 (Applications). *For a given API context* Γ *and secret context* Δ*, an application is a heaplet and an expression* (H, e) *such that* $\Gamma, \Delta \vdash (H, e)$.

Definition 5 (Read-only attackers). *We say an application* $\Gamma, \Delta \vdash (H, e)$ *is a* read-only attacker *if, for all libraries* $\Gamma \vDash L$*, initial states* S_0 *such that* $L \mid \Delta \mid H \vDash S_0$*, and traces* $T \in \mathsf{traces}(\langle S_0[\Delta][L] \mid e[\Delta][L]\rangle)$, $\Gamma \vdash \mathsf{read\text{-}only}\ T$.

Definition 6 (Memory-safe attackers). *We say an application* $\Gamma, \Delta \vdash (H, e)$ *is a* read-only attacker *if, for all libraries* $\Gamma \vDash L$*, initial states* S_0 *such that* $L \mid \Delta \mid H \vDash S_0$*, and traces* $T \in \mathsf{traces}(\langle S_0[\Delta][L] \mid e[\Delta][L]\rangle)$, $\Gamma \vdash \mathsf{mem\text{-}safe}\ T$.

To capture speculative attackers we consider (non-speculatively) memory safe attackers run in the speculative semantics. To capture concurrent observers we consider read-only attackers' concurrent traces.

$$\boxed{\Gamma \vdash \mathsf{read\text{-}only}\ T}$$

$$\frac{\begin{array}{c} A = {\tau_1}^{\mathtt{lib}\to\mathsf{app}} \diamond \overline{\delta_A}^{\mathsf{app}} \diamond {\tau_2}^{\mathsf{app}\to\mathtt{lib}} \qquad L = {\tau_2}^{\mathsf{app}\to\mathtt{lib}} \diamond \overline{\delta_L}^{\mathtt{lib}} \diamond {\tau_3}^{\mathtt{lib}\to\mathsf{app}} \\ \tau_2 = \mathsf{call}\ z_f \Rightarrow z_f \in \mathsf{dom}(\Gamma) \qquad \mathsf{wf\text{-}read\text{-}only}\ \overline{\delta_A} \qquad \mathsf{wf\text{-}read\text{-}only}\ \overline{\delta_L} \implies \Gamma \vdash \mathsf{read\text{-}only}\ T \end{array}}{\Gamma \vdash \mathsf{read\text{-}only}\ A \circ L \circ T}$$

$$\frac{\mathsf{wf\text{-}read\text{-}only}\ \overline{\delta}}{\Gamma \vdash \mathsf{read\text{-}only}\ \tau^{\mathtt{lib}\to\mathsf{app}} \diamond \overline{\delta}^{\mathsf{app}} \diamond (\mathsf{end}\ v)^{\mathsf{app}\to\mathtt{lib}}}$$

$$\boxed{\Gamma \vdash \mathsf{mem\text{-}safe}\ T}$$

$$\frac{\begin{array}{c} A = {\tau_1}^{\mathtt{lib}\to\mathsf{app}} \diamond \overline{\delta_A}^{\mathsf{app}} \diamond {\tau_2}^{\mathsf{app}\to\mathtt{lib}} \qquad L = {\tau_2}^{\mathsf{app}\to\mathtt{lib}} \diamond \overline{\delta_L}^{\mathtt{lib}} \diamond {\tau_3}^{\mathtt{lib}\to\mathsf{app}} \\ \tau_2 = \mathsf{call}\ z_f \Rightarrow z_f \in \mathsf{dom}(\Gamma) \qquad \mathsf{wf\text{-}mem\text{-}safe}\ \overline{\delta_A} \qquad \mathsf{wf\text{-}read\text{-}only}\ \overline{\delta_L} \implies \Gamma \vdash \mathsf{read\text{-}only}\ T \end{array}}{\Gamma \vdash \mathsf{mem\text{-}safe}\ A \circ L \circ T}$$

$$\frac{\mathsf{wf\text{-}mem\text{-}safe}\ \overline{\delta}}{\Gamma \vdash \mathsf{mem\text{-}safe}\ \tau^{\mathtt{lib}\to\mathsf{app}} \diamond \overline{\delta}^{\mathsf{app}} \diamond (\mathsf{end}\ v)^{\mathsf{app}\to\mathtt{lib}}}$$

$$\boxed{\mathsf{wf\text{-}read\text{-}only}\ \overline{\delta}}$$

$$\frac{}{\mathsf{wf\text{-}read\text{-}only}\ \mathsf{new}_p\ z@z_b} \qquad \frac{}{\mathsf{wf\text{-}read\text{-}only}\ \mathsf{read}_b\ v \hookleftarrow z_b[z_o]}$$

$$\frac{}{\mathsf{wf\text{-}read\text{-}only}\ \mathsf{write}_{\mathsf{ib}}\ v \mapsto z_b[z_o]} \qquad \frac{}{\mathsf{wf\text{-}read\text{-}only}\ \mathsf{call}\ z} \qquad \frac{}{\mathsf{wf\text{-}read\text{-}only}\ \mathsf{branch}\ v}$$

$$\frac{}{\mathsf{wf\text{-}read\text{-}only}\ \mathsf{protect}_p} \qquad \frac{}{\mathsf{wf\text{-}read\text{-}only}\ 0} \qquad \frac{}{\mathsf{wf\text{-}read\text{-}only}\ \emptyset}$$

$$\frac{\mathsf{wf\text{-}read\text{-}only}\ \delta \qquad \mathsf{wf\text{-}read\text{-}only}\ \overline{\delta}}{\mathsf{wf\text{-}read\text{-}only}\ \delta \diamond \overline{\delta}}$$

$\boxed{\text{wf-mem-safe } \overline{\delta}}$

$$\frac{}{\text{wf-mem-safe new}_p\ z@z_b} \qquad \frac{}{\text{wf-mem-safe read}_{\text{ib}}\ v \hookleftarrow z_b[z_o]}$$

$$\frac{}{\text{wf-mem-safe write}_{\text{ib}}\ v \mapsto z_b[z_o]} \qquad \frac{}{\text{wf-mem-safe call } z} \qquad \frac{}{\text{wf-mem-safe branch } v}$$

$$\frac{}{\text{wf-mem-safe protect}_p} \qquad \frac{}{\text{wf-mem-safe } 0} \qquad \frac{}{\text{wf-mem-safe } \emptyset}$$

$$\frac{\text{wf-mem-safe } \delta \qquad \text{wf-mem-safe } \overline{\delta}}{\text{wf-mem-safe } \delta \diamond \overline{\delta}}$$

Fig. 19. Attacker model judgments

B.1 Security definitions

$$\begin{array}{rcl}
\text{ct}(\epsilon^{\ell}) & \triangleq & \text{ct}(\epsilon) \\
\text{ct}(\tau^{\ell \to \ell'}) & \triangleq & \text{ct}(\tau) \\
\text{ct}(\text{begin}) & \triangleq & 0 \\
\text{ct}(\text{end } v) & \triangleq & \text{end } v \\
\text{ct}(\text{call } z) & \triangleq & \text{call } z \\
\text{ct}(\text{ret } v) & \triangleq & 0 \\
\text{ct}(\text{new}_p\ z@z_b) & \triangleq & \text{new}_p\ z@z_b \\
\text{ct}(\text{read}_b\ v \hookleftarrow z_b[z_o]) & \triangleq & \text{read} \hookleftarrow z_b[z_o] \\
\text{ct}(\text{write}_b\ v \mapsto z_b[z_o]) & \triangleq & \text{write} \mapsto z_b[z_o] \\
\text{ct}(\text{branch } v) & \triangleq & \text{branch } v \\
\text{ct}(\text{fence}) & \triangleq & 0 \\
\text{ct}(\text{protect}_p) & \triangleq & 0 \\
\text{ct}(0) & \triangleq & 0
\end{array}$$

Fig. 20. Constant time events

$e_\Gamma ::= f(\overline{v})$ with $f \in \Gamma$.

Definition 7 (Classical constant time). *We say a library* $\Gamma \vDash L$ *is* classically constant time *if, for all secret contexts* Δ*, classical "applications"* $\Gamma, \Delta \vdash (H, e_\Gamma)$*, and initial states* S_0, S_0' *such that* $L \mid \Delta \mid H \vDash S_0 = S_0'$*, we have that for all traces* $\langle S_0[\Delta][L] \mid e_\Gamma[\Delta][L]\rangle \downarrow^{\overline{\delta}} \langle S_1 \mid v\rangle$ *there exists a trace* $\langle S_0'[\Delta][L] \mid e_\Gamma[\Delta][L]\rangle \downarrow^{\overline{\delta'}} \langle S_1' \mid v\rangle$ *such that* $\text{ct}(\overline{\delta}) = \text{ct}(\overline{\delta'})$ *and* $S_1(\text{dom}(H)) = S_1'(\text{dom}(H))$.

Definition 8 (Robust constant time). *We say a library* $\Gamma \vDash L$ *is* robustly constant time *for an attacker class* $\text{pred}_{\Gamma,\Delta}$ *if, for all secret contexts* Δ*, applications* $\Gamma, \Delta \vdash (H, e)$ *such that* $\text{pred}_{\Gamma,\Delta}(H, e)$*, and initial states* S_0, S_0' *such that* $L \mid \Delta \mid H \vDash S_0 = S_0'$ *we have that* $\text{ct}(\text{traces}(\langle S_0[\Delta][L] \mid e[\Delta][L]\rangle)) = \text{ct}(\text{traces}(\langle S_0'[\Delta][L] \mid e[\Delta][L]\rangle))$.

Definition 9 (Robust constant time for concurrent observers). *We say a library* $\Gamma \vDash L$ *is* robustly constant time for concurrent observers *if, for all secret contexts* Δ*, read-only applications* $\Gamma, \Delta \vdash (H, e)$*, and initial states* S_0*,* S'_0 *such that* $L \mid \Delta \mid H \vDash S_0 = S'_0$ *we have that* $\mathsf{ct}(\mathsf{concurrentTraces}(\langle S_0[\Delta][L] \mid e[\Delta][L]\rangle)) = \mathsf{ct}(\mathsf{concurrentTraces}(\langle S'_0[\Delta][L] \mid e[\Delta][L]\rangle))$.

Definition 10 (Classical speculative constant time). *We say a library* $\Gamma \vDash L$ *is* classically speculative constant time *with respect to a speculation oracle* $\mathsf{spec} : A \times S \times e \to A \times d$ *if, for all secret contexts* Δ*, classical "applications"* $\Gamma, \Delta \vdash (H, e_\Gamma)$*, initial states* S_0*,* S'_0 *such that* $L \mid \Delta \mid H \vDash S_0 = S'_0$*, microarchitectural states* $a : A$*,* $\Phi_0 = \{S = S_0[\Delta][L], a = a, \Xi = \bullet\}$*, and* $\Phi'_0 = \{S = S'_0[\Delta][L], a = a, \Xi = \bullet\}$*, we have that for all traces* $\langle \Phi_0 \mid e_\Gamma[\Delta][L]\rangle \downarrow_S^{\overline{\delta}} \langle \Phi_1 \mid v\rangle$ *there exists a trace* $\langle \Phi'_0 \mid e_\Gamma[\Delta][L]\rangle \downarrow_S^{\overline{\delta'}} \langle \Phi'_1 \mid v\rangle$ *such that* $\mathsf{ct}(\overline{\delta}) = \mathsf{ct}(\overline{\delta'})$*,* $\Phi_1.S(\mathsf{dom}(H)) = \Phi'_1.S(\mathsf{dom}(H))$*, and* $\Phi_1.a = \Phi'_1.a$*.*

Definition 11 (Robust speculative constant time). *We say a library* $\Gamma \vDash L$ *is* robustly speculative constant time *with respect to a speculation oracle* $\mathsf{spec} : A \times S \times e \to A \times d$ *if, for all secret contexts* Δ*, memory-safe applications* $\Gamma, \Delta \vdash (H, e)$*, initial states* S_0*,* S'_0 *such that* $L \mid \Delta \mid H \vDash S_0 = S'_0$*, microarchitectural states* $a : A$*,* $\Phi_0 = \{S = S_0[\Delta][L], a = a, \Xi = \bullet\}$*, and* $\Phi'_0 = \{S = S'_0[\Delta][L], a = a, \Xi = \bullet\}$*, we have that* $\mathsf{ct}(\mathsf{specTraces}(\langle \Phi_0 \mid e[\Delta][L]\rangle)) = \mathsf{ct}(\mathsf{specTraces}(\langle \Phi'_0 \mid e[\Delta][L]\rangle))$.

Definition 12 (Robust speculative constant time for concurrent observers). *We say a library* $\Gamma \vDash L$ *is* robustly speculative constant time for concurrent observers *with respect to a speculation oracle* $\mathsf{spec} : A \times S \times e \to A \times d$ *if, for all secret contexts* Δ*, memory-safe applications* $\Gamma, \Delta \vdash (H, e)$*, initial states* S_0*,* S'_0 *such that* $L \mid \Delta \mid H \vDash S_0 = S'_0$*, microarchitectural states* $a : A$*,* $\Phi_0 = \{S = S_0[\Delta][L], a = a, \Xi = \bullet\}$*, and* $\Phi'_0 = \{S = S'_0[\Delta][L], a = a, \Xi = \bullet\}$*, we have that* $\mathsf{ct}(\mathsf{concurrentSpecTraces}(\langle \Phi_0 \mid e[\Delta][L]\rangle)) = \mathsf{ct}(\mathsf{concurrentSpecTraces}(\langle \Phi'_0 \mid e[\Delta][L]\rangle))$.

C Compilers

$$
\begin{aligned}
\text{on-return}_f(v) &= f(v)\\
\text{on-return}_f(x) &= f(x)\\
\text{on-return}_f(e_1; e_2) &= e_1; \text{on-return}_f(e_2)\\
\text{on-return}_f(\texttt{fence}) &= f(\texttt{fence})\\
\text{on-return}_f(\texttt{if } e \texttt{ then } e_1 \texttt{ else } e_2) &= \texttt{if } e \texttt{ then } \text{on-return}_f(e_1) \texttt{ else } \text{on-return}_f(e_2)\\
\text{on-return}_f(e) &= \texttt{let } x = e \texttt{ in } f(x)
\end{aligned}
$$

$$
\begin{aligned}
\text{add-protect}(e) &= \texttt{protect}_{\texttt{public}}; e\\
\text{add-fence}(e) &= \texttt{fence}; e\\[1em]
\text{copy-in}(\bullet, e) &= e\\
\text{copy-in}(f \mapsto \overline{(x,\ e_{len})} :: rest, e) &= \texttt{let } y = \mathsf{new}_{\mathsf{public}}\ e_{len} \texttt{ in } \mathsf{copy}(y, x, e_{len}); \text{copy-in}(e[y/x])\\[1em]
\text{copy-out}(\bullet, e) &= e\\
\text{copy-out}(f \mapsto \overline{(x,\ e_{len})} :: rest, e) &= \mathsf{copy}(x, y, e_{len}); \text{copy-out}(e)
\end{aligned}
$$

$$
\begin{array}{rcl}
\mathbb{C}_{\text{ro}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) & = & \lambda\overline{x}.e_{new} \\
 & & \quad \text{where } e_{subst} = e[\mathsf{new}_{\mathsf{protected}}\ e'/\mathsf{new}_p\ e'] \\
 & & \quad \text{where } e_{body} = \mathsf{on\text{-}return}_{\mathsf{add\text{-}protect}}(e_{subst}) \\
 & & \quad \text{where } e_{new} = \mathtt{protect}_{\mathsf{protected}}; e_{body} \\
 & & \quad \text{if } f \in \Gamma \\
\mathbb{C}_{\text{ro}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) & = & \lambda\overline{x}.(e[\mathsf{new}_{\mathsf{protected}}\ e'/\mathsf{new}_p\ e']) \\
 & & \quad \text{if } f \notin \Gamma
\end{array}
$$

$$
\begin{array}{rcl}
\mathbb{C}_{\text{spec}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) & = & \lambda\overline{x}.\mathsf{on\text{-}return}_{\mathsf{add\text{-}fence}}(e_{new}) \\
 & & \quad \text{where } \lambda\overline{x}.e_{new} = \mathbb{C}_{\text{ro}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) \\
 & & \quad \text{if } f \in \Gamma \\
\mathbb{C}_{\text{spec}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) & = & \mathbb{C}_{\text{ro}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) \\
 & & \quad \text{if } f \notin \Gamma
\end{array}
$$

$$
\begin{array}{rcl}
\mathbb{C}_{\text{co}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) & = & \lambda\overline{x}.\mathsf{on\text{-}return}_{\mathsf{copy\text{-}out}(used)}(e_{new}) \\
 & & \quad \text{where } e_{new} = \mathsf{copy\text{-}in}(used, e) \\
 & & \quad \text{if } f \in \Gamma \\
\mathbb{C}_{\text{co}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) & = & f \mapsto (z_f,\ \lambda\overline{x}.e) \\
 & & \quad \text{if } f \notin \Gamma
\end{array}
$$

$$
\begin{array}{rcl}
\mathbb{C}_{\text{ro-co}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) & = & \mathbb{C}_{\text{ro}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e_{new}) \\
 & & \quad \text{where } \lambda\overline{x}.e_{new} = \mathbb{C}_{\text{co}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) \\
 & & \quad \text{if } f \in \Gamma \\
\mathbb{C}_{\text{ro-co}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) & = & \mathbb{C}_{\text{ro}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) \\
 & & \quad \text{if } f \notin \Gamma
\end{array}
$$

$$
\begin{array}{rcl}
\mathbb{C}_{\text{spec-co}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) & = & \mathbb{C}_{\text{spec}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e_{new}) \\
 & & \quad \text{where } \lambda\overline{x}.e_{new} = \mathbb{C}_{\text{co}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) \\
 & & \quad \text{if } f \in \Gamma \\
\mathbb{C}_{\text{spec-co}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) & = & \mathbb{C}_{\text{spec}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e)) \\
 & & \quad \text{if } f \notin \Gamma
\end{array}
$$

$$
\begin{array}{rcl}
\mathbb{C}_{\mathsf{A}}(\Gamma \vDash \bullet) & = & \bullet \\
\mathbb{C}_{\mathsf{A}}(\Gamma \vDash f \mapsto (z_f,\ \lambda\overline{x}.e) :: L) & = & f \mapsto (z_f,\ \mathbb{C}_{\mathsf{A}}(\Gamma, f \mapsto (z_f,\ \lambda\overline{x}.e))) :: \mathbb{C}_{\mathsf{A}}(\Gamma \vDash L)
\end{array}
$$

For the concurrent compilers we assume we have been provided a labeling of arguments $\overline{used} :$ $f \mapsto \overline{(x,\ e_{len})}$ for all functions in Γ. The variables are the argument names and the expressions are the length of the buffer. The correctness condition for the labeling is the following: For all secret contexts Δ, classical "applications" $\Gamma, \Delta \vdash (H,\ e_\Gamma)$, and initial states $\mathbb{C}_{\text{co}}(\Gamma \vDash L) \mid \Delta \mid H \vDash S$, if $T \in \mathsf{traces}(\langle S[\Delta][\mathbb{C}_{\text{co}}(\Gamma \vDash L)] \mid e_\Gamma[\Delta][\mathbb{C}_{\text{co}}(\Gamma \vDash L)]\rangle)$ and $\overline{\mu}$ is the set of memory events in T, then $\overline{\mu} = \overline{\mu_1} \diamond \overline{\mu_2} \diamond \overline{\mu_3}$ where

(1) $\overline{\mu_1} = \overline{\mathsf{new}_{\mathtt{public}}\ v_{len}@z'_b \diamond \mathsf{read}_{\mathsf{ib}}\ v \hookleftarrow z_b[z_o] \diamond \mathsf{write}_{\mathsf{ib}}\ v \mapsto z'_b[z_o]}$ and $z_b \in \mathsf{dom}(H)$
(2) for all $\mu_2 = \mathsf{read}_b\ v \hookleftarrow z_b[z_o]$ or $\mu_2 = \mathsf{write}_b\ v \mapsto z_b[z_o]$, $z_b \notin \mathsf{dom}(H)$
(3) $\overline{\mu_3} = \overline{\mathsf{read}_{\mathsf{ib}}\ v \hookleftarrow z'_b[z_o] \diamond \mathsf{write}_{\mathsf{ib}}\ v \mapsto z_b[z_o]}$

and if $\langle S[\Delta][L] \mid e_\Gamma[\Delta][L]\rangle \downarrow^{T_1} \langle S' \mid v\rangle$ then there exists a T_1' such that $\langle S[\Delta][\mathbb{C}_{co}(\Gamma \vDash L)] \mid e_\Gamma[\Delta][\mathbb{C}_{co}(\Gamma \vDash L)]\rangle \downarrow^{T_1} \langle S'' \mid v\rangle$.

D Compiler proofs

$$\boxed{\Delta \mid H \vDash_{\texttt{protected}} S = S}$$

$$\frac{S.H = H \qquad S.p = \texttt{app}}{\bullet \mid H \vDash_{\texttt{protected}} S = S} \qquad \frac{\begin{array}{c} S(z_{loc}).size = S'(z_{loc}).size = z_{len} \\ S(z_{loc}).p = S'(z_{loc}).p = \texttt{protected} \\ \Delta \mid H \vDash_{\texttt{protected}} S[H \coloneqq S.H - \{z_{loc}\}] = S'[H \coloneqq S'.H - \{z_{loc}\}] \end{array}}{x \mapsto (z_{loc}, z_{len}) :: \Delta \mid H \vDash_{\texttt{protected}} S = S'}$$

D.1 Read-only protections

Lemma 1 (Functions are fixed). *For all* $\langle S_1 \mid \overline{K_1{}^{\ell_{K1}}} :: e_1{}^{\ell_1}\rangle \overset{\bar{\epsilon}}{\Rightarrow}{}^* \langle S_2 \mid \overline{K_2{}^{\ell_{K2}}} :: e_2{}^{\ell_2}\rangle$,

$$\left\{z_f \mapsto \lambda_\ell \overline{x}.e \mid S_1(z_f) = \lambda_\ell \overline{x}.e\right\} = \left\{z_f \mapsto \lambda_\ell \overline{x}.e \mid S_2(z_f) = \lambda_\ell \overline{x}.e\right\}.$$

Proof. By induction on the operational semantics. □

Lemma 2 (Constant time is memory safe). *If* $\Gamma \vDash L$ *is classically constant time and we have a secret context* Δ*, classical "application"* $\Gamma, \Delta \vdash (H, e_\Gamma)$ *and an initial state* $L \mid \Delta \mid H \vdash S$*, then for all* $T \in \text{trace}(\langle S[\Delta][L] \mid e[\Delta][L]\rangle)$ *and* $\delta \in T$*,* wf-mem-safe δ.

Proof. By induction on the operational semantics. □

Lemma 3 ($\mathbb{C}_{ro}$ preserves classic constant time). *If* $\Gamma \vDash L$ *is classically constant time then* $\mathbb{C}_{ro}\Gamma \vDash L$ *is classically constant time.*

Proof. By induction on the compiler. □

Lemma 4 ($\mathbb{C}_{ro}$ only allocates protected memory). *If* $\Gamma \vDash L$ *is classically constant time and we have a secret context* Δ*, classical "application"* $\Gamma, \Delta \vdash (H, e_\Gamma)$ *and an initial state* $L \mid \Delta \mid H \vdash S$*, then if* $\langle S[\Delta][L] \mid e[\Delta][L]\rangle \downarrow^T \langle S' \mid v\rangle$*, then* $\{S(z_b) \mid S(z_b).p = \texttt{public}\} = \{S'(z_b) \mid S'(z_b).p = \texttt{public}\}$

Proof. By induction on the compiler and the assumption of classical constant time. □

Definition 13. *We define our state invariant,* $\text{inv}(S, S')$ *as follows:*

(1) $S.p = S'.p = \texttt{public}$
(2) $\forall z : \mathbb{Z}.S(z) \neq S'(z) \Rightarrow S(z).p = S'(z).p = \texttt{protected}$
(3) $\text{dom}(S) = \text{dom}(S')$

$$\mathbb{A}\llbracket \tau^{\texttt{lib}\to\mathsf{app}} \diamond \overline{\delta^{\mathsf{app}}} \diamond \tau^{\mathsf{app}\to\texttt{lib}} \rrbracket \triangleq \left\{ (S,\, S',\, \overline{K^{\ell}},\, e,\, S_1,\, S_1',\, \overline{K_1{}^{\ell_1}},\, e_1) \;\middle|\; \begin{array}{l} \forall \overline{\epsilon_1}, \overline{\epsilon_2}.\ \overline{\epsilon_1} \diamond \overline{\epsilon_2} = \overline{\delta^{\mathsf{app}}} \Rightarrow \\ \quad \exists S_0, S_0', \overline{K_0{}^{\ell_0}}, e_0. \\ \qquad \langle S \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \xRightarrow{\overline{\epsilon_1}}{}^{*} \langle S_0 \mid \overline{K_0{}^{\ell_0}} :: e_0{}^{\mathsf{app}} \rangle \\ \qquad \langle S' \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \xRightarrow{\overline{\epsilon_1}}{}^{*} \langle S_0' \mid \overline{K_0{}^{\ell_0}} :: e_0{}^{\mathsf{app}} \rangle \\ \qquad \mathsf{inv}(S_0, S_0') \\ \langle S \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \xRightarrow{\overline{\delta^{\mathsf{app}}} \diamond \tau^{\mathsf{app}\to\texttt{lib}}}{}^{*} \langle S_1 \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\texttt{lib}} \rangle \\ \langle S' \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \xRightarrow{\overline{\delta^{\mathsf{app}}} \diamond \tau^{\mathsf{app}\to\texttt{lib}}}{}^{*} \langle S_1' \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\texttt{lib}} \rangle \\ \mathsf{inv}(S_1, S_1') \end{array} \right\}$$

$$\mathbb{L}\llbracket \tau^{\mathsf{app}\to\texttt{lib}} \diamond \overline{\delta^{\texttt{lib}}} \diamond \tau^{\texttt{lib}\to\mathsf{app}} \rrbracket \triangleq \left\{ (S,\, S',\, \overline{K^{\ell}},\, e,\, S_1,\, S_1',\, \overline{K_1{}^{\ell_1}},\, e_1) \;\middle|\; \begin{array}{l} \exists \delta'. \\ \quad \langle S \mid \overline{K^{\ell}} :: e^{\texttt{lib}} \rangle \xRightarrow{\overline{\delta^{\texttt{lib}}} \diamond \tau^{\texttt{lib}\to\mathsf{app}}}{}^{*} \langle S_1 \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\mathsf{app}} \rangle \\ \quad \langle S' \mid \overline{K^{\ell}} :: e^{\texttt{lib}} \rangle \xRightarrow{\overline{\delta'^{\texttt{lib}}} \diamond \tau^{\texttt{lib}\to\mathsf{app}}}{}^{*} \langle S_1' \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\mathsf{app}} \rangle \\ \quad \mathsf{ct}(\overline{\delta^{\texttt{lib}}}) = \mathsf{ct}(\overline{\delta'^{\texttt{lib}}}) \\ \quad \mathsf{inv}(S_1, S_1') \end{array} \right\}$$

$$\mathbb{T}\llbracket A \circ L \circ T \rrbracket \triangleq \left\{ (\langle S \mid \overline{K^{\ell}} :: e \rangle,\, \langle S' \mid \overline{K^{\ell}} :: e \rangle) \;\middle|\; \begin{array}{l} \exists S_1, S_1', S_2, S_2', \overline{K_1{}^{\ell_1}}, e_1, \overline{K_2{}^{\ell_2}}, e_2. \\ \quad (S,\, S',\, \overline{K^{\ell}},\, e,\, S_1,\, S_1',\, \overline{K_1{}^{\ell_1}},\, e_1) \in \mathbb{A}\llbracket A \rrbracket \\ \quad (S_1,\, S_1',\, \overline{K_1{}^{\ell_1}},\, e_1,\, S_2,\, S_2',\, \overline{K_2{}^{\ell_2}},\, e_1) \in \mathbb{L}\llbracket L \rrbracket \\ \quad (\langle S_2 \mid \overline{K_2{}^{\ell_2}} :: e_2{}^{\mathsf{app}} \rangle,\, \langle S_2' \mid \overline{K_2{}^{\ell_2}} :: e_2{}^{\mathsf{app}} \rangle)) \in \mathbb{T}\llbracket T \rrbracket \end{array} \right\}$$

$$\mathbb{T}\llbracket \tau^{\texttt{lib}\to\mathsf{app}} \diamond \overline{\delta^{\mathsf{app}}} \diamond (\mathsf{end}\ v)^{\mathsf{app}\to\texttt{lib}} \rrbracket \triangleq \left\{ (\langle S \mid \overline{K^{\ell}} :: e \rangle,\, \langle S' \mid \overline{K^{\ell}} :: e \rangle) \;\middle|\; \begin{array}{l} \exists S_1, S_1', \overline{K_1{}^{\ell_1}}. \\ \quad (S,\, S',\, \overline{K^{\ell}},\, e,\, S_1,\, S_1',\, \overline{K_1{}^{\ell_1}},\, v) \in \mathbb{A}\llbracket \overline{\delta^{\mathsf{app}}} \diamond (\mathsf{end}\ v)^{\mathsf{app}\to\texttt{lib}} \rrbracket \end{array} \right\}$$

Fig. 21. Semantic interpretation of non-speculative traces

Lemma 5. *If* $\langle S \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \xRightarrow{\overline{\delta^{\mathsf{app}}}}{}^{*} \langle S_1 \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\mathsf{app}} \rangle$, wf-read-only $\overline{\delta^{\mathsf{app}}}$, *and* $\mathsf{inv}(S, S')$, *then there exists an* S_1' *such that* $\langle S' \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \xRightarrow{\overline{\delta^{\mathsf{app}}}}{}^{*} \langle S_1' \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\mathsf{app}} \rangle$ *and* $\mathsf{inv}(S_1, S_1')$.

Proof. We proceed by induction on the reduction relation. The zero step case is immediate by assumption. For the inductive case we have that $\langle S \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \xRightarrow{\overline{\delta_0{}^{\mathsf{app}}}}{}^{*} \langle S_0 \mid \overline{K_0{}^{\ell_0}} :: e_0{}^{\mathsf{app}} \rangle \xRightarrow{\delta^{\mathsf{app}}} \langle S_1 \mid$

$\overline{K_1}^{\ell_1} :: {e_1}^{\mathsf{app}}\rangle$ and must show that $\langle S' \mid \overline{K^\ell} :: e^{\mathsf{app}}\rangle \xRightarrow{\overline{\delta_0}^{\mathsf{app}} \diamond \delta^{\mathsf{app}}}{}^* \langle S_1 \mid \overline{K_1}^{\ell_1} :: {e_1}^{\mathsf{app}}\rangle$ such that $\mathsf{inv}(S_1, S_1')$.

By our inductive hypothesis we have that there exists an S_0' such that $\langle S' \mid \overline{K^\ell} :: e^{\mathsf{app}}\rangle \xRightarrow{\overline{\delta_0}^{\mathsf{app}}}{}^* \langle S_0 \mid \overline{K_0}^{\ell_0} :: {e_0}^{\mathsf{app}}\rangle$ and $\mathsf{inv}(S_0, S_0')$. We proceed by case analysis on δ.

The only cases that interact with the state (the only parts that differ) are dereferencing, writes, allocation, and protect. In all other cases we let $S_1' = S_0'$ and then it is immediate that $\langle S_0' \mid \overline{K_0}^{\ell_0} :: {e_0}^{\mathsf{app}}\rangle \xRightarrow{\delta^{\mathsf{app}}} \langle S_0' \mid \overline{K_1}^{\ell_1} :: {e_1}^{\mathsf{app}}\rangle$ and the invariant holds by assumption.

In the case that $\delta = \mathsf{read}_{z_b[z_o]}\ b \hookleftarrow v$ we have, by Conditions 1 and 2, that for all locations z_b' such that $\mathsf{accessible}(S_0', z_b')$, $S_0(z_b') = S_0'(z_b')$. By Condition 3 we then have that $\{z_b' \mid \mathsf{accessible}(S_0, z_b')\} = \{z_b' \mid \mathsf{accessible}(S_0', z_b')\}$. Therefore the dereference can take the same step $\mathsf{read}_{z_b[z_o]}\ b \hookleftarrow v$ under S_0' and we let $S_1' = S_0'$.

The case for $\delta = \mathsf{write}_{z_b[z_o]}\ b \mapsto v$ is identical except that we let $S_1' = S_0'(z_b).v[z_0 := v]$. By Conditions 1 and 2 we have that $S_0'(z_b).p = \mathtt{public}$ and therefore we get that $\mathsf{inv}(S_1, S_1')$.

For $\delta = \mathsf{new}_p\ z_{len}@z_b$ we rely on Condition 3, which means that z_b is also fresh in S_0'.

We have by assumption that the application does not contain protect statements so $\delta = \mathsf{protect}_p$ is a contradiction. □

Lemma 6. *For a classically constant time library* $\Gamma \vDash L$, *if* $\mathbb{C}_{\mathrm{ro}}(\Gamma \vDash L) \vDash S$, $\mathbb{C}_{\mathrm{ro}}(\Gamma \vDash L) \vDash S'$, $T \in \mathsf{traces}(\langle S \mid \overline{K^\ell} :: e^{\mathsf{app}}\rangle)$, $\Gamma \vdash \mathsf{read\text{-}only}\ T$, *and* $\mathsf{inv}(S, S')$, *then* $(\langle S \mid \overline{K^\ell} :: e\rangle, \langle S' \mid \overline{K^\ell} :: e\rangle) \in \mathbb{T}[\![T]\!]$.

Proof. We proceed by induction on T. Wolog we show the case for $T = A \circ L \circ T'$ (the $T = \tau^{\mathtt{lib}\to\mathsf{app}} \diamond \overline{\delta^{\mathsf{app}}} \diamond (\mathsf{end}\ v)^{\mathsf{app}\to\mathtt{lib}}$ case follows identical reasoning as the A subtrace.) We split the proof into instantiating the different subtrace relations.

Case $\exists S_1, S_1', \overline{K_1}^{\ell_1}, e_1.\ (S,\ S',\ \overline{K^\ell},\ e,\ S_1,\ S_1',\ \overline{K_1}^{\ell_1},\ e_1) \in \mathbb{A}[\![A]\!]$**:** By definition we have that $A = \tau^{\mathtt{lib}\to\mathsf{app}} \diamond \overline{\delta^{\mathsf{app}}} \diamond \tau^{\mathsf{app}\to\mathtt{lib}}$. By assumption we have that for all $\overline{\epsilon_1} \diamond \overline{\epsilon_2} = \overline{\delta^{\mathsf{app}}}$ there exists some S_0, $\overline{K_0}^{\ell_0}$, and e_0 such that $\langle S \mid \overline{K^\ell} :: e^{\mathsf{app}}\rangle \xRightarrow{\overline{\epsilon_1}}{}^* \langle S_0 \mid \overline{K_0}^{\ell_0} :: {e_0}^{\mathsf{app}}\rangle$. Lemma 5 then guarantees that the prefix conditions hold.

To show that the overall application trace conditions hold let $\overline{\epsilon_1} = \overline{\delta^{\mathsf{app}}}$. By the above logic we have that there exist S_0, $\overline{K_0}^{\ell_0}$, and e_0 such that $\langle S \mid \overline{K^\ell} :: e^{\mathsf{app}}\rangle \xRightarrow{\overline{\delta^{\mathsf{app}}}}{}^* \langle S_0 \mid \overline{K_0}^{\ell_0} :: {e_0}^{\mathsf{app}}\rangle$, $\langle S' \mid \overline{K^\ell} :: e^{\mathsf{app}}\rangle \xRightarrow{\overline{\delta^{\mathsf{app}}}}{}^* \langle S_0' \mid \overline{K_0}^{\ell_0} :: {e_0}^{\mathsf{app}}\rangle$, and $\mathsf{inv}(S_0, S_0')$. By assumption we have that there exists some S_1, $\overline{K_1}^{\ell_1}$, and e_1 such that $\langle S_0 \mid \overline{K_0}^{\ell_0} :: {e_0}^{\mathsf{app}}\rangle \xRightarrow{\tau^{\mathtt{lib}\to\mathsf{app}}} \langle S_1 \mid \overline{K_1}^{\ell_1} :: {e_1}^{\mathtt{lib}}\rangle$. By $\Gamma \vdash \mathsf{read\text{-}only}\ T$ we have that $\tau = \mathsf{call}\ z_f$ and $z_f \in \mathsf{dom}(\Gamma)$. By inversion $e_0 = z_f(\overline{v})$. By Lemma 1, $\mathbb{C}_{\mathrm{ro}}(\Gamma \vDash L) \vDash S$, and $\mathbb{C}_{\mathrm{ro}}(\Gamma \vDash L) \vDash S$ we have that $S(z_f) = S(z_f')$. The remaining conditions of $\mathbb{A}$ then follow immediately. ■

Case $\exists S_2, S_2', \overline{K_2}^{\ell_2}, e_2.\ (S_1,\ S_1',\ \overline{K_1}^{\ell_1},\ e_1,\ S_2,\ S_2',\ \overline{K_2}^{\ell_2},\ e_2) \in \mathbb{L}[\![L]\!]$**:** By definition we have that $L = \tau^{\mathsf{app}\to\mathtt{lib}} \diamond \overline{\delta^{\mathtt{lib}}} \diamond {\tau_2}^{\mathtt{lib}\to\mathsf{app}}$. By the above we have that

$$\langle S \mid \overline{K^\ell} :: e^{\mathsf{app}}\rangle \xRightarrow{\overline{\delta^{\mathsf{app}}} \diamond \tau^{\mathsf{app}\to\mathtt{lib}}}{}^* \langle S_1 \mid \overline{K_1}^{\ell_1} :: {e_1}^{\mathtt{lib}}\rangle$$

and

$$\langle S' \mid \overline{K^\ell} :: e^{\mathsf{app}}\rangle \xRightarrow{\overline{\delta^{\mathsf{app}}} \diamond \tau^{\mathsf{app}\to\mathtt{lib}}}{}^* \langle S_1' \mid \overline{K_1}^{\ell_1} :: {e_1}^{\mathtt{lib}}\rangle.$$

By assumption we have that $\langle S_1 \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\texttt{lib}}\rangle \xRightarrow{\overline{\delta^{\texttt{lib}}} \diamond \tau_2{}^{\texttt{lib}\to\texttt{app}}}{}^* \langle S_2 \mid \overline{K_2{}^{\ell_2}} :: e_2{}^{\texttt{app}}\rangle$ By $\Gamma \vdash \texttt{read-only}\ T$ we have that $\tau = \texttt{call}\ z_f$ where $z_f \in \texttt{cod}(\mathbb{C}_{\texttt{ro}}(\Gamma \vDash L))$. We may therefore apply the assumption that L is classically constant time and Lemma 3 to get that there exist S_2', $\overline{K_2{}^{\ell_2}}$, e_2, and $\overline{\delta'}$ such that $\langle S_1' \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\texttt{lib}}\rangle \xRightarrow{\overline{\delta'^{\texttt{lib}}} \diamond \tau_2{}^{\texttt{lib}\to\texttt{app}}}{}^* \langle S_2' \mid \overline{K_2{}^{\ell_2}} :: e_2{}^{\texttt{app}}\rangle$ where $\texttt{ct}(\overline{\delta'^{\texttt{lib}}}) = \texttt{ct}(\overline{\delta'^{\texttt{lib}}})$. $\texttt{inv}(S_2, S_2')$ follows by the fact that new events must be the same due to the constant timeness and Lemma 4. ■

Case $(\langle S_2 \mid \overline{K_2{}^{\ell_2}} :: e_2\rangle, \langle S_2' \mid \overline{K_2{}^{\ell_2}} :: e_2\rangle) \in \mathbb{T}[\![T']\!]$**:** By the prior states being in $\mathbb{L}$ we have that $\texttt{inv}(S_2, S_2')$. By inversion on $\Gamma \vdash \texttt{read-only}\ T$ we have that $\Gamma \vdash \texttt{read-only}\ T'$. By Lemma 1 $\mathbb{C}_{\texttt{ro}}(\Gamma \vDash L) \vDash S$ and $\mathbb{C}_{\texttt{ro}}(\Gamma \vDash L) \vDash S$. We then apply our inductive hypothesis for T' and our proof is complete. ■

□

Lemma 7. *For a classically constant time library* $\Gamma \vDash L$*, secret context* Δ*, read-only application* $\Gamma, \Delta \vDash (H, e)$*, and initial states* $\mathbb{C}_{\texttt{ro}}(\Gamma \vDash L) \mid \Delta \mid H \vDash S_0 = S_0'$*, for all* $T \in \texttt{traces}(\langle S_0[\Delta][\mathbb{C}_{\texttt{ro}}(\Gamma \vDash L)] \mid e[\Delta][\mathbb{C}_{\texttt{ro}}(\Gamma \vDash L)]\rangle)$*,*

$$(\langle S_0[\Delta][\mathbb{C}_{\texttt{ro}}(\Gamma \vDash L)] \mid e[\Delta][\mathbb{C}_{\texttt{ro}}(\Gamma \vDash L)]\rangle, \langle S_0'[\Delta][\mathbb{C}_{\texttt{ro}}(\Gamma \vDash L)] \mid e[\Delta][\mathbb{C}_{\texttt{ro}}(\Gamma \vDash L)]\rangle) \in \mathbb{T}[\![T]\!].$$

Proof. We first show that $\texttt{inv}(S_0[\Delta][\mathbb{C}_{\texttt{ro}}(\Gamma \vDash L)], S_0'[\Delta][\mathbb{C}_{\texttt{ro}}(\Gamma \vDash L)])$. Condition 3 holds by inversion on $L \mid \Delta \mid H \vDash S_0 = S_0'$. Condition 2 holds by inversion on $\Delta \mid H \vDash_{\texttt{protected}} S_0 = S_0'$. Condition 1 holds by inversion on $L \mid \Delta \mid H \vDash S_0 = S_0'$. By assumption and Lemma 2 we have that $\Gamma \vdash \texttt{read-only}\ T$. Our goal then follows by Lemma 6. □

Lemma 8 (FTLR). *If* $(\langle S \mid \overline{K^\ell} :: e\rangle, \langle S' \mid \overline{K^\ell} :: e\rangle) \in \mathbb{T}[\![T]\!]$*, then* $T \in \texttt{traces}(\langle S \mid \overline{K^\ell} :: e^{\texttt{app}}\rangle)$ *and there exists a trace* $T' \in \texttt{traces}(\langle S' \mid \overline{K^\ell} :: e^{\texttt{app}}\rangle)$ *such that* $\texttt{ct}(T) = \texttt{ct}(T')$.

Proof. By induction on T. □

Theorem 2 ($\mathbb{C}_{\texttt{ro}}$ guarantees read-only robust constant time). *If* $\Gamma \vDash L$ *is classically constant time and does not contain any* $\texttt{protect}_p$ *subterms, then* $\mathbb{C}_{\texttt{ro}}(\Gamma \vDash L)$ *is robustly constant time for read-only attackers (that do not contain* $\texttt{protect}_p$*).*

Proof. By Lemma 6 and Lemma 8. □

D.2 Speculative protections

For this section we consider a fixed speculator.

Definition 14. *We say a step* $\langle \Phi_1 \mid e_1\rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow} \langle \Phi_2 \mid e_2\rangle$ is speculating *if* $\Phi_1.\Xi \neq \bullet$ *and* $\Phi_2.\Xi \neq \bullet$*. We say a subtrace* $\langle \Phi_1 \mid e_1\rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}{}^* \langle \Phi_2 \mid e_2\rangle$ *is* speculating *if all steps from* $\langle \Phi_1 \mid e_1\rangle$ *to* $\langle \Phi_2 \mid e_2\rangle$ *along* $\overline{\delta}$ *are speculating.*

σ	$::=$	$(\delta, \mathcal{N})$	non-speculative trace
	$\mid$	$(\overline{\delta}, \mathcal{S})$	speculatively rolled-back trace
	$\mid$	$[\overline{\sigma}, \overline{\delta}]$	speculatively committed trace with $\overline{\delta}$ speculated on

Σ	$::=$	$(\epsilon, \mathcal{N})$
	$\mid$	$(\overline{\delta}, \mathcal{S})$

$$\mid \quad [\overline{\Sigma}, \overline{\epsilon}]$$

$$\boxed{\text{crunch} : \overline{\sigma} \rightarrow \overline{\delta}}$$
$$\boxed{\text{crunch} : \overline{\Sigma} \rightarrow \overline{\delta}}$$

$$\begin{aligned}
\text{crunch}\,(\bullet) &\triangleq \bullet \\
\text{crunch}\,((\delta,\ \mathcal{N}) :: \overline{\sigma}) &\triangleq \delta \diamond \text{crunch}(\overline{\sigma}) \\
\text{crunch}\left((\overline{\delta},\ \mathcal{S}) :: \overline{\sigma}\right) &\triangleq 0 :: \overline{\delta} \diamond 0 :: \text{crunch}(\overline{\sigma}) \\
\text{crunch}\left([\overline{\sigma}, \overline{\delta}] :: \overline{\sigma'}\right) &\triangleq 0 :: \text{crunch}(\overline{\sigma}) \diamond \overline{\delta} \diamond \text{crunch}(\overline{\sigma'}) \\
\text{crunch}\left(\overline{\Sigma}\right) &\triangleq \text{crunch}\left(\text{unlabel}(\overline{\Sigma})\right)
\end{aligned}$$

$$\boxed{\text{nonspec} : \overline{\Sigma} \rightarrow \overline{\epsilon}}$$

$$\begin{aligned}
\text{nonspec}\,(\bullet) &\triangleq \bullet \\
\text{nonspec}\left((\epsilon,\ \mathcal{N}) :: \overline{\Sigma}\right) &\triangleq \epsilon :: \text{nonspec}\left(\overline{\Sigma}\right) \\
\text{nonspec}\left((\overline{\delta},\ \mathcal{S}) :: \overline{\Sigma}\right) &\triangleq \text{nonspec}\left(\overline{\Sigma}\right) \\
\text{nonspec}\left([\overline{\Sigma}, \overline{\epsilon}] :: \overline{\Sigma'}\right) &\triangleq \overline{\epsilon} \diamond \text{nonspec}(\overline{\Sigma}) \diamond \text{nonspec}(\overline{\Sigma'})
\end{aligned}$$

$$\boxed{\text{specValid}\ \overline{\Xi}}$$

$$\frac{}{\text{specValid}\ \bullet} \qquad \frac{\text{specValid}\ (\text{addEvents}(\overline{\Xi}, \overline{\delta} \diamond \overline{\mu}))}{\text{specValid}\ (S,\ e,\ \overline{\delta},\ \overline{\mu}) :: \overline{\Xi}}$$

$$\boxed{(\langle \Phi \mid e \rangle \xhookrightarrow{\overline{\delta}}\!\!\twoheadrightarrow^{+} \langle \Phi \mid e \rangle)|_{\mathcal{S}} = \sigma}$$

$$\frac{\text{length}(\Phi_1.\Xi) = \text{length}(\Phi_2.\Xi)}{(\langle \Phi_1 \mid e_1 \rangle \xhookrightarrow{\delta}\!\!\twoheadrightarrow \langle \Phi_2 \mid e_2 \rangle)|_{\mathcal{S}} = (\delta,\ \mathcal{N})}$$

$$\frac{\begin{array}{c}\text{length}(\Phi_1.\Xi) + 1 = \text{length}(\Phi_2.\Xi) = \text{length}(\Phi_3.\Xi) = \text{length}(\Phi_4.\Xi) + 1 \\ \Phi_3.\Xi = (\Phi_1.S,\ e_1,\ \overline{\delta_3},\ \overline{\mu}) :: \overline{\Xi} \qquad (\langle \Phi_2 \mid e_2 \rangle \xhookrightarrow{\overline{\delta_2}}\!\!\twoheadrightarrow^{*} \langle \Phi_3 \mid e_3 \rangle)|_{\mathcal{S}^*} = \overline{\sigma}\end{array}}{(\langle \Phi_1 \mid e_1 \rangle \xhookrightarrow{0}\!\!\twoheadrightarrow \langle \Phi_2 \mid e_2 \rangle \xhookrightarrow{\overline{\delta_2}}\!\!\twoheadrightarrow^{*} \langle \Phi_3 \mid e_3 \rangle \xhookrightarrow{\overline{\delta_3}}\!\!\twoheadrightarrow \langle \Phi_4 \mid e_4 \rangle)|_{\mathcal{S}} = [\overline{\sigma}, \overline{\delta_3}]}$$

$$\frac{\begin{array}{c}\text{length}(\Phi_1.\Xi) + 1 = \text{length}(\Phi_2.\Xi) = \text{length}(\Phi_3.\Xi) = \text{length}(\Phi_4.\Xi) + 1 \\ \Phi_3.\Xi = (\widehat{\Phi_1.S,\ e_1},\ \overline{\delta}) :: \overline{\Xi} \qquad (\langle \Phi_2 \mid e_2 \rangle \xhookrightarrow{\overline{\delta_2}}\!\!\twoheadrightarrow^{*} \langle \Phi_3 \mid e_3 \rangle)|_{\mathcal{S}^*} = \overline{\sigma}\end{array}}{(\langle \Phi_1 \mid e_1 \rangle \xhookrightarrow{0}\!\!\twoheadrightarrow \langle \Phi_2 \mid e_2 \rangle \xhookrightarrow{\overline{\delta_2}}\!\!\twoheadrightarrow^{*} \langle \Phi_3 \mid e_3 \rangle \xhookrightarrow{0}\!\!\twoheadrightarrow \langle \Phi_4 \mid e_1 \rangle)|_{\mathcal{S}} = (\text{crunch}(\overline{\sigma}),\ \mathcal{S})}$$

$$\boxed{(\langle \Phi \mid e \rangle \xhookrightarrow{\overline{\delta}}\!\!\!\!\twoheadrightarrow^{*} \langle \Phi \mid e \rangle)|_{\mathcal{S}^*} = \overline{\sigma}}$$

$$\frac{}{(\langle \Phi_1 \mid e_1 \rangle \xhookrightarrow{\bullet}\!\!\!\!\twoheadrightarrow^{0} \langle \Phi_1 \mid e_1 \rangle)|_{\mathcal{S}^*} = \bullet}$$

$$\frac{(\langle \Phi_1 \mid e_1 \rangle \xhookrightarrow{\overline{\delta_1}}\!\!\!\!\twoheadrightarrow^{+} \langle \Phi_2 \mid e_2 \rangle)|_{\mathcal{S}} = \sigma_1 \qquad (\langle \Phi_2 \mid e_2 \rangle \xhookrightarrow{\overline{\delta_2}}\!\!\!\!\twoheadrightarrow^{*} \langle \Phi_3 \mid e_3 \rangle)|_{\mathcal{S}^*} = \overline{\sigma_2}}{(\langle \Phi_1 \mid e_1 \rangle \xhookrightarrow{\overline{\delta_1}}\!\!\!\!\twoheadrightarrow^{+} \langle \Phi_2 \mid e_2 \rangle \xhookrightarrow{\overline{\delta_2}}\!\!\!\!\twoheadrightarrow^{*} \langle \Phi_3 \mid e_3 \rangle)|_{\mathcal{S}^*} = \sigma_1 :: \overline{\sigma_2}}$$

$$\boxed{\text{last-nonspec} : \overline{\Sigma} \rightharpoonup \epsilon}$$

$$\begin{aligned}
\left((\epsilon, \mathcal{N}) :: \overline{\Sigma}\right) &\triangleq \begin{cases} \epsilon' & \text{when last-nonspec}\left(\overline{\Sigma}\right) = \epsilon' \\ \epsilon & \text{otherwise} \end{cases} \\
\text{last-nonspec}\left((\overline{\delta}, \mathcal{S}) :: \overline{\Sigma}\right) &\triangleq \text{last-nonspec}\left(\overline{\Sigma}\right) \\
\text{last-nonspec}\left([\overline{\Sigma}, \overline{\epsilon}] :: \overline{\Sigma'}\right) &\triangleq \begin{cases} \epsilon' & \text{when last-nonspec}\left(\overline{\Sigma'}\right) = \epsilon' \\ \epsilon' & \text{when last-nonspec}\left(\overline{\Sigma}\right) = \epsilon' \\ \epsilon & \text{when last}(\overline{\epsilon}) = \epsilon \end{cases}
\end{aligned}$$

$$\boxed{\text{label} : \epsilon \longrightarrow \ell}$$
$$\boxed{\text{label} : \overline{\Sigma} \longrightarrow \ell}$$
$$\boxed{\text{label}_{\ell} : \Sigma \longrightarrow \ell}$$

$$\begin{aligned}
\text{label}\left(\delta^{\ell}\right) &\triangleq \ell \\
\text{label}\left(\tau^{\ell \to \ell'}\right) &\triangleq \ell'
\end{aligned}$$

$$\begin{aligned}
\text{label}\,(\bullet) &\triangleq \text{app} \\
\text{label}\left(\overline{\Sigma}\right) &\triangleq \text{label}\left(\text{last-nonspec}\left(\overline{\Sigma}\right)\right) \\
\text{label}_{\ell}\,(\Sigma) &\triangleq \begin{cases} \text{label}(\epsilon) & \text{when last-nonspec}\,(\Sigma :: \bullet) = \epsilon \\ \ell & \text{otherwise} \end{cases}
\end{aligned}$$

$$\boxed{\langle \Phi \mid \overline{K^{\ell}} :: e^{\ell} \rangle \xrightarrow{\Sigma} \langle \Phi \mid \overline{K^{\ell}} :: e^{\ell} \rangle}$$

$$\frac{\begin{array}{c} \langle \Phi_1.S \mid \overline{{K_1}^{\ell_{K1}}} :: {e_1}^{\ell_1} \rangle \xRightarrow{\text{nonspec}(\Sigma)}{}^{*} \langle \Phi_2.S \mid \overline{{K_2}^{\ell_{K2}}} :: {e_2}^{\text{label}_{\ell_1}(\Sigma)} \rangle \\ \langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \xhookrightarrow{\text{crunch}(\Sigma)}\!\!\!\!\twoheadrightarrow^{*} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle \\ (\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \xhookrightarrow{\text{crunch}(\Sigma)}\!\!\!\!\twoheadrightarrow^{*} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle)|_{\mathcal{S}} = \text{unlabel}(\Sigma) \end{array}}{\langle \Phi_1 \mid \overline{{K_1}^{\ell_{K1}}} :: {e_1}^{\ell_1} \rangle \xrightarrow{\Sigma} \langle \Phi_2 \mid \overline{{K_2}^{\ell_{K2}}} :: e^{\text{label}_{\ell_1}(\Sigma)} \rangle}$$

Lemma 9. *If* $\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \xhookrightarrow{\overline{\delta}}\!\!\!\!\twoheadrightarrow^{*} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle$ *and* $(\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \xhookrightarrow{\overline{\delta}}\!\!\!\!\twoheadrightarrow^{*} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle)|_{\mathcal{S}} = \sigma$, *then* $\Phi_2.S = \text{commit}(\Phi_1.S, \text{nonspec}(\sigma))$.

Proof. By simultaneous induction on $(\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}{}^{*} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle)|_{\mathcal{S}} = \sigma$ and $\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}{}^{*} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle$ □

Lemma 10. *If* $(\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}{}^{*} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle)|_{\mathcal{S}} = \sigma$, *then* $\overline{\delta} = \mathsf{crunch}(\overline{\sigma})$.

Proof. By induction on $(\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}{}^{*} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle)|_{\mathcal{S}} = \sigma$. □

Lemma 11. *If* $(\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}{}^{*} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle)|_{\mathcal{S}^*} = \overline{\sigma_1}$ *and*

$$(\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}{}^{*} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle)|_{\mathcal{S}^*} = \overline{\sigma_2},$$

then $\overline{\sigma_1} = \overline{\sigma_2}$.

Proof. By induction on the derivation of σ_1. □

Lemma 12. *If* $\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}{}^{*} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle$ *and* $\Phi_1.\Xi = \Phi_2.\Xi = \bullet$, *then there exists a unique* $\overline{\sigma}$ *such that* $(\langle \Phi_1 \mid e_1 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}{}^{*} \langle \Phi_2 \mid e_2 \rangle)|_{\mathcal{S}^*} = \overline{\sigma}$.

Proof. Existence follows by induction on the operational semantics with the state of specValid $\overline{\Xi}$ as an invariant. Uniqueness follows by Lemma 11. □

Lemma 13. *If* $\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}{}^{*} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle$ *and* $\Phi_1.\Xi = \Phi_2.\Xi = \bullet$, *then for all* ℓ_{K1} *and* ℓ_1 *there exist unique* $\overline{\Sigma}$, *and* ℓ_{K2} *such that* $\langle \Phi_1 \mid \overline{{K_1}^{\ell_{K1}}} :: {e_1}^{\ell_1} \rangle \overset{\overline{\Sigma}}{\rightarrow}{}^{*} \langle \Phi_2 \mid \overline{{K_2}^{\ell_{K2}}} :: {e_2}^{\mathsf{label}(\overline{\Sigma})} \rangle$ *and* $\mathsf{crunch}(\overline{\Sigma}) = \overline{\delta}$.

Proof. By Lemma 12 there exists a unique $\overline{\sigma}$ such that $(\langle \Phi_1 \mid e_1 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}{}^{*} \langle \Phi_2 \mid e_2 \rangle)|_{\mathcal{S}^*} = \overline{\sigma}$. We construct the labels and non-speculative reduction by induction on $\overline{\sigma}$. The remaining conditions follow by assumption and Lemma 10 □

Lemma 14 (Speculation can't change protection). *If* $\langle \Phi_0 \mid e_0 \rangle \overset{0}{\hookrightarrow\!\!\!\rightarrow} \langle \Phi_1 \mid e_1 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}{}^{*} \langle \Phi_2 \mid e_2 \rangle$ *such that* $\Phi_0.\Xi = \bullet$, $\Phi_0.S.p = \mathtt{public}$, $\Phi_1.\Xi \neq \bullet$, *and the subtrace* $\langle \Phi_1 \mid e_1 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\rightarrow}{}^{*} \langle \Phi_2 \mid e_2 \rangle$ *is speculating, then* $\mathsf{read}_b\ v \hookleftarrow z_b[z_o] \in \overline{\delta} \Rightarrow \Phi_0.S(z_b) \neq \mathtt{protected}$ *and* $\mathsf{write}_b\ v \mapsto z_b[z_o] \in \overline{\delta} \Rightarrow \Phi_0.S(z_b) \neq \mathtt{protected}$.

Proof. By induction on the operational semantics with the invariant that that if $\Phi_2.S.p = \mathtt{protected}$, then $\mathsf{protect}_p$ exists in $\Phi_2.\Xi$. □

Definition 15. *We define our speculative state invariant,* $\mathsf{sinv}(\Phi, \Phi')$ *as follows:*

(1) $\mathsf{inv}(\Phi.S, \Phi'.S)$
(2) $\Phi.a = \Phi'.a$
(3) $\Phi.\Xi = \Phi'.\Xi$

$$\mathbb{E}_{\mathsf{app}}[\![(\epsilon, \mathcal{N})]\!] \triangleq \left\{ \begin{array}{l} (\Phi, \overline{K^{\ell}}, e, \Phi', \overline{K^{\ell}}, e, \\ \quad \Phi_1, \overline{{K_1}^{\ell_1}}, \Phi_1', e_1, \overline{{K_1}^{\ell_1}}, e_1) \end{array} \middle| \begin{array}{l} \langle \Phi \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \xrightarrow{(\epsilon, \mathcal{N})} \langle \Phi_1 \mid \overline{{K_1}^{\ell_1}} :: {e_1}^{\mathsf{app}} \rangle \\ \langle \Phi' \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \xrightarrow{(\epsilon, \mathcal{N})} \langle \Phi_1' \mid \overline{{K_1}^{\ell_1}} :: {e_1}^{\mathsf{app}} \rangle \\ \mathsf{sinv}(\Phi_1, \Phi_1') \end{array} \right\}$$

$$\mathbb{E}_{\texttt{lib}}[\![(\delta^{\texttt{lib}},\ \mathcal{N})]\!] \triangleq \left\{ \begin{array}{l} (\Phi, \overline{K^{\ell}}, e, \Phi', \overline{K'^{\ell'}}, e', \\ \ \Phi_1, \overline{K_1{}^{\ell_1}}, e_1, \Phi'_1, \overline{K'_1{}^{\ell'_1}}, e'_1) \end{array} \middle| \begin{array}{l} \exists \delta'. \\ \quad \langle \Phi \mid \overline{K^{\ell}} :: e^{\texttt{lib}} \rangle \xrightarrow{(\delta^{\texttt{lib}}, \mathcal{N})} \langle \Phi_1 \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\texttt{lib}} \rangle \\ \quad \langle \Phi' \mid \overline{K'^{\ell'}} :: e'^{\texttt{lib}} \rangle \xrightarrow{(\delta'^{\texttt{lib}}, \mathcal{N})} \langle \Phi'_1 \mid \overline{K'_1{}^{\ell'_1}} :: e'_1{}^{\texttt{lib}} \rangle \\ \quad \mathsf{ct}(\delta) = \mathsf{ct}(\delta') \end{array} \right\}$$

$$\mathbb{E}_{\texttt{lib}}[\![(\tau^{\texttt{lib}\to\mathsf{app}},\ \mathcal{N})]\!] \triangleq \left\{ \begin{array}{l} (\Phi, \overline{K^{\ell}}, e, \Phi', \overline{K^{\ell}}, e, \\ \ \Phi_1, \overline{K_1{}^{\ell_1}}, e_1, \Phi'_1, \overline{K_1{}^{\ell_1}}, e_1) \end{array} \middle| \begin{array}{l} \langle \Phi \mid \overline{K^{\ell}} :: e^{\texttt{lib}} \rangle \xrightarrow{(\tau^{\texttt{lib}\to\mathsf{app}}, \mathcal{N})} \langle \Phi_1 \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\mathsf{app}} \rangle \\ \langle \Phi' \mid \overline{K^{\ell}} :: e^{\texttt{lib}} \rangle \xrightarrow{(\tau^{\texttt{lib}\to\mathsf{app}}, \mathcal{N})} \langle \Phi'_1 \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\mathsf{app}} \rangle \\ \mathsf{sinv}(\Phi_1, \Phi'_1) \end{array} \right\}$$

$$\mathbb{E}_{\mathsf{app}}[\![(\overline{\delta},\ \mathcal{S})]\!] \triangleq \left\{ \begin{array}{l} (\Phi, \overline{K^{\ell}}, e, \Phi', \overline{K^{\ell}}, e, \\ \ \Phi_1, \overline{K^{\ell}}, e, \Phi'_1, \overline{K^{\ell}}, e) \end{array} \middle| \begin{array}{l} \langle \Phi \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \xrightarrow{(\overline{\delta}, \mathcal{S})} \langle \Phi_1 \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \\ \langle \Phi' \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \xrightarrow{(\overline{\delta}, \mathcal{S})} \langle \Phi'_1 \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \\ \mathsf{sinv}(\Phi_1, \Phi'_1) \end{array} \right\}$$

$$\mathbb{E}_{\texttt{lib}}[\![(\overline{\delta},\ \mathcal{S})]\!] \triangleq \left\{ \begin{array}{l} (\Phi, \overline{K^{\ell}}, e, \Phi', \overline{K'^{\ell'}}, e', \\ \ \Phi_1, \overline{K^{\ell}}, e, \Phi'_1, \overline{K'^{\ell'}}, e') \end{array} \middle| \begin{array}{l} \exists \overline{\delta'}. \\ \quad \langle \Phi \mid \overline{K^{\ell}} :: e^{\texttt{lib}} \rangle \xrightarrow{(\overline{\delta}, \mathcal{S})} \langle \Phi_1 \mid \overline{K^{\ell}} :: e^{\texttt{lib}} \rangle \\ \quad \langle \Phi' \mid \overline{K'^{\ell'}} :: e'^{\texttt{lib}} \rangle \xrightarrow{(\overline{\delta'}, \mathcal{S})} \langle \Phi'_1 \mid \overline{K'^{\ell'}} :: e'^{\texttt{lib}} \rangle \\ \quad \mathsf{ct}(\overline{\delta}) = \mathsf{ct}(\overline{\delta'}) \end{array} \right\}$$

$$\mathbb{E}_{\mathsf{app}}[\![[\overline{\Sigma}, \overline{\epsilon}]]\!] \triangleq \left\{ \begin{array}{l} (\Phi, \overline{K^{\ell}}, e, \Phi', \overline{K^{\ell}}, e, \\ \ \Phi_1, \overline{K_1{}^{\ell_1}}, e_1, \Phi'_1, \overline{K_1{}^{\ell_1}}, e_1) \end{array} \middle| \begin{array}{l} \exists \overline{\epsilon'}. \\ \quad \langle \Phi \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \xrightarrow{[\overline{\Sigma}, \overline{\epsilon}]} \langle \Phi_1 \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\mathsf{label}([\overline{\Sigma}, \overline{\epsilon}])} \rangle \\ \quad \langle \Phi' \mid \overline{K^{\ell}} :: e^{\mathsf{app}} \rangle \xrightarrow{[\overline{\Sigma}, \overline{\epsilon'}]} \langle \Phi'_1 \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\mathsf{label}([\overline{\Sigma}, \overline{\epsilon}])} \rangle \\ \quad \mathsf{label}_{\mathsf{app}}([\overline{\Sigma}, \overline{\epsilon}]) = \mathsf{app} \Rightarrow \mathsf{sinv}(\Phi_1, \Phi'_1) \\ \quad \mathsf{ct}(\overline{\epsilon}) = \mathsf{ct}(\overline{\epsilon'}) \end{array} \right\}$$

$$\mathbb{E}_{\texttt{lib}}[\![[\overline{\Sigma},\overline{\epsilon}]]\!] \triangleq \left\{ \begin{array}{l} (\Phi,\overline{K^{\ell}},e,\Phi',\overline{K'^{\ell'}},e', \\ \quad \Phi_1,\overline{K_1{}^{\ell_1}},e_1,\Phi_1',\overline{K_1'^{\ell_1'}},e_1') \end{array} \middle| \begin{array}{l} \exists\overline{\Sigma'},\overline{\epsilon'}. \\ \quad \langle\Phi \mid \overline{K^{\ell}} :: e^{\texttt{lib}}\rangle \xrightarrow{[\overline{\Sigma},\overline{\epsilon}]} \langle\Phi_1 \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\texttt{lib}}\rangle \\ \quad \langle\Phi' \mid \overline{K'^{\ell'}} :: e'^{\texttt{lib}}\rangle \xrightarrow{[\overline{\Sigma'},\overline{\epsilon'}]} \langle\Phi_1' \mid \overline{K_1'^{\ell_1'}} :: e_1'^{\texttt{lib}}\rangle \\ \quad \mathsf{ct}(\mathsf{crunch}([\overline{\Sigma},\overline{\epsilon}])) = \mathsf{ct}(\mathsf{crunch}([\overline{\Sigma'},\overline{\epsilon'}])) \end{array} \right\}$$

$$\mathbb{S}_{\ell}[\![\Sigma :: \overline{\Sigma'}]\!] \triangleq \left\{ (\langle\Phi \mid \overline{K^{\ell_K}} :: e^{\ell}\rangle,\ \langle\Phi' \mid \overline{K'^{\ell_K'}} :: e'^{\ell}\rangle) \middle| \begin{array}{l} \exists\Phi_1,\overline{K_1{}^{\ell_{K1}}},e_1,\Phi_1',\overline{K_1'^{\ell_{K1}'}},e_1',\Phi_2,\overline{K_2{}^{\ell_{K2}}},e_2,\Phi_2',\overline{K_2'^{\ell_{K2}'}},e_2'. \\ \quad (\Phi,\overline{K^{\ell_K}},e,\Phi',\overline{K'^{\ell_K'}},e', \\ \quad\quad \Phi_1,\overline{K_1{}^{\ell_{K1}}},e_1,\Phi_1',\overline{K_1'^{\ell_{K1}'}},e_1') \in \mathbb{E}_{\ell}[\![\Sigma]\!] \\ \quad (\Phi_1,\ \Phi_1',\ \overline{K_1{}^{\ell_{K1}}},\ e_1,\ \Phi_2,\ \Phi_2',\ \overline{K_2{}^{\ell_{K2}}},\ e_2) \in \mathbb{S}_{\mathsf{label}(\Sigma)}[\![\overline{\Sigma'}]\!] \end{array} \right\}$$

$$\mathbb{S}_{\mathsf{app}}[\![\bullet]\!] \triangleq \left\{ (\langle\Phi \mid \overline{K^{\ell_K}} :: e^{\mathsf{app}}\rangle,\ \langle\Phi' \mid \overline{K^{\ell_K}} :: e^{\mathsf{app}}\rangle) \,\middle|\, \mathsf{sinv}(\Phi,\Phi') \right\}$$
$$\mathbb{S}_{\texttt{lib}}[\![\bullet]\!] \triangleq \left\{ (\langle\Phi \mid \overline{K^{\ell_K}} :: e^{\texttt{lib}}\rangle,\ \langle\Phi' \mid \overline{K'^{\ell_K'}} :: e'^{\texttt{lib}}\rangle) \right\}$$

Lemma 15 ($\mathbb{C}_{\mathsf{spec}}$ preserves classic speculative constant time). *If $\Gamma \vDash L$ is classically speculative constant time then $\mathbb{C}_{\mathsf{spec}}\Gamma \vDash L$ is classically speculative constant time.*

Proof. By induction on the compiler. □

Lemma 16 ($\mathbb{C}_{\mathsf{spec}}$ prevents speculative returns). *If $\Gamma \vDash L$ is classically speculative constant time with respect to a speculation oracle* $\mathsf{spec} : A \times S \times e \longrightarrow A \times d$, *then for all secret contexts Δ, classical "applications" $\Gamma, \Delta \vdash (H, e_{\Gamma})$, initial states S_0, S_0' such that $L \mid \Delta \mid H \vDash S_0$, microarchitectural states $a : A$, $\Phi_0 = \{S = S_0[\Delta][L], a = a, \Xi = \bullet\}$, and $e_0 = e_{\Gamma}[\Delta][\mathbb{C}_{\mathsf{spec}}(\Gamma \vDash L)]$, then* $\mathsf{begin} \diamond \overline{\delta} \diamond \mathsf{end}\ v \in \mathsf{specTraces}(\langle\Phi_0 \mid e_0\rangle)$ *implies that there exists a $\overline{\Sigma}$ such that* $\mathsf{nonspec}(\overline{\Sigma}) = \mathsf{begin} \diamond \overline{\delta} \diamond \mathsf{end}\ v$ *and the tail of $\overline{\Sigma}$ is of the form* $(\mathsf{fence}^{\texttt{lib}}, \mathcal{N}) \diamond (\overline{\delta_1}, \mathcal{S})^* \diamond (\mathsf{ret}\ v^{\texttt{lib}\to\mathsf{app}}, \mathcal{N}) \diamond (\mathsf{end}\ v^{\mathsf{app}\to\texttt{lib}}, \mathcal{N})$ *or* $(\mathsf{fence}^{\texttt{lib}}, \mathcal{N}) \diamond [(\overline{\delta_1}, \mathcal{S})^* \diamond (\mathsf{ret}\ v^{\texttt{lib}\to\mathsf{app}}, \mathcal{N}), 0] \diamond (\mathsf{end}\ v^{\mathsf{app}\to\texttt{lib}}, \mathcal{N})$.

Proof. By induction on the compiler. □

Lemma 17 ($\mathbb{S}$ extension). *If*

$$\langle\Phi \mid \bullet :: e^{\mathsf{app}}\rangle \xrightarrow{\overline{\Sigma_1}}{}^{*} \langle\Phi_1 \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\mathsf{label}(\overline{\Sigma_1})}\rangle \xrightarrow{\Sigma_2} \langle\Phi_2 \mid \overline{K_2{}^{\ell_2}} :: e_2{}^{\mathsf{label}_{\mathsf{label}(\overline{\Sigma_1})}(\Sigma_2)}\rangle$$

and $(\langle\Phi \mid \bullet :: e^{\mathsf{app}}\rangle, \langle\Phi' \mid \bullet :: e^{\mathsf{app}}\rangle) \in \mathbb{S}_{\mathsf{app}}[\![\overline{\Sigma_1}]\!]$, *then if*

(1) *There exist $\overline{\Sigma_1'}, \Phi_1', \overline{K_1'^{\ell_1'}}$, and e_1' such that* $\langle\Phi' \mid \bullet :: e^{\mathsf{app}}\rangle \xrightarrow{\overline{\Sigma_1'}}{}^{*} \langle\Phi_1' \mid \overline{K_1'^{\ell_1'}} :: e_1'^{\mathsf{label}(\overline{\Sigma_1})}\rangle$.
(2) *If* $\mathsf{label}(\overline{\Sigma_1}) = \mathsf{app}$, *then* $\overline{K_1{}^{\ell_1}} :: e_1 = \overline{K_1'^{\ell_1'}} :: e_1'$ *and* $\mathsf{sinv}(\Phi_1, \Phi_1')$.
(3) *Let* $\langle\Phi.S \mid \bullet :: e^{\mathsf{app}}\rangle \xRightarrow{\overline{\delta_0} \diamond (\mathsf{call}\ z_f)^{\mathsf{app}\to\texttt{lib}} \diamond \overline{\delta_1}}{}^{*} \langle\Phi_1.S \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\mathsf{label}(\overline{\Sigma_1})}\rangle$ *such that* $(\mathsf{call}\ z_f')^{\mathsf{app}\to\texttt{lib}} \notin \overline{\delta_1}$ *and* $\overline{\delta_0} \diamond (\mathsf{call}\ z_f)^{\mathsf{app}\to\texttt{lib}} \diamond \overline{\delta_1} = \mathsf{nonspec}(\overline{\Sigma_1})$. *Then* $\langle\Phi'.S \mid \bullet :: e^{\mathsf{app}}\rangle \xRightarrow{\overline{\delta_0'} \diamond (\mathsf{call}\ z_f)^{\mathsf{app}\to\texttt{lib}} \diamond \overline{\delta_1'}}{}^{*}$

$\langle \Phi_1'.S \mid \overline{K_1'^{\ell_1'}} :: e_1'^{\mathsf{label}(\overline{\Sigma_1})} \rangle$ *such that* $(\mathsf{call}\ z_f')^{\mathsf{app}\to\mathtt{lib}} \notin \overline{\delta_1'}$ *and* $\overline{\delta_0'} \diamond (\mathsf{call}\ z_f)^{\mathsf{app}\to\mathtt{lib}} \diamond \overline{\delta_1'} = \mathsf{nonspec}(\overline{\Sigma_1'})$.

(4) $\mathsf{ct}(\mathsf{crunch}(\overline{\Sigma_1})) = \mathsf{ct}(\mathsf{crunch}(\overline{\Sigma_1'}))$

imply that there exist Φ_2', $\overline{K_2'^{\ell_2'}}$, *and* e_2' *such that* $(\Phi_1, \overline{K_1^{\ell_1}}, e_1, \Phi_1', \overline{K_1'^{\ell_1'}}, e_1', \Phi_2, \overline{K_2^{\ell_2}}, e_2, \Phi_2', \overline{K_2'^{\ell_2'}}, e_2') \in \mathbb{E}_{\mathsf{label}(\overline{\Sigma_1})}[\![\Sigma_2]\!]$, *then* $(\langle \Phi \mid \bullet :: e^{\mathsf{app}} \rangle, \langle \Phi' \mid \bullet :: e^{\mathsf{app}} \rangle) \in \mathbb{S}_{\mathsf{app}}[\![\overline{\Sigma_1} \diamond \Sigma_2]\!]$.

Proof. By induction on $\overline{\Sigma_1}$. □

Lemma 18. *If* $\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \overset{0}{\hookrightarrow\!\!\!\!\rightarrow} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\!\rightarrow}{}^{*} \langle \Phi_3 \mid \overline{K_3} :: e_3 \rangle$, $\Phi_1.\Xi = \bullet$, $\Phi_2.\Xi \neq \bullet$, $\langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\!\rightarrow}{}^{*} \langle \Phi_3 \mid \overline{K_3} :: e_3 \rangle$ *is speculating, and* $\mathsf{sinv}(\Phi_1, \Phi_1')$, *then* $\langle \Phi_1' \mid \overline{K_1} :: e_1 \rangle \overset{0}{\hookrightarrow\!\!\!\!\rightarrow} \langle \Phi_2' \mid \overline{K_2} :: e_2 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\!\rightarrow}{}^{*} \langle \Phi_3' \mid \overline{K_3} :: e_3 \rangle$ *and* $\mathsf{sinv}(\Phi_3, \Phi_3')$.

Proof. We proceed by induction on $\overline{\delta}$. We have by $\mathsf{sinv}(\Phi_1, \Phi_1')$ that $\langle \Phi_1' \mid \overline{K_1} :: e_1 \rangle \overset{0}{\hookrightarrow\!\!\!\!\rightarrow} \langle \Phi_2' \mid \overline{K_2} :: e_2 \rangle$, completing our base case. In the case where $\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \overset{0}{\hookrightarrow\!\!\!\!\rightarrow} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle \overset{\overline{\delta}}{\hookrightarrow\!\!\!\!\rightarrow}{}^{*} \langle \Phi_3 \mid \overline{K_3} :: e_3 \rangle \overset{\delta_3}{\hookrightarrow\!\!\!\!\rightarrow} \langle \Phi_4 \mid \overline{K_4} :: e_4 \rangle$ we must show that $\mathsf{sinv}(\Phi_3, \Phi_3')$ implies that $\langle \Phi_3' \mid \overline{K_3} :: e_3 \rangle \overset{\delta_3}{\hookrightarrow\!\!\!\!\rightarrow} \langle \Phi_4' \mid \overline{K_4} :: e_4 \rangle$. This follows by case analysis on e_3, the fact that $\mathsf{inv}(\Phi_3.S, \Phi_3'.S)$, and Lemma 14. □

Lemma 19. *For a classically speculatively constant time library* $\Gamma \vDash L$, *secret context* Δ, *memory-safe application* $\Gamma, \Delta \vDash (H, e_a)$, *initial states* $\mathbb{C}_{\mathsf{spec}}(\Gamma \vDash L) \mid \Delta \mid H \vDash S_0 = S_0'$, $S = S_0[\Delta][\mathbb{C}_{\mathsf{spec}}(\Gamma \vDash L)]$, $S' = S_0'[\Delta][\mathbb{C}_{\mathsf{spec}}(\Gamma \vDash L)]$, $\Phi = \{S = S, a = a, \Xi = \bullet\}$; $\Phi' = \{S = S', a = a, \Xi = \bullet\}$, *and* $e = e_a[\Delta][\mathbb{C}_{\mathsf{spec}}(\Gamma \vDash L)]$, *if* $\langle \Phi \mid \bullet :: e^{\mathsf{app}} \rangle \xrightarrow{\overline{\Sigma}}{}^{*} \langle \Phi_1 \mid \overline{K_1^{\ell_1}} :: e_1^{\mathsf{label}(\overline{\Sigma})} \rangle$, *then* $(\langle \Phi \mid \bullet :: e^{\mathsf{app}} \rangle, \langle \Phi' \mid \bullet :: e^{\mathsf{app}} \rangle) \in \mathbb{S}_{\mathsf{app}}[\![\overline{\Sigma}]\!]$.

Proof. We proceed by induction on $\overline{\Sigma}$. In the base case we have $\overline{\Sigma} = \bullet$, where $\mathsf{sinv}(\Phi, \Phi')$ follows from our assumptions.

In our inductive case we have $\overline{\Sigma} = \overline{\Sigma_1} \diamond \Sigma_2$ and therefore $\langle \Phi \mid \bullet :: e^{\mathsf{app}} \rangle \xrightarrow{\overline{\Sigma_1}}{}^{*} \langle \Phi_1 \mid \overline{K_1^{\ell_1}} :: e_1^{\mathsf{label}(\overline{\Sigma_1})} \rangle \xrightarrow{\Sigma_2} \langle \Phi_2 \mid \overline{K_2^{\ell_2}} :: e_2^{\mathsf{label}_{\mathsf{label}(\overline{\Sigma_1})}(\Sigma_2)} \rangle$. By our inductive hypothesis we have that $(\langle \Phi \mid \bullet :: e^{\mathsf{app}} \rangle, \langle \Phi' \mid \bullet :: e^{\mathsf{app}} \rangle) \in \mathbb{S}_{\mathsf{app}}[\![\overline{\Sigma_1}]\!]$. By Lemma 17 we may assume that there exist $\overline{\Sigma_1'}$, Φ_1', $\overline{K_1'^{\ell_1'}}$, and e_1' such that $\langle \Phi' \mid \bullet :: e^{\mathsf{app}} \rangle \xrightarrow{\overline{\Sigma_1'}}{}^{*} \langle \Phi_1' \mid \overline{K_1'^{\ell_1'}} :: e_1'^{\mathsf{label}(\overline{\Sigma_1})} \rangle$ and $\mathsf{label}(\overline{\Sigma_1}) = \mathsf{app} \Rightarrow \overline{K_1^{\ell_1}} :: e_1 = \overline{K_1'^{\ell_1'}} :: e_1' \wedge \mathsf{sinv}(\Phi_1, \Phi_1')$. We must then show that there exist Φ_2', $\overline{K_2'^{\ell_2'}}$, and e_2' such that $(\Phi_1, \overline{K_1^{\ell_1}}, e_1, \Phi_1', \overline{K_1'^{\ell_1'}}, e_1', \Phi_2, \overline{K_2^{\ell_2}}, e_2, \Phi_2', \overline{K_2'^{\ell_2'}}, e_2') \in \mathbb{E}_{\mathsf{label}(\overline{\Sigma_1})}[\![\Sigma_2]\!]$.

We proceed by case analysis on $\mathsf{label}(\overline{\Sigma_1})$ and Σ_2. We first consider all of the cases where $\mathsf{label}(\overline{\Sigma_1}) = \mathsf{app}$. By assumption $\overline{K_1^{\ell_1}} :: e_1 = \overline{K_1'^{\ell_1'}} :: e_1'$ and $\mathsf{sinv}(\Phi_1, \Phi_1')$. We pick $\overline{K_2'^{\ell_2'}} :: e_2' = \overline{K_2^{\ell_2}} :: e_2$. By inversion on $\langle \Phi_1 \mid \overline{K_1^{\ell_1}} :: e_1^{\mathsf{label}(\overline{\Sigma_1})} \rangle \xrightarrow{\Sigma_2} \langle \Phi_2 \mid \overline{K_2^{\ell_2}} :: e_2^{\mathsf{label}_{\mathsf{label}(\overline{\Sigma_1})}(\Sigma_2)} \rangle$ we have

(1) $\langle \Phi_1.S \mid \overline{K_1^{\ell_{K1}}} :: e_1^{\ell_1} \rangle \xRightarrow{\mathsf{nonspec}(\Sigma)}{}^{*} \langle \Phi_2.S \mid \overline{K_2^{\ell_{K2}}} :: e_2^{\mathsf{label}_{\ell_1}(\Sigma)} \rangle$

(2) $\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \overset{\mathsf{crunch}(\Sigma)}{\hookrightarrow\!\!\!\!\rightarrow}{}^{*} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle$

(3) $(\langle \Phi_1 \mid \overline{K_1} :: e_1 \rangle \overset{\mathsf{crunch}(\Sigma)}{\hookrightarrow\!\!\!\!\rightarrow}{}^{*} \langle \Phi_2 \mid \overline{K_2} :: e_2 \rangle)|_{\mathcal{S}} = \mathsf{unlabel}(\Sigma)$

Case $\Sigma_2 = (\epsilon, \mathcal{N})$ **:** We must show that there exists a Φ_2' such that

(1) $\langle \Phi_1 \mid \overline{K_1^{\ell_1}} :: e_1^{\mathsf{app}} \rangle \xrightarrow{(\epsilon, \mathcal{N})} \langle \Phi_2 \mid \overline{K_2^{\ell_2}} :: e_2^{\mathsf{app}} \rangle$

(2) $\langle \Phi_1' \mid \overline{K_1}^{\ell_1} :: e_1{}^{\mathsf{app}}\rangle \xrightarrow{(\epsilon, \mathcal{N})} \langle \Phi_2' \mid \overline{K_2}^{\ell_2} :: e_2{}^{\mathsf{app}}\rangle$
(3) $\mathsf{sinv}(\Phi_2, \Phi_2')$

The first follows immediately by assumption.

For the second we first show that there exists an S_2' such that $\langle \Phi_1'.S \mid \overline{K_1}^{\ell_{K1}} :: e_1{}^{\ell_1}\rangle \overset{\epsilon}{\Rightarrow}{}^* \langle S_2' \mid \overline{K_2}^{\ell_{K2}} :: e_2{}^{\mathsf{label}_{\ell_1}((\epsilon, \mathcal{N}))}\rangle$. This follows by Lemma 6 and the definitions of $\mathbb{T}$ and $\mathbb{A}$. We then let $\Phi_2' = \{S = S_2', a = \mathit{spec}(\Phi_1.a, e_1).2, \Xi = \Phi_2.\Xi\}$. $\langle \Phi_1' \mid \overline{K_1} :: e_1\rangle \xhookrightarrow{\mathsf{unlabel}(\epsilon)}{}^* \langle \Phi_2' \mid \overline{K_2} :: e_2\rangle$ by the same logic as Lemma 5 (by inversion $\mathit{spec}(\Phi_1.a, e_1) = \mathsf{nonspec}$ so the rule SPEC-β applies). $(\langle \Phi_1' \mid \overline{K_1} :: e_1\rangle \xhookrightarrow{\mathsf{unlabel}(\epsilon)}{}^* \langle \Phi_2' \mid \overline{K_2} :: e_2\rangle)|_{\mathcal{S}} = \mathsf{unlabel}(\epsilon)$ by unrolling of definitions and assumption.

The speculative invariant follows by Lemma 6, the definitions of $\mathbb{T}$ and $\mathbb{A}$ and the fact that $\Phi_1.a = \Phi_1'.a$ and $\Phi_1.\Xi = \Phi_1.\Xi$. ■

Case $\Sigma_2 = (\overline{\delta}, \mathcal{S})$ **:** By inversion we have that $\langle \Phi_1 \mid \overline{K_1} :: e_1\rangle \xhookrightarrow{0::\overline{\delta}}{}^* \langle \Phi_3 \mid \overline{K_3} :: e_3\rangle \xhookrightarrow{0} \langle \Phi_2 \mid \overline{K_1} :: e_1\rangle$. By Lemma 18 we have that $\langle \Phi_1' \mid \overline{K_1} :: e_1\rangle \xhookrightarrow{0::\overline{\delta}}{}^* \langle \Phi_3' \mid \overline{K_3} :: e_3\rangle$ and $\mathsf{sinv}(\Phi_3, \Phi_3')$. Therefore $\langle \Phi_3' \mid \overline{K_3} :: e_3\rangle \xhookrightarrow{0} \langle \Phi_2' \mid \overline{K_1} :: e_1\rangle$. The other conditions follow immediately. ■

Case $\Sigma_2 = [\overline{\Sigma_3}, \overline{\epsilon}]$ **:** We must show that there exists a $\overline{\epsilon'}$ such that $\langle \Phi_1' \mid \overline{K_1}^{\ell_1} :: K[e_1]^{\mathsf{app}}\rangle \xrightarrow{[\overline{\Sigma}, \overline{\epsilon'}]} \langle \Phi_2' \mid \overline{K_2}^{\ell_2} :: e_2{}^{\mathsf{label}([\overline{\Sigma}, \overline{\epsilon}])}\rangle$, $\mathsf{label}_{\mathsf{app}}([\overline{\Sigma}, \overline{\epsilon}]) = \mathsf{app} \Rightarrow \mathsf{sinv}(\Phi_2, \Phi_2')$, and $\mathsf{ct}(\overline{\epsilon}) = \mathsf{ct}(\overline{\epsilon'})$.

We case split on whether $(\mathsf{call}\ z_f)^{\mathsf{app}\to\mathtt{lib}} \in \overline{\epsilon}$. If it isn't then we let $\overline{\epsilon'} = \overline{\epsilon}$. By inversion we have that $\langle \Phi_1 \mid \overline{K_1} :: e_1\rangle \xhookrightarrow{0} \langle \Phi_3 \mid \overline{K_3} :: e_3\rangle \xhookrightarrow{\overline{\delta}}{}^* \langle \Phi_4 \mid \overline{K_4} :: e_4\rangle \xhookrightarrow{\overline{\epsilon}} \langle \Phi_2 \mid \overline{K_2} :: e_2\rangle$. By Lemma 6 and the definitions of $\mathbb{T}$ and $\mathbb{A}$ on the underlying non-speculative subtrace for e_1 we have that $\langle \Phi_1' \mid \bullet :: e_1\rangle \xhookrightarrow{\mathsf{unlabel}(\overline{\epsilon})}{}^* \langle \Phi_1'' \mid \bullet :: v\rangle$ and therefore $\langle \Phi_1' \mid \overline{K_1} :: e_1\rangle \xhookrightarrow{0} \langle \Phi_3' \mid \overline{K_3} :: e_3\rangle$ with $\mathsf{sinv}(\Phi_3, \Phi_3')$. By Lemma 18 we have that $\langle \Phi_1' \mid \overline{K_1} :: e_1\rangle \xhookrightarrow{0::\overline{\delta}}{}^* \langle \Phi_4' \mid \overline{K_4} :: e_4\rangle$ and $\mathsf{sinv}(\Phi_3, \Phi_3')$. Our goals then follow immediately.

If $(\mathsf{call}\ z_f)^{\mathsf{app}\to\mathtt{lib}} \in \overline{\epsilon}$, then by Lemma 6 we have that there is some underlying non-speculative trace $\overline{\epsilon'}$ such that $\mathsf{ct}(\overline{\epsilon}) = \mathsf{ct}(\overline{\epsilon'})$ which contains the entire call into the library. By the definition of classical constant time, Lemma 4, and $\mathsf{sinv}(\Phi_1, \Phi_1')$, this call leaves no traces in unprotected memory (and therefore no reads or writes to unprotected memory may be invalidated). Therefore we may once again apply the same reasoning to get that $\langle \Phi_1' \mid \overline{K_1}^{\ell_1} :: K[e_1]^{\mathsf{app}}\rangle \xrightarrow{[\overline{\Sigma}, \overline{\epsilon'}]} \langle \Phi_2' \mid \overline{K_2}^{\ell_2} :: e_2{}^{\mathsf{label}([\overline{\Sigma}, \overline{\epsilon}])}\rangle$. ■

We next consider all of the cases where $\mathsf{label}(\overline{\Sigma_1}) = \mathtt{lib}$. By assumption we may let $\langle \Phi.S \mid \bullet :: e^{\mathsf{app}}\rangle \xRightarrow{\overline{\delta_0} \diamond (\mathsf{call}\ z_f)^{\mathsf{app}\to\mathtt{lib}} \diamond \overline{\delta_1}}{}^* \langle \Phi_1.S \mid \overline{K_1}^{\ell_1} :: e_1{}^{\mathsf{label}(\overline{\Sigma_1})}\rangle$ such that $(\mathsf{call}\ z_f')^{\mathsf{app}\to\mathtt{lib}} \notin \overline{\delta_1}$ and $\overline{\delta_0} \diamond (\mathsf{call}\ z_f)^{\mathsf{app}\to\mathtt{lib}} \diamond \overline{\delta_1} = \mathsf{nonspec}(\overline{\Sigma_1})$ and then have that $\langle \Phi'.S \mid \bullet :: e^{\mathsf{app}}\rangle \xRightarrow{\overline{\delta_0'} \diamond (\mathsf{call}\ z_f)^{\mathsf{app}\to\mathtt{lib}} \diamond \overline{\delta_1'}}{}^* \langle \Phi_1'.S \mid \overline{K_1'}^{\ell_1'} :: e_1'{}^{\mathsf{label}(\overline{\Sigma_1})}\rangle$ such that $(\mathsf{call}\ z_f')^{\mathsf{app}\to\mathtt{lib}} \notin \overline{\delta_1'}$ and $\overline{\delta_0'} \diamond (\mathsf{call}\ z_f)^{\mathsf{app}\to\mathtt{lib}} \diamond \overline{\delta_1'} = \mathsf{nonspec}(\overline{\Sigma_1'})$. By the assumption that the attacker is memory safe, $z_f \in \mathsf{dom}(\mathbb{C}_{\mathsf{spec}}(\Gamma \vDash L))$.

Case $\Sigma_2 = (\delta^{\mathtt{lib}}, \mathcal{N})$ **:** Follows by the above assumption, the assumption that the library is classically speculative constant time, and Lemma 15. ■

Case $\Sigma_2 = (\tau^{\texttt{lib}\to\mathsf{app}}, \mathcal{N})$ **:** Follows by the above assumption, the assumption that the library is classically speculative constant time, Lemma 15, and the same reasoning using Lemma 4 as in Lemma 6. ∎

Case $\Sigma_2 = (\overline{\delta}, \mathcal{S})$ **:** Follows by the above assumption, the assumption that the library is classically speculative constant time, Lemma 15, and Lemma 16. ∎

Case $\Sigma_2 = [\overline{\Sigma_3}, \overline{\epsilon}]$ **:** Follows by the above assumption, the assumption that the library is classically speculative constant time, Lemma 15, and Lemma 16. ∎

□

Lemma 20 (FTLR). *If* $(\langle \Phi \mid \bullet :: e^{\mathsf{app}} \rangle, \langle \Phi' \mid \bullet :: e^{\mathsf{app}} \rangle) \in \mathbb{S}_{\mathsf{app}}[\![\overline{\Sigma}]\!]$*, then there exist* $\overline{\Sigma'}$, Φ_1, $\overline{K_1{}^{\ell_1}}$, e_1, Φ'_1, $\overline{K'_1{}^{\ell'_1}}$*, and* e'_1 *such that* $\langle \Phi \mid \bullet :: e^{\mathsf{app}} \rangle \xrightarrow{\overline{\Sigma}}{}^{*} \langle \Phi_1 \mid \overline{K_1{}^{\ell_1}} :: e_1{}^{\mathsf{label}(\overline{\Sigma})} \rangle$ *and* $\langle \Phi' \mid \bullet :: e^{\mathsf{app}} \rangle \xrightarrow{\overline{\Sigma'}}{}^{*} \langle \Phi'_1 \mid \overline{K'_1{}^{\ell'_1}} :: e'_1{}^{\mathsf{label}(\overline{\Sigma})} \rangle$ *and* $\mathsf{ct}(\mathsf{crunch}(\overline{\Sigma})) = \mathsf{ct}(\mathsf{crunch}(\overline{\Sigma'}))$.

Proof. By induction on $\overline{\Sigma}$. □

Theorem 3 ($\mathbb{C}_{\mathsf{spec}}$ guarantees robust speculative constant time). *If* $\Gamma \vDash L$ *is classically speculative constant time for a speculation oracle spec and does not contain any* `protect`$_p$ *subterms, then* $\mathbb{C}_{\mathsf{spec}}(\Gamma \vDash L)$ *is robustly speculatively constant time (for attackers that do not contain* `protect`$_p$*).*

Proof. By Lemma 13, Lemma 19, and Lemma 20. □

D.3 Concurrent protections

Theorem 4 ($\mathbb{C}_{\text{ro-co}}$ guarantees robust constant time for concurrent observers). *If* $\Gamma \vDash L$ *is classically constant time and does not contain any* `protect`$_p$ *subterms, then* $\mathbb{C}_{\text{ro-co}}(\Gamma \vDash L)$ *is robustly constant time for concurrent observers (that do not contain* `protect`$_p$*).*

Theorem 5 ($\mathbb{C}_{\text{spec-co}}$ guarantees robust speculative constant time for concurrent observers). *If* $\Gamma \vDash L$ *is classically speculative constant time for a speculation oracle spec and does not contain any* `protect`$_p$ *subterms, then* $\mathbb{C}_{\text{spec-co}}(\Gamma \vDash L)$ *is robustly speculatively constant time (for attackers that do not contain* `protect`$_p$*).*

Both proofs exactly follow the structure of their non-concurrent counterparts with the addition of maintaining the state invariant during library subtraces as well. This ensures that *at all times* the only memory that varies is protected from concurrent observers.

E Evaluation additional data

	READ-ONLY			SPECULATIVE			
Size	Q_1	**Median overhead**	Q_3	Q_1	**Median overhead**	Q_3	Baseline cycles
1	0.59%	**3.92%**	10.07%	0.63%	**4.01%**	10.18%	7.14e+03
128	0.14%	**1.46%**	5.00%	0.13%	**1.65%**	4.99%	1.60e+04
256	0.01%	**0.80%**	2.67%	0.02%	**0.82%**	2.65%	2.69e+04
512	-0.03%	**0.45%**	1.70%	0.00%	**0.56%**	1.68%	4.93e+04
1024	-0.07%	**0.29%**	0.98%	-0.08%	**0.34%**	1.05%	9.42e+04
2048	-0.26%	**0.16%**	0.65%	-0.15%	**0.15%**	0.71%	1.82e+05

Table 17. aead overhead by size

	READ-ONLY			SPECULATIVE			
Size	Q_1	**Median overhead**	Q_3	Q_1	**Median overhead**	Q_3	Baseline cycles
29	-0.58%	**0.50%**	0.64%	-3.53%	**0.65%**	3.59%	5.67e+05
59	-0.27%	**0.38%**	0.69%	-3.36%	**0.68%**	3.85%	5.65e+05
117	-0.07%	**0.49%**	0.71%	-3.67%	**0.64%**	3.06%	5.59e+05
231	-0.24%	**0.38%**	0.65%	-3.60%	**0.58%**	3.53%	5.50e+05
453	-0.50%	**0.26%**	0.72%	-3.67%	**0.59%**	2.97%	9.03e+05
709	-0.70%	**0.21%**	0.49%	-3.60%	**0.58%**	3.79%	1.13e+06
2711	-0.49%	**0.06%**	0.73%	-3.57%	**0.59%**	3.53%	3.92e+06
4237	-0.57%	**-0.09%**	0.41%	-3.61%	**0.48%**	3.38%	6.09e+06

Table 18. encrypt overhead by size

	READ-ONLY			SPECULATIVE			
Size	Q_1	**Median overhead**	Q_3	Q_1	**Median overhead**	Q_3	Baseline cycles
29	-1.09%	**-0.12%**	0.96%	-0.57%	**0.27%**	1.83%	2.99e+08
59	-1.20%	**-0.21%**	0.67%	-1.05%	**-0.08%**	1.27%	2.99e+08
117	-0.98%	**-0.22%**	0.48%	-0.65%	**-0.04%**	1.55%	3.08e+08
231	-1.08%	**-0.22%**	0.50%	-0.62%	**0.05%**	1.88%	3.09e+08
453	-1.16%	**-0.29%**	0.48%	-0.96%	**0.01%**	1.71%	3.09e+08
709	-1.13%	**-0.26%**	0.48%	-0.96%	**-0.04%**	1.39%	3.09e+08
2711	-1.22%	**-0.22%**	0.48%	-0.99%	**-0.09%**	1.28%	2.99e+08
4237	-1.00%	**-0.11%**	0.67%	-1.00%	**-0.22%**	0.97%	3.e+08

Table 19. sign overhead by size