AUTOMATICALLY ELIMINATING SPECULATIVE LEAKS WITH BLADE

MARCO VASSENA, KLAUS V. GLEISSENTHALL, RAMI GÖKHAN KICI, DEIAN STEFAN, AND RANJIT JHALA

Abstract We introduce BLADE, a new approach to automatically and efficiently synthesizing provably correct repairs for transient execution vulnerabilities like Spectre. BLADE is built on the insight that to stop speculative execution attacks, it suffices to cut the dataflow from expressions that speculatively introduce secrets (sources) to those that leak them through the cache (sinks), rather than prohibiting speculation altogether. We formalize this insight in a static type system that (1) types each expression as either transient, i.e., possibly containing speculative secrets or as being stable, and (2) prohibits speculative leaks by requiring that all sink expressions are stable. We introduce protect, a new abstract primitive for fine grained speculation control that can be implemented via existing architectural mechanisms, and show how our type system can automatically synthesize a minimal number of protect calls needed to ensure the program is secure. We evaluate BLADE by using it to repair several verified, yet vulnerable WebAssembly implementations of cryptographic primitives. BLADE can fix existing programs that leak via speculation automatically, without user intervention, and efficiently using two orders of magnitude fewer fences than would be added by existing compilers, and thereby ensuring security with minimal performance overhead.

1 Introduction

Implementing secure cryptographic algorithms is hard. The code must not only be functionally correct, memory safe, and efficient, it must also avoid divulging secrets indirectly through side channels like control-flow, memory-access patterns, or execution time. Consequently, much recent work focuses on how to ensure implementations do not leak secrets e.g., via type systems [12, 43], verification [4], and program transformations [6].

Unfortunately, these efforts are foiled by speculative execution. Even if secrets are closely controlled via guards and access checks, the processor can simply ignore those checks when executing speculatively. Even though these speculative windows are short, an attacker can exploit them to leak secrets in turn.

In principle, memory fences block speculation, and hence, offer a way to recover the original security guarantees. In practice, however, fences pose a confounding dilemma. Programmers can restore security by conservatively inserting fences after every load, but then endure the huge performance costs. Alternatively, they can rely on heuristic approaches for inserting fences [41], but forgo guarantees about the absence of side-channels. Since missing even only one fence can allow an attacker to leak the whole process memory, Chrome isolates web sites in different processes to avoid the risk altogether [9].

In this paper, we introduce BLADE, a new, fully automatic approach to provably and efficiently eliminate speculation-based leakage. BLADE is based on the key insight that to prevent leaking data via speculative execution, it is unnecessary to stop all speculation as done by traditional memory fences. Instead, it suffices to cut the data flow from expressions (sources) that speculatively introduce secrets to those that leak them through the cache (sinks). We develop this insight into an automatic enforcement algorithm via four contributions.

1. A Semantics for Speculation. Our first contribution is a formal operational semantics for a simple While language, inspired by a previous formal model of a speculative and out-of-order processor [13]. Our semantics captures the essence of speculation-based attacks by modeling precisely how speculation can occur and what an attacker can observe via speculation (§ 3). To prevent leakage, we propose and formalize the semantics of an abstract primitive called protect that captures the essence of several hardware mechanisms proposed in recent work [2, 36]. Crucially, and in contrast to a regular fence which stops all speculation, protect only stops speculation for a given variable. For example \( x := \text{protect}(e) \) ensures that \( e \)'s value is only assigned to \( x \) after \( e \) has been assigned its stable, non-speculative value. Ideally and in the long run, protect should be implemented in hardware. However commodity processors do not support fine grained control of speculation, therefore, as a short term solution, we implement protect in software via speculative load hardening (SLH) [11].

2. A Type System for Speculation. Our second contribution is an approach to conservatively approximating the dynamic semantics of speculation via a static type system that types each expression as either transient (\( T \)), i.e., expressions that may contain speculative secrets, or stable (\( S \)), i.e., those that cannot (§ 4.1). Our system prohibits speculative leaks by requiring that all sink expressions
that can influence intrinsic attacker visible behavior (e.g., cache addresses) are typed as stable. In fact, the type system does not rely on user-provided security annotations to identify sensitive sources and public sinks. Instead, our system conservatively rejects programs that exhibit any flow of information from transient sources to stable sinks, and therefore can detect speculative vulnerabilities also in legacy unannotated cryptographic code. We connect the static and dynamic semantics by proving that well-typed programs are indeed secure (Â§ 5), i.e., satisfy a correctness condition called speculative non-interference [18] which states that the program does not leak under speculative execution more than it would under sequential execution.

3. Automatic Protection. Existing programs that are free of protect statements are likely insecure under speculation and will be rejected by our type system. Thus, our third contribution is an algorithm that automatically synthesizes a minimal number of protect statements to ensure that the program satisfies speculative non-interference (Â§ 4.2). To this end, we extend the type checker to construct a def-use graph that captures the data-flow between program expressions. The presence of a path from transient sources to stable sinks in the graph indicates a potential speculative leak in the program. To repair the program, we only need to identify a cut-set, a set of variables whose removal eliminates all the leaky paths in the graph. We show that inserting a protect statement for each variable in a cut-set suffices to yield a program that is well-typed, and hence, secure with respect to speculation (§5.3). Happily, finding such cuts is an instance of the classic max-flow/min-cut problem, so existing polynomial time algorithms let us efficiently synthesize protect statements that resolve the dilemma of enforcing security with minimal performance overhead.

4. Evaluation. Our final contribution is an implementation of our method in a fully automatic push-a-button tool called BLADE, and an evaluation using BLADE to repair verified yet vulnerable (to transient execution attacks) programs: the WebAssembly implementations of the signal messaging Protocol and its respective cryptographic libraries [32], and a number of verified cryptographic algorithms from [42] (§ 6). WebAssembly (WASM) is an attractive language for writing efficient and secure implementation of cryptographic algorithms because WASM code runs at near native speed, it is portable, and validated, and hence immune to classic stack-smashing attacks [19]. Even though BLADE operates on WASM code, we believe that our approach can scale to other programming languages as well. Our evaluation shows that BLADE can automatically compute fixes for existing programs without requiring user annotations. Compared to an existing fully automatic protection as implemented in existing compilers (notably Clang), BLADE inserts without user intervention two orders of magnitude fewer fences and thus imposes negligible performance overhead.

2 Overview

In this section, we present two potential speculative execution vulnerabilities in HACL*— a verified cryptographic library — that were discovered by BLADE and discuss how BLADE repairs the vulnerabilities by inserting protect statements. We then show how BLADE computes the repairs via our minimal fence inference algorithm and finally how BLADE proves that the repairs are indeed correct, via our transient-flow type system.

2.1 Two Speculation Bugs and Their Fixes

Figure 1 shows a code fragment from a function in the implementation of the SHA2 hash in HACL*. Though BLADE operates on WebAssembly, we present equivalent simplified C code for readability. The function takes as input a pointer input_len, validates the input (line 3), loads from memory the public length of the hash (line 4), calculates a target address dst3 (line 5), and finally pads the buffer pointed to by dst3 (line 8).

```c
void SHA2_update_last(int *input_len, ...) {
    if (!valid(input_len)) { ... }
    int len = *input_len;
    int *dst3 = ... + len;
    int *dst3 = protect(. + len);
    ... *dst3 = pad;
    ... }
```

Figure 1. Code fragment from the HACL* SHA2 implementation, containing a potential speculative execution vulnerability that leaks explicitly through the cache by writing memory at a secret-tainted address (line 8). The patch computed by BLADE is shown in green.
address containing secret data (e.g., the secret key used by
the hash function) and calls the function again, this time
with an invalid pointer. As a result of the mistraining, the
branch predictor causes the processor to skip validation
and erroneously load the secret into len, which in turn, is
used to calculate pointer dst3. The buffer pointed to by
dst3 is then written in line 8, completing the attack. Even
though pointer dst3 is incorrect due to miströsion and
the write will therefore be squashed, its side-effects persist,
and therefore remain visible to the attacker. The attacker
can then extract the target address — and thereby the secret — via cache timing measurements [17].

**Preventing the Attack: Memory Fences.** Since the
attack exploits the fact that input validation is specula-
tively skipped, we can prevent it by making sure that the
buffer in line 8 is not written until the input has been
validated. To mitigate these class of attacks, Intel [21] and
AMD [5] recommend inserting a speculation barrier after
critical validation check-points. Following this strategy,
we would place a memory fence right before line 8. This
fence would stop all speculative execution past the fence,
\textit{i.e.}, no statements after the fence are executed until all
previous statements (including input validation) have
been completed. While the effects of the fence prevent the
attack, they are more restrictive than necessary and incur
high performance cost [37].

**Preventing the Attack Efficiently.** We propose an alter-
native way to stop speculation from reaching the write
in line 8 through a new primitive called \texttt{protect}. Rather
than eliminate all speculation, \texttt{protect} only stops spec-
ulation along a particular data-path. We use \texttt{protect}
to patch the program in line 6. Instead of assigning pointer
dst3 directly as in line 5, the expression that computes
the address is guarded by a \texttt{protect} statement. This en-
sures that the value assigned to \texttt{dst3} is always guaranteed
to use \texttt{len}’s final, nonspeculative value. Therefore, writing
to \texttt{dst3} in line 8 prevents any invalid secret-tainted
address from speculatively reaching the store, where it
could be leaked to the attacker.

**Implementation of \texttt{protect}**. Our \texttt{protect} primitive
provides an abstract interface for fine grained control
of speculation. This approach is promising because it
can eliminate speculation-based leaks precisely and only
when needed (indeed a similar API has been proposed in [33]). However, whether \texttt{protect} can eliminate leaks with tolerable runtime overhead depends on its concrete implementation.

Ideally and in the long run, \texttt{protect} should be
implemented in hardware. Unfortunately existing process-
sors provide only coarse grained control over speculation
through memory fence instructions. Nevertheless, recently
proposed microprocessor designs [2, 36] provide new hard-
ware mechanisms to control speculation, which suggests

```c
void SHA2_update_last(int *input_len,...)
{
  if (! valid(input_len)) { ... }
  int len = protect(*input_len);
  ...
  for ( i = 0; i < len + ...)
  dst2[i] = 0;
  ...
}
```

**Figure 2.** SHA2 code fragment containing a potential speculative execution vulnerability that leaks implicitly through a control-flow dependency. The patch computed by BLADE is shown in green.

that \texttt{protect could be implemented efficiently in hard-
ware in the future.}

Alternatively, \texttt{protect could be implemented in soft-
ware and some software mitigations already exist [35].}
However, the adoption of these solutions is challenging
because they require some cooperation between a modi-
\texttt{fied operating system, the compiler and the hardware, as
well as explicit annotations for secret data. Instead, we
propose a self-contained software-only solution based on
Speculative Load Hardening (SLH), software mitigation
deployed in the code generated by Clang [11]. At a high
level, SLH closes speculative leaks by stalling individual ar-
ray reads until the corresponding bounds-check condition
gets resolved. To support our proposal, we formalize the
ideal hardware-based and the software SLH-based seman-
tics of \texttt{protect} (Âğ3.4) and prove that both primitives
can eliminate speculation-based leaks. Then, we evaluate
how many hardware-based and SLH-based \texttt{protect}
statements are needed to patch HACL* and measure the
runtime performance overhead of the software implementa-
tion of \texttt{protect in Section 6.}

2. Leaking Through a Control-Flow Dependency.
Figure 2 shows a code fragment taken from the same func-
tion as in Figure 1. The code contains a second potential
vulnerability, but in contrast to Figure 1 the vulnerability
leaks secrets implicitly, through a control-flow depen-
dency.

The function reads from memory a (public) integer
\texttt{len} (line 4), which determines the number of initializa-
tion rounds in the condition of the for-loop (line 6). Like
the previous vulnerability, the function is harmless under
sequential execution, but leaks under speculation. As be-
fore, the attacker manipulates the pointer \texttt{input_len} to
point to a secret after mistraining the branch predictor to
skip validation. But instead of leaking the secret directly
through the data cache, the attacker can leak the value indirectly through a control-flow dependency, e.g., via the instruction cache and non-secret dependent lines of the data cache. In particular, the secret determines how often the initialization loop (line 6) is executed during speculation, and therefore an attacker can make secret dependent observations via instruction- and data-cache timing attacks. Like in the previous example, BLADE eliminates the vulnerability by placing a \texttt{protect} statement in line 4.

### 2.2 Computing

#### Fixes Via Minimal Fence Inference

BLADE automatically infers the placement of these \texttt{protect} statements. We illustrate this process using a simple running example Ex1 shown in Figure 3. The code reads two values from an array \((x := a[i_1] \text{ and } y := a[i_2])\), adds them \((z := x + y)\), and indexes another array with the result \((w := b[z])\). We assume that all array operations are implicitly bounds-checked and thus no explicit validation code is needed.

Like the examples above, Ex1 contains a speculative execution vulnerability: the array reads may skip their bounds check and so \(x\) and \(y\) can contain transient secrets \((i.e.,\) secrets introduced by misspecification). This secret data then flows to \(z\), and finally leaks through the data cache by the array read \(b[z]\).

#### Def-Use Graph.

To secure the program, we need to cut the dataflow between the array reads which could introduce \textit{transient} secret values into the program, and the index in the array read where they are leaked through the cache. For this, we first build a \textit{def-use graph} whose nodes and directed edges capture the data dependencies between the expressions and variables of a program. For example, consider (a subset of) the def-use graph of program Ex1 in Figure 4. In the graph, the edge \(x \rightarrow x + y\) indicates that \(x\) is used to compute \(x + y\). To track how transient values propagate in the def-use graph, we extend the graph with the special circle node \(T\), which represents the source of \textit{transient} values of the program. Since reading memory creates transient values, we connect the \(T\) node to all nodes containing expressions that explicitly read memory, e.g., \(T \rightarrow a[i_1]\). Following the data dependencies along the edges of the def-use graph, we can see that node \(T\) is transitively connected to node \(z\), which indicates that \(z\) can contain transient data at runtime. To detect insecure uses of transient values, we then extend the graph with the special circle node \(S\), which represents the sink of \textit{stable} \((i.e.,\) non-transient) values of a program. Intuitively, this node draws all the values of a program that \textit{must} be stable to avoid transient execution attacks. Therefore, we connect all expression used as array indices in the program to the \(S\) node, e.g., \(z \rightarrow S\). The fact that the graph in Figure 4 contains a \textit{path} from \(T\) to \(S\) indicates that transient data flows through data dependencies into (what should be) a stable index expression and thus the program may be leaky.

#### Cutting the Dataflow.

In order to make the program safe, we need to cut the data-flow between \(T\) and \(S\) by introducing \texttt{protect} statements. This problem can be equivalently restated as follows: find a cut-set, \(i.e.,\) a set of variables, such that removing the variables from the graph eliminates all paths from \(T\) to \(S\). Each choice of cut-set defines a way to repair the program: simply add a \texttt{protect} statement for each variable in the set. Figure 4 contains two choices of cut-sets, shown as dotted lines. The cut-set on the left requires two protect statements, for variables \(x\) and \(y\) respectively, corresponding to the \textit{green} patch in Figure 3. The cut-set on the right is \textit{minimal}, it requires only a single protect, for variable \(z\), and corresponds to the \textit{green} patch in Figure 3. Intuitively, minimal cut-sets are preferable because they result in patches that introduce as few \texttt{protect} as needed and therefore allow more speculation safely. Luckily, the problem of finding a minimal cut-set is a classic graph problem called Min-Cut/Max-Flow, for which efficient polynomial-time algorithms exist [1]. For simplicity, BLADE applies the basic version of this algorithm and synthesizes patches that include a minimal number of \texttt{protect}, regardless of their position in the code and how many times they can be executed. We can extend this algorithm to consider also additional criteria \((e.g.,\) proximity to the sink node and the presence of loops\) when searching for a minimal cut set.

---

1. We omit some irrelevant nodes and edges of the graph for readability.

---

**Figure 3.** Ex1: Running Example. The optimal patch computed by BLADE is shown in green. A sub-optimal patch is shown in orange.

**Figure 4.** Subset of the def-use graph of Ex1. The dashed lines identify two valid choices of cut-sets. The left cut requires removing two nodes and thus inserting two \texttt{protect} statements. The right cut shows a minimal solution, which only requires removing a single node.
2.3 Proving

Correctness via Transient-Flow Types

Technically, BLADE needs only the def-use graph of a program to detect vulnerabilities and synthetize patches. However, it is convenient to use a type system to show that patched programs are provably secure, i.e., they satisfy a semantic security condition. Clearly, the type system simplifies the security analysis because one can reason about program execution directly rather than through generic flows of information in the def-use graph. Furthermore, restricting the security analysis only to the type system makes the security proofs independent from the specific algorithm used to compute the program repairs (e.g., the Max-Flow/Min-Cut algorithm). As long as the repaired program type checks, the formal guarantees of BLADE remain valid even if we apply different algorithms to compute the patches (e.g., the variations on the Max-Flow/Min-Cut problem mentioned above). To show that the patches obtained from cutting the def-use graph of a given program are correct (i.e., they make the program well-typed), our transient-flow type system constructs its def-use graph from the type-constraints generated during type inference.

Typing Judgement. The type system statically assigns a transient-flow type to each variable: a variable is typed as transient (written as $T$), if it can contain transient data (i.e., potential secrets) at run-time, and as stable (written as $S$), otherwise. Given a typing environment $\Gamma$ which assigns a transient flow type to each variable, and a command $c$, the type system defines a judgement $\Gamma \vdash c$ saying that $c$ is free of speculative execution bugs. The type system enforces that transient expressions may not be used in positions that may leak their value by affecting memory reads and writes, e.g., they may not be used as array indices and in loop conditions. Additionally, it requires that transient expressions may not be assigned to stable variables, except through the use of protect. To show that our type system indeed prevents speculative execution attacks, we define a semantics for speculative execution of a while language (Section 3) and prove that well-typed programs do not leak speculatively more than sequentially, that is by executing their statements in-order and without speculation (see Section 5).

Type Inference. Given an input program, we construct the corresponding def-use graph by collecting the type constraints generated during type inference. Type inference is formalized by a typing-inference judgment $\Gamma, \text{Prot} \vdash c \Rightarrow k$, which extends the typing judgment from above with (1) a set of protected variables Prot (the cut-set), and (2) a set of type-constraints $k$ (the def-use graph). At a high level, type inference has 3 steps: (i) generate a set of constraints under an initial typing environment and protected set that allow any program to type-check, (ii) construct the def-use graph from the constraints and find a cut-set (the final protected set), and (iii) compute the final typing environment which types the variables in the cut-set as stable. To characterize the security of a still un repaired program after type inference, we define a typing judgment $\Gamma, \text{Prot} \vdash c$, where unprotected variables are explicitly accounted for in the Prot set. Intuitively, the program is secure if we promise to insert a protect statement for each variable in Prot.

To repair programs, we simply honor this promise and insert a protect statement for each variable in the protected set of the typing judgment obtained above. Once repaired, the program type checks under an empty protected set and with the final typing environment.

2.4 Attacker Model

Before moving to the details of our semantics and transient type system, we discuss the attacker model considered in this work. The attacker runs cryptographic code on a speculative out-of-order processor and, as usual, can choose the values of public inputs and observe public outputs, but may not read secret data (e.g., cryptographic keys) in registers and memory. Additionally, the attacker can influence how programs are speculatively executed through the branch predictor and choose the instructions execution order in the processor pipeline. The effects of these actions are observable through the cache and are otherwise invisible at the ISA level. In particular, while programs run, the attacker can take precise timing measurements through the data- and instruction-cache with a cache-line granularity, which may disclose secret data covertly. These features allow the attacker to mount Spectre-PHT attacks [22, 23] and leak data through FLUSH+RELOAD [47] and PRIME+PROBE [38] cache side-channels attacks. We do not consider speculative attacks that rely on the Return Stack Buffer (e.g., Ret2Spec [27] and [24]), Branch Target Buffer (Spectre-BTB [23]), or the Store To Load optimization (Spectre-STL [10]). We similarly do not consider attacks that do not use the cache to exfiltrate data, e.g., port contention (SMoTherSpectre [7]) and Meltdown attacks [10, 26], since hardware fixes address them.

3 A Semantics for Speculation

We now formalize the concepts presented in the overview. We start by giving a formal semantics for a while language with speculative execution. Figure 5 presents the language’s surface syntax. Values consist of Booleans $b$, pointers $n$ represented as natural numbers, and arrays $a$. Array length and base address are given by functions $\text{length}(\cdot)$ and $\text{base}(\cdot)$. In addition to variable assignments, pointer dereferences, array stores, conditionals and loops, our language features a special command that helps prevent

\footnote{The judgment $\Gamma \vdash c$ is just a short-hand for $\Gamma, \varnothing \vdash c$.}
Configurations and Reduction Relation. We formally specify our semantics as a reduction relation between processor configurations. A configuration \(\langle \text{is}, \text{cs}, \mu, \rho \rangle\) consists of a queue of in-flight instructions \(\text{is}\), a stack of commands \(\text{cs}\), a memory \(\mu\), and a map \(\rho\) from variables to values. A reduction step \(C \xrightarrow{d} C'\) denotes that, under directive \(d\), configuration \(C\) is transformed into \(C'\) and generates observation \(o\). To execute a program \(c\) with initial memory \(\mu\) and variable map \(\rho\), the processor initializes the configuration with an empty reorder buffer and inserts the program into the command stack, i.e., \(\langle [], [], \mu, \rho \rangle\). Then, the execution proceeds until both the reorder buffer and the stack in the configuration are empty, i.e., we reach a configuration of the form \(\langle [], [], \mu', \rho' \rangle\), for some final memory store \(\mu'\) and variable map \(\rho'\).

We now discuss the semantics rules of each execution stage and then those for our security primitive.

### 3.1 Fetch Stage

The fetch stage flattens the input command into a sequence of instructions which it stores in the reorder buffer. Figure 7 presents selected rules; the remaining rules are in Appendix A. Rule [Fetch-Seq] pops command \(c_1; c_2\) from the commands stack and pushes the two subcommands for further processing. [Fetch-Asgn] pops an assignment from the commands stack and appends the corresponding processor instruction \(x := e\) at the end of the reorder buffer. Rule [Fetch-Ptr-Load] is similar and simply translates pointer dereferences to the corresponding load instruction. Arrays provide a memory-safe interface to read and write memory: the processor injects bounds-checks when fetching commands that read and write arrays. For example, rule [Fetch-Array-Load] expands command \(x := e_1[e_2]\) into the corresponding pointer

---

**Figure 5.** Surface Syntax.

**Figure 6.** Processor Syntax.
dereference, but guards the command with a bounds-check condition. First, the rule generates the condition \( e = e_2 < \text{length}(e_1) \) and calculates the address of the indexed element \( e' = \text{base}(e_1) + e_2 \). Then, it replaces the array read on the stack with command \texttt{if e then x := *e' else fail} to abort the program and prevent the buffer overrun if the bounds check fails. Later, we show that speculative out-of-order execution can simply ignore the bounds check guard and cause the processor to transiently read memory at an invalid address. Rule [\texttt{Fetch-If-True}] fetches a conditional branch from the stack and, following the prediction provided in directive \texttt{fetchtrue}, speculative branch \( e \) will evaluate to true. Thus, the processor inserts the corresponding instruction \texttt{guard} with a fresh guard identifier \( p \) in the reorder buffer and pushes the then-branch \( c_1 \) onto the stack \( cs \). Importantly, the guard instruction stores the else-branch together with a copy of the current commands stack (i.e., \( cs : = cs + [i] \)) as a rollback stack to restart the execution in case of misprediction.

### 3.2 Execute Stage

In the execute stage, the processor evaluates the operands of instructions in the reorder buffer and rolls back the program state whenever it detects a misprediction.

**Transient Variable Map.** To evaluate operands in the presence of out-of-order execution, we need to take into account how previous, possibly unresolved assignments in the reorder buffer affect the variable map. In particular, we need to ensure that an instruction cannot execute if it depends on a preceding assignment whose value is still unknown. To update variable map \( \rho \) with the pending assignments in reorder buffer \( is \), we define a function \( \phi(is, \rho) \), called the transient variable map. The function walks through the reorder buffer, registers each resolved assignment instruction \( x := v \) in the variable map (through function update \( \rho(x \mapsto v) \)) and marks variables from pending assignments (i.e., \( x := e \), \( x := \text{load}(e) \), and \( x := \text{protect}(r) \)) to \texttt{undefined \( (\rho(x \mapsto \bot)) \)}, making their respective values unavailable to following instructions.

**Execute Rule and Auxiliary Relation.** Step rules for the reduction relation are shown in Fig. 8. Rule [\texttt{EXECUTE}] executes the \( n \)-th instruction in the reorder buffer, following the directive \texttt{exec n}. For this, the rule splits the reorder buffer into prefix \( is_1 \), \( n \)-th instruction \( i \) and suffix \( is_2 \). Next, it computes the transient variable map \( \phi(is_1, \rho) \) and executes a transition step under the new map using an auxiliary relation \( \rightsquigarrow \). Notice that [\texttt{EXECUTE}] does not update the store or the variable map (the transient map is simply discarded). These changes are performed later in the retire stage.

The rules for the auxiliary relation are shown in Fig. 8. The relation transforms a tuple \( \langle is_1, is_2, cs, cs' \rangle \) consisting of prefix, suffix and current instruction \( i \) into a tuple \( \langle is', cs' \rangle \).
specifying the reorder buffer and command stack obtained by executing \( i \). For example, rule \([\text{EXEC-ASGN}]\) evaluates the right-hand side of the assignment \( x := e \) where \([e]^{\rho}\) denotes the value of \( e \) under \( \rho \). The premise \( v = [e]^{\rho} \) ensures that the expression is defined \( i.e., \) it does not evaluate to \( \bot \). Then, the rule substitutes the computed value into the assignment \( (x := v) \), and reinserts the instruction back into its original position in the reorder buffer.

**Guards and Rollback.** Rules \([\text{EXEC-BRANCH-OK}]\) and \([\text{EXEC-BRANCH-MISpredict}]\) resolve guard instructions. In rule \([\text{EXEC-BRANCH-OK}]\), the predicted and computed value of the guard expression match, and the processor only has to replace the guard with a \( \text{nop} \). In contrast, in rule \([\text{EXEC-BRANCH-MISpredict}]\) the predicted and computed value differ \( ([e]^{\rho} \neq b) \). This causes the processor to revert the program state and issue a rollback observation. For the rollback, the processor discards the instructions past the guard \( i.e., \) \( is_2 \) and substitutes the current commands stack \( cs \) with the rollback stack \( cs' \) which causes execution to revert to the alternative branch.

**Loads.** Rule \([\text{EXEC-LOAD}]\) executes a memory load. The rule computes the address \( (n = [e]^{\rho}) \), retrieves the value at that address from memory \( (\mu(n)) \) and rewrites the load into an assignment \( (x := \mu(n)) \). Inserting the assignment into the reorder buffer allows transiently forwarding the loaded value to later instructions. The premise \( \text{store}(_{\_}) \notin is_1 \) prevents the processor from reading stale data from memory: if the load aliases with a preceding (but pending) store, ignoring the store would produce a stale read. To record that the load is issued *speculatively*, the observation \( \text{read}(n, ps) \) stores list \( ps \) containing the identifiers of the guards still pending in the reorder buffer. Function \( (is) \) simply extracts the identifiers of the guard instructions in the buffer \( is \).

### 3.3 Retire Stage

The retire stage removes completed instructions from the reorder buffer and propagates their changes to variable map and memory store. While instructions are executed out-of-order, they are retired in-order to preserve the illusion of sequential execution to the user. For this reason, the rules for the retire stage in Figure 9 always remove the first instruction in the reorder buffer. For example, rule \([\text{RETIRE-NOP}]\) removes \( \text{nop} \) from the front of the reorder buffer. Rules \([\text{RETIRE-ASGN}]\) and \([\text{RETIRE-STORE}]\) remove the resolved assignment \( x := v \) and instruction \( \text{store}(n,v) \) from the reorder buffer and update the variable map \( (\rho[x \mapsto v]) \) and the memory store \( (\mu[n \mapsto v]) \) respectively. Rule \([\text{RETIRE-FAIL}]\) aborts the program by emptying reorder buffer and command stack and generates a \( \text{fail} \) observation, simulating a processor raising an exception \( i.e., \) a page fault.

| Rule | Premise | Conclusion |
|------|---------|------------|
| **Retire-Nop** | \( \langle \text{nop} : is, cs, \mu, \rho \rangle \) | \( \langle \text{re} : is, cs, \mu, \rho \rangle \) |
| **Retire-Asgn** | \( \langle x := v : is, cs, \mu, \rho \rangle \) | \( \langle \text{re} : is, cs, \mu[\mu[x \mapsto v] \mapsto \rho] \rangle \) |
| **Retire-Store** | \( i \Rightarrow \text{store}(n,v) \) | \( \langle \text{re} : is, cs, \mu, \rho \rangle \) |
| **Retire-Fail** | \( \langle \text{fail} : is, cs, \mu, \rho \rangle \) | \( \langle \text{fail} : [] \rangle \rangle \rangle ) \cup \langle \mu, \rho \rangle \) |

**Figure 9. Retire stage.**

**Example.** We demonstrate how the attacker can leak a secret from program Ex1 (Fig. 3) in our model. First, the attacker instructs the processor to fetch all the instructions, suppling prediction \( \text{true} \) for all bounds-check conditions. Figure 10 shows the resulting buffer and how it evolves after each attacker directive, which instruct the processor to speculatively execute the load instructions and the assignment (but not the guard instructions). Memory \( \mu \) and variable map \( \rho \) are shown on the right. Directive \( \text{exec} \ 2 \) executes the first load instruction by computing the memory address \( 2 = [[\text{base}(a)] + i_1]^{\rho} \) and replacing the instruction with the assignment \( x := \mu(2) \) containing the loaded value. Directive \( \text{exec} \ 4 \) transiently reads array \( a \) past its bound, at index 2, reading into the memory \( (\mu(3) = 42) \) of secret array \( s[0] \) and generates the corresponding observation. Finally, the processor forwards the values of \( x \) and \( y \) through the *transient* variable map \( \rho[x \mapsto \mu(2), y \mapsto \mu(3)] \) to compute their sum in the fifth instruction, \( (z := 42) \), which is then used as an index in the last instruction and leaked to the attacker via observation \( \text{read}(42,[1,2,3]) \).

### 3.4 Protect

Next, we turn to the rules that formalize the semantics of \( \text{protect}(\cdot) \) as an ideal hardware primitive and then its software implementation via SLH.

**Protect in Hardware.** Instruction \( x := \text{protect}(r) \) assigns the value of \( r \), only after all previous guard instructions have been executed, \( i.e., \) when the value has become stable and no more rollbacks are possible. Figure 11 formalizes this intuition. Rule \([\text{FETCH-PROTECT-EXPR}]\) fetches protect commands involving simple expressions \( (x := \text{protect}(e)) \) and inserts the corresponding protect instruction in the reorder buffer. Rule \([\text{FETCH-PROTECT-ARRAY}]\) piggy-backs on the previous rule by splitting a protect of an array \( \text{read}(x := \text{protect}(e_1[e_2])) \) into a separate
Reorder Buffer

|   | exec 2 | exec 4 | exec 5 | exec 7 |
|---|--------|--------|--------|--------|
| 1 | guard((x < length(a)) \rightarrow \{fail\}, 1) | exec 2 | exec 4 | exec 5 | exec 7 |
| 2 | x := load(base(a) + x) | x := µ(2) |      |      |      |
| 3 | guard((y < length(a)) \rightarrow \{fail\}, 2) | y := µ(3) |      |      |      |
| 4 | y := load(base(a) + y) |      |      |      |      |
| 5 | z := x + y |      |      |      |      |
| 6 | guard((z < length(b)) \rightarrow \{fail\}, 3) | z := 42 |      |      |      |
| 7 | w := load(base(b) + z) |      |      |      |      |

Observations: read(2,[1]), read(3, [1,2]), \( \epsilon \), read(42,[1,2,3])

Figure 10. Leaking execution of running example Ex1.

### Variable Map

|   | Memory Layout | Variable Map |
|---|---------------|--------------|
|   | \( \mu(0) = 0 \) | \( \rho(i_1) = 1 \) |
|   | \( b[0] = 0 \) | \( \rho(i_2) = 2 \) |
|   | \( a[0] = 0 \) |      |
|   | \( a[1] = 0 \) |      |
|   | \( s[0] = 42 \) |      |

### Semantics of Protect

- **Fetch-Protect-Array**
  \[ c = (x := \text{protect}(e_1[e_2])) \]
  \[ c_1 = (x := e_1[e_2]) \]
  \[ c_2 = (x := \text{protect}(e)) \]
  \[ (is,c:cs,\mu,\rho) \xrightarrow{\text{fetch}} (is, e_1 : e_2 : cs, \mu, \rho) \]

- **Fetch-Protect-Expr**
  \[ c = (x := \text{protect}(e)) \]
  \[ i = (x := \text{protect}(e)) \]
  \[ (is,c:cs,\mu,\rho) \xrightarrow{\text{fetch}} (is + i : [i], cs, \mu, \rho) \]

- **Exec-Protect**
  \[ i = (x := \text{protect}(e)) \]
  \[ v = [e]_\rho \quad i' = (x := \text{protect}(v)) \]
  \[ (is_1,i,is_2,cs) \xrightarrow{\mu,\rho} (is_1 + i' : [i'] + is_2, cs) \]

- **Exec-Protect2**
  \[ i = (x := \text{protect}(v)) \]
  \[ \text{guard}(\_ , \_ ) \notin is_1 \quad i' = (x := v) \]
  \[ (is_1,i,is_2,cs) \xrightarrow{\mu,\rho} (is_1 + i' : [i'] + is_2, cs) \]

Figure 11. Semantics of \( \text{protect}(\cdot) \) as a hardware primitive (selected rules).

### Example
Consider again Ex1 and the execution shown in Figure 10. In the repaired program, \( x + y \) is wrapped in a \( \text{protect} \) statement. As a result, directive \( \text{exec} 5 \) produces value \( z := 42 \), instead of \( z := 42 \) which prevents instruction 7 from executing (as its target address is undefined), until all guards are resolved. This in turn prevents the leaking of the transient value.

**Protect in software.** SLH rewrites array reads by injecting artificial data-dependencies between bound checks and the corresponding addresses in load instructions, thus transforming control-flow dependencies into data-flow dependencies. These data-dependencies validate control-flow decisions at runtime by stalling speculative loads until the processor resolves their bounds check conditions. Formally, we replace rule [Fetch-Protect-Array] with rule [Fetch-Protect-SLH] in Figure 12. The rule computes the bounds check condition \( e = e_2 < \text{length}(e_1) \), the target address \( e' = (base(e_1) + e_2 \), and generate commands that abort the execution if the check fails, like for regular array reads. Additionally, the rule generates regular commands that (i) assign the result of the bounds check to a reserved variable \( r \) (\( c_1 = r := e \)), (ii) conditionally update the variable with a bitmask consisting of all 1s or 0s (\( c_2 = r := r \oplus \overline{r} \)), and (iii) mask off the target address with the bitmask (\( c_3 = x := (e' \& r) \)). Since the target address in command \( c_3 \) depends on register \( r \), the processor cannot read memory until the bounds check is resolved. If the check succeeds, the bitmask \( r = \overline{r} \) leaves the target address unchanged \( ([e']^\rho = [e' \& r]_\rho) \) and the processor reads the correct address normally. Otherwise,
the bitmask \(r = 0\) zeros out the target address and the processor loads speculatively only from the constant address \(0 = \llbracket c & \mathbb{B} \rrbracket^\rho\). (We assume that the processor reserves the first memory cell and initializes it with a dummy value, e.g., \(\mu(0) = 0\).) Notice that this solution works under the assumption that the processor does not evaluate the conditional update \(r := r \oplus \Gamma; 0\) speculatively. We can easily enforce that by compiling conditional updates to non-speculative instructions available on commodity processors (e.g., the conditional move instruction CMOV on x86).

**Example.** Consider again Ex1. The optimal patch protect\((x + y)\) cannot be executed on existing processors without support for a generic protect(·) primitive. Nevertheless, we can repair the program by applying SLH to the individual array reads, i.e., \(x := \text{protect}(a[i_1])\) and \(y := \text{protect}(a[i_2])\).

## 4 Type System and Inference

In Section 4.1, we present a transient-flow type system which statically rejects programs that can potentially leak through transient execution attacks. Given an annotated program, we apply constraint-based type inference [3, 30] to generate its use-def graph and reconstruct type information (Section 4.2). Then, reusing off-the-shelf Max-Flow/Min-Cut algorithms, we analyze the graph and locate potential speculative vulnerabilities in the form of a variable min-cut set. Finally, using a simple program repair algorithm we patch the program by inserting a minimum number of protect so that it cannot leak speculatively anymore (Figure 14).

### 4.1 Type System

Our type system assigns a transient-flow type to expressions and tracks how transient values propagate within programs, rejecting programs in which transient values reach commands which may leak them. An expression can either be typed as stable (S) indicating that it cannot contain transient values during execution, or as transient (T) indicating that it can. These types form a 2-point lattice [25], which allows stable expressions to be typed as transient, but not vice versa, i.e., we define a can-flow-to relation \(\sqsubseteq\) such that \(S \sqsubseteq T\), but \(T \not\sqsubseteq S\).

**Typing Expressions.** Given a typing environment for variables \(\Gamma \in \text{Var} \rightarrow \{S, T\}\), the typing judgement \(\Gamma \vdash r : \tau\) assigns a transient-flow type \(\tau\) to \(r\). Figure 13 presents selected rules (see Appendix C for the rest). The shaded part of the rules generates type constraints during type inference and are explained later. Values can assume any type. Variables are assigned their respective type from the environment. Rule [Bop] propagates the type of the operands to the result of binary operators \(\oplus \in \{+, \ast\}\). Finally, rule [Array-Read] assigns the transient type to array reads as the array may potentially be indexed out of bounds during speculation. Importantly, the rule requires the array index to be stable to prevent programs from leaking through the cache.

**Typing Commands.** Given a set of protected variables Prot, we define a typing judgment \(\Gamma, \text{Prot} \vdash c\) for commands. Intuitively, a command \(c\) is well-typed under environment \(\Gamma\) and set Prot, if \(c\) does not leak, under the assumption that the expressions assigned to all variables in Prot are protected using the protect(·) primitive. Figure 13b shows our typing rules. Rule [Asgn] disallows assignments from transient to stable variables (as \(T \not\sqsubseteq S\)).

![Figure 13. Transient flow type system and type constraints generation (selected rules).](image-url)
Rule [PROTECT] relaxes this policy as long as the right-hand side is explicitly protected.\(^3\) Intuitively, the result of \texttt{protect}(\(\cdot\)) is stable and it can thus flow securely to variables of any type. Rule [ASGN-PROT] is similar, but instead of requiring an explicit \texttt{protect}(\(\cdot\)) statement, it demands that the variable is accounted for in the protected set Prot. This is secure because all assignments to variables in Prot will eventually be protected through the repair function discussed later in this section.

**Implicit Flows.** To prevent programs from leaking data implicitly through their control flow, rule [IF-THEN-ELSE] requires the branch condition to be stable. This might seem overly restrictive, at first: why can’t we accept a program that branches on transient data, as long as it does not perform any attacker-observable operations (e.g., memory reads and writes) along the branches? Indeed, classic information-flow control (IFC) type systems (e.g., [40]) take this approach by keeping track of an explicit program counter label. Unfortunately, such permissiveness is un-sound under speculation. Even if a branch does not contain observable behavior, the value of the branch condition can be leaked by the instructions that follow a mispredicted branch. In particular, the rollback caused by a misprediction may cause to repeat load and store instructions after the mispredicted branch, thus revealing whether the attacker guessed the value of the branch condition.

**Example.** Consider the following program: \(\text{if } tr \text{ then } \{ x := 0 \} \text{ else } \{ \text{skip} \}; y := a[0] \). The program can leak the value of \(tr\) during speculative execution. To see that, assume that the processor predicts that \(tr\) will evaluate to \texttt{true}. Then, the processor speculatively executes the then-branch (\(x:=0\)) and the load instruction (\(y:=a[0]\)), before resolving the condition. If \(tr\) is \texttt{true}, the memory trace of the program contains a single read observation. However, if \(tr\) is \texttt{false}, the processor detects a misprediction, restarts the execution from the other branch (\texttt{skip}) and executes the array read, producing a rollback and two read observations. From these observations, an attacker could potentially make inferences about the value of \(tr\). Consequently, if \(tr\) is typed as \(T\), our type system rejects the program as unsafe.

4.2 Type Inference

We now present our type inference algorithm.

**Constraints.** We start by collecting a set of constraints \(k\) via typing judgement \(\Gamma, \text{Prot} \vdash s \Rightarrow k\). For this, we define a dummy environment \(\Gamma^*\) and protected set \(\text{Prot}^*\), such that \(\Gamma^*, \text{Prot}^* \vdash c \Rightarrow k\) holds for any command \(c\), (i.e., we let \(\Gamma^* = \lambda x. S\) and include all variables in the cut-set) and use it to extract the set of constraints \(k\). The syntax for constraints is shown in Figure 22. The constraints relate atoms which represent the unknown type of variables, \(i.e., \alpha_x\) for \(x\), and expression, \(i.e., r\). Constraints record can-flow-to relationships between the atoms and lattice values \(T\) and \(S\). They are accumulated via operator \(\cup\), where we identify \(k_1 \cup \ldots \cup k_n\) with the set \{\(k_1, \ldots, k_n\)\}.

**Solutions and Satisfiability.** We define the solution to a set of constraints as a function \(\sigma\) from atoms to flow types, \(i.e., \sigma \in \text{ATOMS} \rightarrow \{T,S\}\), and extend solutions to map \(T\) and \(S\) to themselves. For a set of constraints \(k\) and a solution function \(\sigma\), we write \(\sigma \vdash k\) to say that the constraints \(k\) are satisfied under solution \(\sigma\). A solution \(\sigma\) satisfies \(k\), if all can-flow-to constraints hold, when the atoms are replaced by their values under \(\sigma\). We say that a set of constraints \(k\) is satisfiable, if there is a solution \(\sigma\) such that \(\sigma \vdash k\).

**Def-Use Graph & Paths.** The constraints generated by our type system give rise to the def-use graph of the type-checked program. For a set of constraints \(k\), we call a sequence of atoms \(a_1...a_n\) a \texttt{path} in \(k\), if \(a_i \subseteq a_{i+1} \in k\) for \(i \in \{1,...,n-1\}\) and say that \(a_1\) is the path’s entry and \(a_n\) its exit. A \(T,S\) path is a path with entry \(T\) and exit \(S\). A set of constraints \(k\) is satisfiable if and only if there is no \(T,S\) path in \(k\), such as a path would correspond to a derivation of \texttt{false}. If \(k\) is satisfiable, we can compute a solution \(\sigma(k)\) by letting \(\sigma(k)(a) = T\), if there is a path with entry \(T\) and exit \(a\), and \(S\) otherwise.

**Cuts.** If a set of constraints is unsatisfiable, we can make it satisfiable by removing some of the nodes in its graph or equivalently protecting some of the variables. A set of atoms \(A\) \texttt{cuts} a path \(a_1...a_n\), if some \(a \in A\) occurs along the path, \(i.e., \exists a \in A \land i \in \{1,...,n\}\) such that \(a_i = a\). We call \(A\) a cut-set for a set of constraints \(k\), if \(A\) cuts all \(T,S\) paths in \(k\). A cut-set \(A\) is minimal for \(k\), if all other cut-sets \(A'\) contain as many or more atoms than \(A\), \(i.e., \#A \leq \#A'\).

**Extracting Types From Cuts.** From a set of variables \(A\) such that \(A\) is a cut-set of constraints \(k\), we can extract a typing environment \(\Gamma(k,A)\) as follows: for an atom \(\alpha_x\), we define \(\Gamma(k,A)(x) = T\), if there is a path with entry \(T\) and exit \(\alpha_x\) in \(k\) that is not cut by \(A\), and let \(\Gamma(k,A)(x) = S\) otherwise.

**Proposition 1 (Type Inference).** If \(\Gamma^*, \text{Prot}^* \vdash c \Rightarrow k\) and \(A\) is a set of variables that cut \(k\), then \(\Gamma(k,A), A \vdash s\).

**Remark.** To infer a repair using exclusively SLH-based \texttt{protect} statements, we simply restrict our cut-set to only include variables that are assigned from an array read.

**Example.** Consider again EX1 in Figure 3. The graph defined by the constraints \(k\), given by \(\Gamma^*, \text{Prot}^* \vdash \text{EX1} \Rightarrow k\) is shown in Figure 4, where we have omitted \(\alpha\)-nodes. The constraints are not satisfiable, since there are \(T,S\) paths. Both \(\{x,y\}\) and \(\{z\}\) are cut-sets, since they cut each \(T,S\) path, however, the set \(\{z\}\) contains only one element and

\(^3\) Readers familiar with information-flow control may see an analogy between \texttt{protect} and the \texttt{declassify} primitive of some IFC languages [29].
Atom \[ a ::= \alpha | r \]
Constraint \[ k ::= a \subseteq S \mid T \subseteq a \mid a \subseteq a \mid k \cup k \mid \emptyset \]
Solution \[ \sigma \in \text{ATOMS} \Rightarrow \{S,T\} \]

Figure 14. Constraint Syntax.

is therefore minimal. The typing environment \( \Gamma(k,\{x,y\}) \) extracted from the sub-optimal cut \( \{x,y\} \) types all variables as \( S \), while the typing extracted from the optimal cut, i.e., \( \Gamma(k,\{z\}) \) types \( x \) and \( y \) as \( T \) and \( z \), \( i_1 \) and \( i_2 \) as \( S \). By Proposition 2 both \( \Gamma(k,\{x,y\}),\{x,y\} \vdash \text{EX1} \) and \( \Gamma(k,\{z\}),\{z\} \vdash \text{EX1} \) hold.

4.3 Program Repair
As a final step, our repair algorithm \( \text{repair}(c, \text{Prot}) \) traverses program \( c \) and inserts a \( \text{protect}(\cdot) \) statement for each variable in the cut-set \( \text{Prot} \). Since we assume that programs are in static single assignment form, there is a single assignment \( x := r \) for each variable \( x \in \text{Prot} \), and our repair algorithm simply replaces it with \( x := \text{protect}(r) \).

5 Consistency and Security
We now present two formal results about our speculative semantics and the security of the type system. First, we prove that the semantics from Section 3 is consistent with sequential program execution (Theorem 5.1). Intuitively, programs running on our processor produce the same results (with respect to the memory store and variables) as if their commands were executed in-order and without speculation. The second result establishes that our type system is sound (Theorem 5.2). We prove that the type system enforces a security property similar to speculative non-interference [18]: well-typed programs do not leak speculatively more than they leak sequentially. Full definitions and proofs can be found in Appendix D.

**Consistency.** We write \( C \Downarrow_D C' \) for the complete speculative execution of configuration \( C \) to final configuration \( C' \), which generates a trace of observations \( O \) under list of directives \( D \). Similarly, we write \( \langle \mu, \rho \rangle \Downarrow_D (\mu', \rho') \) for the sequential execution of program \( c \) with initial memory \( \mu \) and variable map \( \rho \) resulting in final memory \( \mu' \) and variable map \( \rho' \). To relate speculative and sequential observations, we define a projection function, written \( O_L \), which removes prediction identifiers, rollbacks, and mispredicted loads and stores.

**Theorem 5.1 (Consistency).** For all programs \( c \), initial memory stores \( \mu \), variable maps \( \rho \), and directives \( D \), such that \( \langle \mu, \rho \rangle \Downarrow_D (\mu', \rho') \) and \( \langle [\cdot], [\cdot], \mu, \rho \rangle \Downarrow_D (\langle [\cdot], [\cdot], \mu'', \rho'' \rangle) \), then \( \mu' = \mu'' \), \( \rho' = \rho'' \), and \( O \cong O_L \).

The theorem ensures equivalence of the final memory stores, variable maps, and observation traces from the sequential and the speculative execution. Notice that trace equivalence is up to permutation, i.e., \( O \cong O_L \), because the processor can execute load and store instructions out-of-order.

**Speculative Non-Interference.** Speculative non-interference is parametric in the security policy which specifies that variables and part of the memory are controlled by the attacker [18]. In the following, we write \( L \) for the set of public variables and memory locations that are observable by the attacker. Two variable maps are indistinguishable to the attacker, written \( \rho_1 \approx_L \rho_2 \), if and only if \( \rho_1(x) = \rho_2(x) \) for all \( x \in L \). Similarly, memory stores are related pointwise, i.e., \( \mu_1 \approx_L \mu_2 \) if all \( \mu_1(n) = \mu_2(n) \) for all \( n \in L \).

**Definition 1 (Speculative Non-Interference).** A program \( c \) satisfies speculative non-interference, written \( \text{SNI}_L(c) \), if and only if for all directives \( D \), memory stores and variable maps such that \( \mu_1 \approx_L \mu_2 \) and \( \rho_1 \approx_L \rho_2 \), let \( C_i = \langle [\cdot],[\cdot],\mu_i,\rho_i \rangle \) for \( i \in \{1,2\} \), if \( C_1 \Downarrow_D C_1, C_2 \Downarrow_D C_2 \), and \( O_1 \Downarrow = O_2 \downarrow \), then \( O_1 = O_2 \).

In the definition above, programs leak by producing different observations and starting from memories and variables indistinguishable to the attacker. Speculative non-interference requires showing absence of leaks for the speculative traces \( (O_1 = O_2) \) assuming that the program does not already leak sequentially \( (O_1 \Downarrow = O_2 \downarrow) \). Notice that here we consider syntactic equivalence for the traces because both executions follow the same list of directives. We now present our soundness theorem: well-typed programs satisfy speculative non-interference.

**Theorem 5.2 (Soundness).** For all programs \( c \) and security policies \( L \), if \( \Gamma \vdash c \), then \( \text{SNI}_L(c) \).

Notice that our type system does not use the security policy \( L \), yet it can enforce speculative non-interference for any security policy. Intuitively, this theorem holds because the type system conservatively prohibits all data flows from transient source to stable sinks, regardless of the sensitivity (public or secret) of the source.

We conclude with a corollary that combines all the components of our protection chain (type inference, type checking and automatic repair) and shows that repaired programs satisfy speculative non-interference.

**Corollary 5.3.** For all programs \( c \), there exists a set of constraints \( k \) such that \( \Gamma^*, \text{Prot}^* \vdash c \Rightarrow k \). Let \( A \) be a set of variables that cut \( k \), then \( \text{SNI}_L(\text{repair}(c,A)) \) for all security policies \( L \).

6 Implementation and Evaluation
We now describe our implementation and evaluate BLADE on an implementation of the Signal secure messaging protocol and various cryptographic algorithms. Our evaluation shows that BLADE can secure existing software
systems against speculative execution attacks automatically. Moreover, BLADE introduces two orders of magnitude less fences than a baseline algorithm implemented in Clang. As a result, the repairs computed by BLADE incur only a minimal performance overhead.

6.1 Implementation

We implemented BLADE in 3500 lines of Haskell code. BLADE takes as input a WebAssembly program, computes a repaired program that is safe under speculative execution and verifies its correctness via type-checking. Internally, BLADE proceeds in three stages. First, BLADE converts the WebAssembly program into an intermediate representation similar to the While language in Figure 5. This simplifies further processing as WebAssembly is a stack-based language, i.e., arguments are not represented directly, but instead kept on an argument stack. Second, BLADE builds the use-def graph (§4.1) of the input program, infers a minimal cut-set (§4.2), and computes the repair (§4.3). Finally, in the last stage, BLADE extracts a typing-environment from the use-def graph and type-checks the repaired program (§4). This independent checking step provides extra confidence that the repaired program indeed does not leak more speculatively, than it does sequentially (§5). Source code will be made available under an open source license.

6.2 Evaluation

We evaluate BLADE by answering three questions: (Q1) Can we apply BLADE to secure existing software? (Q2) How many protect statements does BLADE have to insert in order to secure those systems? and (Q3) How do the inserted protect statements affect the runtime performance?

(Q1) Applicability. To evaluate BLADE’s applicability, we run it on crypto code, which is already carefully written to eschew cache-timing side channels. Our benchmarks are taken from two main sources: first, a verified implementation [32] of the Signal messaging protocol [16], and second, verified implementations of several crypto primitives taken from [42]. In particular, our benchmarks consist of

- The messaging algorithm implemented in module Signal Core and common cryptographic constructions implemented in module Signal Crypto and used in Signal.
- The HACL* SHA2 hash, AES block cypher, Curve25519 elliptic curve function, and ED25519 digital signature used in Signal.
- The SALSA20 stream cypher, SHA2 hash, and TEA block cypher from [42].

The original implementations of our benchmarks are provably free from cache and timing side-channel. However, those proofs considered only a sequential execution model and therefore do not account for the speculative execution vulnerabilities addressed in this work.

Results. Table 1 shows the code size in Webassembly text format, and the runtime of BLADE on each benchmark. The runtime includes translation, repair and type-checking. The results are encouraging: the execution time scales proportionally with the code size and the analysis completes fairly quickly, even for large benchmarks (>60k WASM LOC): the runtime is less than 10s for all of our benchmarks.

(Q2) Number of protect. Next, we evaluate how many protect statements the analysis has to insert to make the programs secure. The results are shown in Table 1. Column B contains our baseline, which replaces all non-constant array reads, i.e., reads whose address depends on a variable, with statement stable_read (Section 3.4), implementing a SLH-like mitigation that masks the address with the array bounds-check condition. This is the proposed mitigation in the Clang compiler [11]. Column P shows the number of protect inserted by BLADE. All benchmarks are modified by the baseline, except for TEA, which is a simple, toy encryption algorithm (that should not be used in practice). In particular, for five of the nine programs, BLADE does not need to insert any fences. Column P/B shows the ratio of protect statements to baseline read masks in percent. For most benchmarks, our analysis has to insert under 3% of fences compared to the baseline. For the SHA2 implementation of HACL* this rises to 11.5%. Across all benchmarks, the number of fences is two orders of magnitude lower than the baseline. Since protect statements are an idealized primitive that are not available in todays hardware, we show the number of stable_read primitives that are needed to implement the protect in column S. The table shows that using stable_read primitives requires inserting more fences by a factor of 1.8x, which underlines the benefits of a hardware implementation of protect.

(Q3) Runtime performance overhead. To evaluate the runtime performance impact of our repair, we compared how a naive placement of fences—applying speculative load hardening to every load of a non-constant address—compares against our approach. We picked the SHA2-512 hash function for this test, and used inputs of size 4KB. Naive fence placement introduced 44 fences while ours introduced only 5. Our measurements showed that while the naive repair algorithm caused 13.9% overhead, the overhead of our minimal fence replacement algorithm was only 0.42%. We used a sample size of 500, and found the relative margin of error of our measurements were less than 0.07%.
Transient Execution Attacks. Since Spectre [23] and Meltdown [26] were announced, many transient execution attacks exploiting different microarchitectural components and side-channels have been discovered and new ones come to light at a steady pace. These attacks leak data across arbitrary security boundaries, including SGX enclaves [15, 39], hypervisors and virtual machines [44], and even remotely over a network [34]. We refer to [10] for a comprehensive systematization.

Detection and Repair. Wu and Wang [45] detect cache side channels via abstract interpretation by augmenting the control-flow to accommodate for speculation. Specsectoc [?] and Pitchfork [13] use symbolic execution on x86 binaries to detect speculative vulnerabilities. Cheang et al. [14] and Bloem et al. [8] apply bounded model checking to detect potential speculative vulnerabilities respectively via 4-ways self-composition and taint-tracking. Almost all these efforts [8, 13, 14, 457] consider only in-order execution (except Pitchfork [13]) for a fixed speculation bound, and focus on vulnerability detection but do not propose techniques to repair vulnerable programs. In contrast, our type system enforces speculative non-interference even when program instructions are executed out-of-order with unbounded speculation and automatically synthesizes repairs. Given a set of untrusted input source, oo7 Wang et al. [41] statically analyzes a binary to detect vulnerable patterns and inserts fences in turn. Our tool, BLADE, not only repairs vulnerable programs without user annotation, but ensures that program patches contain a minimum number of fences. Furthermore, BLADE formally guarantees that repaired programs are free from speculation-based attacks.

Speculative Execution Semantics. There have been several recent proposals for speculative execution semantics [13, 14, 287]. Of those, [13] is closest to ours, and inspired our semantics (e.g., we share the 3-stages pipeline, attacker-supplied directives and the instruction reorder buffer). However their semantics targets an assembly language with direct jumps, while we reason about speculative execution of imperative programs with structured control-flow.

Hardware Mitigations and Secure Design. Both AMD AMD [5] and Intel Intel [21] recommend inserting serializing, fence instructions after bounds checks to protect against Spectre v1 attacks and some compilers followed suit [20, 31]. Unfortunately, these defenses cause significant performance degradation [10]. Taram et al. [36] propose context-sensitive fencing, a hardware-based mitigation that dynamically inserts fences in the instruction stream when dangerous conditions arise. Several secure hardware designs have been studied to remove speculative attacks from future processors. InvisiSpec Yan et al. [46] is a new micro-architecture design that features a special speculative buffer to prevent speculative loads from polluting the cache. STT [2] tracks speculative taints inside the processor micro-architecture and prevent speculative values from reaching instructions that could serve as covert channels. We think our approach could be applied to guide such hardware mitigations by pinpointing the program parts that need to be protected.

7 Related Work

Transient Execution Attacks. Since Spectre [23] and Meltdown [26] were announced, many transient execution attacks exploiting different microarchitectural components and side-channels have been discovered and new ones come to light at a steady pace. These attacks leak data across arbitrary security boundaries, including SGX enclaves [15, 39], hypervisors and virtual machines [44], and even remotely over a network [34]. We refer to [10] for a comprehensive systematization.

Detection and Repair. Wu and Wang [45] detect cache side channels via abstract interpretation by augmenting the control-flow to accommodate for speculation. Specsectoc [?] and Pitchfork [13] use symbolic execution on x86 binaries to detect speculative vulnerabilities. Cheang et al. [14] and Bloem et al. [8] apply bounded model checking to detect potential speculative vulnerabilities respectively via 4-ways self-composition and taint-tracking. Almost all these efforts [8, 13, 14, 457] consider only in-order execution (except Pitchfork [13]) for a fixed speculation bound, and focus on vulnerability detection but do not propose techniques to repair vulnerable programs. In contrast, our type system enforces speculative non-interference even when program instructions are executed out-of-order with unbounded speculation and automatically synthesizes repairs. Given a set of untrusted input source, oo7 Wang et al. [41] statically analyzes a binary to detect vulnerable patterns and inserts fences in turn. Our tool, BLADE, not only repairs vulnerable programs without user annotation, but ensures that program patches contain a minimum number of fences. Furthermore, BLADE formally guarantees that repaired programs are free from speculation-based attacks.

Speculative Execution Semantics. There have been several recent proposals for speculative execution semantics [13, 14, 287]. Of those, [13] is closest to ours, and inspired our semantics (e.g., we share the 3-stages pipeline, attacker-supplied directives and the instruction reorder buffer). However their semantics targets an assembly language with direct jumps, while we reason about speculative execution of imperative programs with structured control-flow.

Hardware Mitigations and Secure Design. Both AMD AMD [5] and Intel Intel [21] recommend inserting serializing, fence instructions after bounds checks to protect against Spectre v1 attacks and some compilers followed suit [20, 31]. Unfortunately, these defenses cause significant performance degradation [10]. Taram et al. [36] propose context-sensitive fencing, a hardware-based mitigation that dynamically inserts fences in the instruction stream when dangerous conditions arise. Several secure hardware designs have been studied to remove speculative attacks from future processors. InvisiSpec Yan et al. [46] is a new micro-architecture design that features a special speculative buffer to prevent speculative loads from polluting the cache. STT [2] tracks speculative taints inside the processor micro-architecture and prevent speculative values from reaching instructions that could serve as covert channels. We think our approach could be applied to guide such hardware mitigations by pinpointing the program parts that need to be protected.

References

[1] Flows in Networks. Princeton University Press, 1962.
[2] Speculative taint tracking (stt): A comprehensive protection for speculatively accessed data. In MICRO, 2019.
[3] Alex Aiken. Constraint-based program analysis. In Radhia Cousot and David A. Schmidt, editors, Static Analysis, pages 1–1, Berlin, Heidelberg, 1996. Springer Berlin Heidelberg. ISBN 978-3-540-70674-8.
[4] José Bacelar Almeida, Manuel Barbosa, Gilles Barthe, François Dupressoir, and Michael Emmi. Verifying constant-time implementations. In Usenix Security, 2016.
[5] AMD. Software techniques for managing speculation on AMD processors. https://developer.amd.com/wp-content/resources/Managing-Speculation-on-AMD-Processors.pdf, 2018.
[6] GILLES BARTHÉ, SANDRINE BLAZY, BENJAMIN GREGOIRE, RÉMI HUTIN, VINCENT LAPORTE, DAVID PICARDIE, and ALIX TRIEU. Formal verification of a constant-time preserving c compiler. In POPL, 2020.
[7] Atri Bhattacharyya, Alexandra Sandulescu, Matthias Neuschwander, Alessandro Sorniotti, Babak Falsafi, Matthias Payer, and Anil Kurmus. Smotherspectre: Exploiting speculative execution through port contention. In Proceedings of the 2019 ACM SIGSAC Conference on Computer and Communications Security, CCS ’19, pages 785–800, New York, NY, USA,
[37] Vadim Tkachenko. 20-30% performance hit from the spectre bug fix on ubuntu. https://www.percona.com/blog/2018/01/23/20-30-performance-hit-spectre-bug-fix-ubuntu/, Jan 2018.

[38] Eran Tromer, Dag Arne Osvik, and Adi Shamir. Efficient cache attacks on aes, and countermeasures. J. Cryptol., 23(1):37–71, January 2010. ISSN 0933-2790. doi: 10.1007/s00145-009-9049-y. URL http://dx.doi.org/10.1007/s00145-009-9049-y.

[39] Jo Van Bulck, Marina Minkin, Ofir Weisse, Daniel Genkin, Baris Kasikci, Frank Piessens, Mark Silberstein, Thomas F. Wenisch, Yuval Yarom, and Raoul Strackx. Foreshadow: Extracting the keys to the Intel SGX kingdom with transient out-of-order execution. In Proceedings of the 27th USENIX Security Symposium. USENIX Association, August 2018. See also technical report Foreshadow-NG [44].

[40] D. Volpano, G. Smith, and C. Irvine. A Sound Type System for Secure Flow Analysis. J. Computer Security, 4(3):167–187, 1996.

[41] Guanhua Wang, Sudipta Chattopadhyay, Ivan Gotovchits, Tulika Mitra, and Abhik Roychoudhury. oo7: Low-overhead defense against spectre attacks via binary analysis. CoRR, abs/1807.05843, 2018. URL http://arxiv.org/abs/1807.05843.

[42] Conrad Watt, John Renner, Natalie Popescu, Sunjay Cauligi, and Deian Stefan. Ct-wasm: Type-driven secure cryptography for the web ecosystem. In POPL, 2019.

[43] Conrad Watt, John Renner, Natalie Popescu, Sunjay Cauligi, and Deian Stefan. Ct-wasm: Type-driven secure cryptography for the web ecosystem. Proc. ACM Program. Lang., 3 (POPL):77:1–77:29, January 2019. ISSN 2475-1421. doi: 10.1145/3290390. URL http://doi.acm.org/10.1145/3290390.

[44] Ofir Weisse, Jo Van Bulck, Marina Minkin, Daniel Genkin, Baris Kasikci, Frank Piessens, Mark Silberstein, Raoul Strackx, Thomas F. Wenisch, and Yuval Yarom. Foreshadow-NG: Breaking the virtual memory abstraction with transient out-of-order execution. Technical report, 2018. See also USENIX Security paper Foreshadow [39].

[45] Meng Wu and Chao Wang. Abstract interpretation under speculative execution. In Proceedings of the 40th ACM SIGPLAN Conference on Programming Language Design and Implementation, PLDI 2019, pages 802–815, New York, NY, USA, 2019. ACM. ISBN 978-1-4503-6712-7. doi: 10.1145/3314221.3314647. URL http://doi.acm.org/10.1145/3314221.3314647.

[46] Mengjia Yan, Jiho Choi, Dimitrios Skarlatos, Adam Morrison, Christopher W. Fletcher, and Josep Torrellas. Invisispec: Making speculative execution invisible in the cache hierarchy. In Proceedings of the 51st Annual IEEE/ACM International Symposium on Microarchitecture, MICRO-51, pages 428–441, Piscataway, NJ, USA, 2018. IEEE Press. ISBN 978-1-5386-6240-3. doi: 10.1109/MICRO.2018.00042. URL https://doi.org/10.1109/MICRO.2018.00042.

[47] Yuval Yarom and Katrina Falkner. Flash-reload: A high resolution, low noise, 33 cache side-channel attack. In 23rd USENIX Security Symposium (USENIX Security 14), pages 719–732, San Diego, CA, August 2014. USENIX Association. ISBN 978-1-931971-15-7. URL https://www.usenix.org/conference/usenixsecurity14/technical-sessions/presentation/yarom.
## A Full Calculus

\begin{table}[h]
\begin{center}
\begin{tabular}{|c|}
\hline
\textbf{Fetch-Skip} \\
\langle is, \text{skip} : cs, \mu, \rho \rangle \xrightarrow{\text{fetch}} \langle is++, \text{nop} : cs, \mu, \rho \rangle \\
\hline
\textbf{Fetch-Assign} \\
\langle is, x := e : cs, \mu, \rho \rangle \xrightarrow{\text{fetch}} \langle is++, [x := e] : cs, \mu, \rho \rangle \\
\hline
\textbf{Fetch-Seq} \\
\langle is, c_1 ; c_2 : cs, \mu, \rho \rangle \xrightarrow{\text{fetch}} \langle is, c_1 ; c_2 : cs, \mu, \rho \rangle \\
\hline
\textbf{Fetch-Ptr-Load} \\
\langle is, x := * : cs, \mu, \rho \rangle \xrightarrow{\text{fetch}} \langle is++, [x := \text{load}(e)] : cs, \mu, \rho \rangle \\
\hline
\textbf{Fetch-Ptr-Store} \\
\langle is, * e_1 := e_2 : cs, \mu, \rho \rangle \xrightarrow{\text{fetch}} \langle is++, \text{store}(e_1, e_2) : cs, \mu, \rho \rangle \\
\hline
\textbf{Fetch-Fail} \\
\langle is, \text{fail} : cs, \mu, \rho \rangle \xrightarrow{\text{fetch}} \langle is++, \text{fail} : cs, \mu, \rho \rangle \\
\hline
\textbf{Fetch-Array-Load} \\
\begin{align*}
&c = x := e_1[e_2] \\
&\text{fresh}(p)
\end{align*} \\
\begin{align*}
e &= e_2 < \text{length}(e_1) \\
&\text{fresh}(p)
\end{align*} \\
\begin{align*}
e' &= \text{base}(e_1) + e_2 \\
&c' = \text{if } e \text{ then } x := e' \text{ else fail}
\end{align*} \\
\langle is, c : cs, \mu, \rho \rangle \xrightarrow{\text{fetch}} \langle is, c' : cs, \mu, \rho \rangle \\
\hline
\textbf{Fetch-Array-Store} \\
\begin{align*}
c &= e_1[e_2] := e_3 \\
&\text{fresh}(p)
\end{align*} \\
\begin{align*}
e &= e_2 < \text{length}(e_1) \\
&\text{fresh}(p)
\end{align*} \\
\begin{align*}
e' &= \text{base}(e_1) + e_2 \\
&c' = \text{if } e \text{ then } * e' := e \text{ else fail}
\end{align*} \\
\langle is, c : cs, \mu, \rho \rangle \xrightarrow{\text{fetch}} \langle is, c' : cs, \mu, \rho \rangle \\
\hline
\textbf{Fetch-If-True} \\
\begin{align*}
c &= \text{if } e \text{ then } c_1 \text{ else } c_2 \\
\text{fresh}(p)
\end{align*} \\
\begin{align*}
i &= \text{guard}(e, c_2 : cs, p)
\end{align*} \\
\langle is, c : cs, \mu, \rho \rangle \xrightarrow{\text{fetchtrue}} \langle is++, [i], c_1 : cs, \mu, \rho \rangle \\
\hline
\textbf{Fetch-If-False} \\
\begin{align*}
c &= \text{if } e \text{ then } c_1 \text{ else } c_2 \\
\text{fresh}(p)
\end{align*} \\
\begin{align*}
i &= \text{guard}(e, c_1 : cs, p)
\end{align*} \\
\langle is, c : cs, \mu, \rho \rangle \xrightarrow{\text{fetchfalse}} \langle is++, [i], c_2 : cs, \mu, \rho \rangle \\
\hline
\textbf{Fetch-While} \\
\begin{align*}
c_1 &= c \text{ while } e \\
&c_2 = \text{if } e \text{ then } c_1 \text{ else skip}
\end{align*} \\
\langle is, \text{while} e c : cs, \mu, \rho \rangle \xrightarrow{\text{fetch}} \langle is, c_2 : cs, \mu, \rho \rangle \\
\hline
\end{tabular}
\end{center}
\end{table}

Figure 15. Fetch stage.
Fig. 16. Execute stage.

Fig. 17. Retire stage.

Fig. 18. Helper functions.
Figure 19. Semantics of $\text{protect}(\cdot)$. 
B Semantics of Stable Read

Current processors do not provide a protect primitive instruction nor the means to implement it on top of existing instructions, in its full generality. However, for array reads, it is possible to replicate the effects of protect by exploiting the same data-dependencies tracking capabilities at the core of the processor pipeline. Indeed, Speculative Load Hardening (SLH), a mitigation technique deployed in the code generated by the CLANG compiler, relies on data-dependencies to secure memory loads automatically [11]. Using our formal model, we give rigorous semantics to SLH and show that it can stop transient execution attacks.

At a high level, SLH injects artificial data-dependencies between the conditions used in branch instructions and the addresses loaded in the following instructions to transform control-flow dependencies into data-flow dependencies. Intuitively, these data-dependencies validate control-flow decisions at runtime by stalling speculative loads until the processor resolves the conditions. Under branch conditions, SLH masks the address of loads instructions in such a way that the processor zeroes out the address if the condition is mispredicted, preventing misloads.

To formalize this mechanism, we extend our processor model as follows. We introduce a new processor instruction \( x := e ? : e \_ \text{true} : e \_ \text{false} \), which corresponds to the conditional move instruction CMOV on x86 processors. This instruction simply assigns the value of \( e_1 \) (resp. \( e_2 \)) to variable \( x \), if the condition \( e \) evaluates to true (resp. false). Importantly, this instruction is not subject to speculation: the processor must first evaluate the condition before it can resolve the assignment. We also extend expressions with the standard bitwise AND operator (\&\&) and write \( \overline{0} \) and \( \overline{1} \) for bit words consisting of all 0 and 1. As usual bitmask \( \overline{0} \) and \( \overline{1} \) are respectively the zero and identity element for \&\&, i.e., \( e \&\& \overline{0} = \overline{0} \) and \( e \&\& \overline{1} = e \).

Figure 20 presents the semantics rules for CMOV and for the stable read command implemented using SLH. Rule [Exec-CMOV] evaluates the condition \( (b = [e] \overline{\rho}) \) of the conditional assignment \( x := e ? : e \_ \text{true} : e \_ \text{false} \) and assigns the corresponding expressions \( x := e_b \). Rule [Fetch-Stable-Read-SLH] fetches command \( x := \text{stable\_read}(e_1,e_2) \), computes the bounds check condition, i.e., the address of the indexed element, and push on the stack the following command.

\[
\begin{align*}
\text{if } r \text{ then} & \quad r := r ? \overline{T} \overline{0}; \\
\text{else} & \quad \text{fail}
\end{align*}
\]

The code above is similar to the code generated by a regular array read, but additionally stores the result of the bounds-check condition in reserved variable \( r \). In the then-branch, the condition is then converted into a suitable bitmask using the non-speculative CMOV instruction i.e., \( r := r ? \overline{T} \overline{0} \), which then masks the address loaded, i.e., \( \&((\text{base}(e_1) + e_2) \& r) \). As a result, the value of the address remains undefined until the processor evaluates the bounds check condition. When the condition resolves, if the index is in bounds \( r = \overline{1} \) and the program reads the correct address \( [e \& \overline{T}] \overline{\rho} = [e] \overline{\rho} \) if the index is out-of-bounds, instead, \( r = \overline{0} \) and the load can only read speculatively from a constant address \( (x := \mu(0)) \), thus closing the leak.

Revised Example. Consider again running example EX1 in Figure 3, where instead of standard array reads, we employ the stable\_read(·) primitive from above. After fetching the program, the addresses of the loads are masked with the respective array bounds-check conditions. Assuming the same memory layout and content as in Figure 10 (except for the fact that arrays are shifted by one position since \( \mu(0) = 0 \) is reserved), the processor resolves the first bounds check and reads the array within its bounds, i.e., \( x := \mu(3) = 0 \). The second load attempts to read the array out of bounds \( (y := a[2]) \), and our countermeasure prevents the buffer overrun by redirecting the load to the dummy value stored at address 0. First, the processor resolves the bounds check, i.e., \( r := \overline{0} \), and forwards it to the load \( y := \text{load}((\text{base}(a) + i_2) \& r) \). Then, the condition zero the address and the processor assigns the dummy value to variable \( y \), i.e., \( y := \mu(0) \). As a result, we always read array \( b \) at index \( z = 0 \) and close the leak.

5 We assume that the first memory cell is reserved to the processor, which initializes it with dummy data, e.g., \( \mu(0) = 0 \).
C Full Type System

**Constraints.** Our typing judgement \( \Gamma, \text{Prot} \vdash s \Rightarrow c \) creates a set of constraints \( c \). The syntax for constraints is shown in Figure 22. The constraints relate \( \text{atoms} \) which either represent the unknown type of a variable \( x \) (\( \alpha_x \)), or the unknown type of an expression (\( r \)). Constraints record can-flow-to-relationships between the atoms and lattice values \( T \) and \( S \). They are accumulated via operator \( \cup \), where we identify \( c_1 \cup \ldots \cup c_n \) with the set \( \{ c_1, \ldots, c_n \} \).

**Solutions and Satisfiability.** We define the solution to a set of constraints as a function \( \sigma \) from atoms to flow types, i.e., \( \sigma \in \text{ATOMS} \Rightarrow \{ T, S \} \), and extend solutions to map \( T \) and \( S \) to themselves. For a set of constraints \( c \) and a solution function \( \sigma \), we write \( \sigma \vdash c \) to say that the constraints \( c \) are satisfied under solution \( \sigma \). The definition of \( \sigma \vdash c \) is shown in the lower part of Figure 22. In short, solution \( \sigma \) satisfies \( c \), if all can-flow-to constraints hold, when the atoms are replaced by their values under \( \sigma \). We say that a set of constraints \( c \) is satisfiable, if there is a solution \( \sigma \) such that \( \sigma \vdash c \).

**Paths.** The constraints generated by our type system give rise to the def-use graph of the type-checked program. For a set of constraints \( c \), we call a sequence of atoms \( a_1 \ldots a_n \) a path in \( c \), if \( a_i \subseteq a_{i+1} \in c \) for \( i \in \{ 1, \ldots, n-1 \} \) and say that \( a_1 \) is the path’s entry and \( a_n \) its exit. A \( T\text{-}S \) path is a path with entry \( T \) and exit \( S \). A set of constraints \( c \) is satisfiable if and only if there is no \( T\text{-}S \) path in \( c \), as such a path would correspond to a derivation of false. If \( c \) is satisfiable, we can compute a solution \( \sigma(c) \) by letting \( \sigma(c)(a) = T \), if there is a path with entry \( T \) and exit \( a \), and \( S \) otherwise.

**Cuts.** If a set of constraints is unsatisfiable, we can make it satisfiable by removing some of the nodes in its graph or equivalently protecting some of the expressions. A set of atoms \( A \) cuts a path \( a_1 \ldots a_n \), if some \( a \in A \) occurs along the path, i.e., there exists \( a \in A \) and \( i \in \{ 1, \ldots, n \} \) such that \( a_i = a \). We call \( A \) a cut-set for a set of constraints \( c \), if \( A \) cuts all \( T\text{-}S \) paths in \( c \), and say that \( A \) is minimal for \( c \), if all other cut-sets \( A' \) contain as many or more atoms than \( A \), i.e., \( \#A \leq \#A' \). The problem of finding a minimal cut-set is an instance of them Min-Cut/Max-Flow problem, and we can reuse existing efficient algorithms [11] to compute a solution.

**Extracting Types From Cuts.** From a set of variables \( A \) such that \( A \) is a cut-set of constraints \( c \), we can extract a typing environment \( \Gamma(c,A) \), as follows: for an atom \( \alpha_x \), we define \( \Gamma(c,A)(x) = T \), if there is a path with entry \( T \) and exit \( \alpha_x \) in \( c \) that is not cut by \( A \), and let \( \Gamma(c,A)(x) = S \) otherwise.

**Type Inference.** To infer typing environment \( \Gamma \) and protected set \( \text{Prot} \) for a statement \( s \), we first define a dummy environment \( \Gamma^* \) and protected set \( \text{Prot}^* \), such that \( \Gamma^*, \text{Prot}^* \vdash s \Rightarrow c \) holds for any statement \( s \), and use it to extract the set of constraints \( c \). For this, we define \( \Gamma^* \) as the environment that maps all variables to \( S \), and \( \text{Prot}^* \) the set of all variables. We then compute a minimal set of variables \( A \) such that \( A \) is a cut-set of \( c \), extract environment \( \Gamma(c,A) \) and use \( A \) as protected set. Statement \( s \) is then guaranteed to type check under the inferred environment.

**Proposition 2 (Type Inference).** If \( \Gamma^*, \text{Prot}^* \vdash s \Rightarrow c \) and \( A \) is a set of variables that cut \( c \), then \( \Gamma(c,A), A \vdash s \).

**Remark.** To infer a repair using \texttt{stable_read} instead of \texttt{protect}, we can restrict our cut-set to only include variables that are assigned values from an array read.

**Example.** Consider again Ex1 in Figure 3. The graph defined by the constraints \( c \), given by \( \Gamma^*, \text{Prot}^* \vdash \text{Ex1} \Rightarrow c \) is shown in Figure 4, where we have omitted \( \alpha \)-nodes. The constraints are not satisfiable, since there are \( T\text{-}S \) paths. Both \( \{ x, y \} \) and \( \{ z \} \) are cut-sets, since they cut each \( T\text{-}S \) path, however, the set \( \{ z \} \) contains only one element and is therefore minimal. Finally, environment \( \Gamma(c,\{x,y\}) \) types all variables as \( S \) and \( \Gamma(c,\{z\}) \) types \( x \) and \( y \) as \( T \) and \( z \), \( i_1 \) and \( i_2 \) as \( S \), and by Proposition 2 both \( \Gamma(c,\{x,y\}), \{x,y\} \vdash \text{Ex1} \) and \( \Gamma(c,\{z\}), \{z\} \vdash \text{Ex1} \) hold.

**Example.** Next, consider the following example Ex3.

\[
x := a[i]; \quad b[y] := x; \quad \text{if } 0 \leq x \text{ then } z := y \text{ else skip}
\]

We show the corresponding graph in Figure 23. As before, the constraints are unsatisfiable due to the path from \( T \) to \( S \). The set \( \{ x \} \) is a minimal cut-set producing environment \( \Gamma(c,\{x\}) \) which types all variables as \( S \). Finally, the typing judgement \( \Gamma(c,\{x\}) \vdash \text{Ex1} \) holds, indicating that the program is secure, given the promise that \( x \) will be protected.

C.1 Examples for Repair

**Example.** Consider again Ex1 in Figure 3 from Section 2. The cut-set shown on the right in Figure 4 produces the repair shown in the comments of Figure 3.

**Example.** Consider again Ex2 and its dataflow graph shown in Figure 23. The cut-set \( \{ x \} \) produces the repaired program below.

\[
x := \text{protect}(a[i]); \quad b[y] := x; \quad \text{if } 0 \leq x \text{ then } z := y \text{ else skip}
\]

C.2 Type Inference

Our type-inference approach is based on type-constraints satisfaction. Intuitively, type constraints restrict the types that variables and expressions may assume in a program. In the constraints, the possible types of variables and expressions are represented by \( \text{atoms} \), unknown types of (sub-)expressions and type variables that can be instantiated with any type that satisfies the constraints. Solving these constraints requires finding a substitution, i.e., a mapping from atoms to concrete transient-flow type,
such that all constraints are satisfied if we instantiate the atoms with their type.

Type inference consists of 3 steps: (i) generate a set of constraints under an initial typing environment and protected set that under-approximates the solution of the constraints, (ii) construct the def-use graph from the constraints and find a cut-set, and (iii) cut the transient-to-stable dataflows in the graph and compute the resulting typing environment.

**Constraint Generation.** We describe the generation of constraints through the typing judgment from Figure 13. Given a typing environment $\Gamma$, a protected set $\text{Prot}$, the judgment $\Gamma, \text{Prot} \vdash r \Rightarrow k$ type checks $r$ and generates type constraints $k$. The syntax for constraints is shown in Figure 14. Constraints are sets of can-flow-to relations involving concrete types ($S$ and $T$) and atoms, i.e., type

For space reasons, the constraints generation is reported next to the typing judgment, but we remark that these are two distinct judgments. In particular the type-constraints generation judgment ignores the set of protected variables $\text{Prot}$, but to avoid confusion we include it in the judgment anyway.
automatically eliminating speculative leaks with blade

\[ S(x) = 0 \]

appendix c for a formal account of the mathematical

variables corresponding to program variables (e.g., \( \alpha_x \) for \( x \)) and unknown types for expressions (e.g., \( r \)). In rule \([\text{VAR}]\), constraint \( x \subseteq \alpha_x \) indicates that the type variable of \( x \) should be at least as transitive as the unknown type of \( x \). This ensures that, if variable \( x \) is transient, then \( \alpha_x \) can only be instantiated with type \( T \). Rule \([\text{Bop}]\) generates constraints \( e_1 \subseteq e_1 \oplus e_2 \) and \( e_2 \subseteq e_1 \oplus e_2 \) to reflect the fact that the unknown type of \( e_1 \oplus e_2 \) should be at least as transitive as the (unknown) type of \( e_1 \) and \( e_2 \). Notice that these constraints correspond exactly to the premises \( \tau_1 \subseteq \tau \) and \( \tau_2 \subseteq \tau \) of the same rule. Similarly, rule \([\text{Array-Read}]\) generates constraints \( e_1 \subseteq S \) and \( e_2 \subseteq S \) for the unknown type of the array and the index respectively. In addition to these, the rule generates also the constraint \( T \subseteq e_1[e_2] \), which forces the type of \( e_1[e_2] \) to be transient. Rule \([\text{Asgn}]\) and \([\text{Asgn-Prot}]\) generate the same constraint \( r \subseteq x \) because we ignore the protected set during constraint generation, as explained in the footnote. In contrast, rule \([\text{Protect}]\) does not generate the constraint \( r \subseteq x \) because \( r \) is explicitly protected. In the other rules, the constraints are generated following a similar scheme.

from the set of constraints, we can construct the use-def graph of the program as outlined in section 2.3. We refer appendix c for a formal account of the mathematical construction.

type inference. to perform type inference on a program \( c \), we first generate a set of constraints \( k \) using the judgment described above, with appropriate initial values for the typing environment and the protected set. Specifically, we start with an environment that types all variables as stable, i.e., \( \Gamma^* = \lambda x . S \) and include all variables in the cutset, i.e., \( \text{Prot}_* \vdash c = k \). From the constraints \( k \), we construct the def-graph and compute a cut-set \( \text{Prot} \), e.g., by applying the min-cut/max-flow algorithm. Then, from the cut-set \( \text{Prot} \) and the program \( c \), we compute a substitution that solves the constraints \( k \), as follows. We remove from the graph all nodes in the cut-set \( \text{Prot} \) (and their corresponding edges), and type all variables reachable from node \( T \) as transient, and all other variables as stable. We update the initial typing environment with these type assignments and obtain the resulting environment \( \Gamma \). Under new environment \( \Gamma \) and protected set \( \text{Prot} \), the un repaired program type checks, i.e., \( \Gamma, \text{Prot} \vdash c \).

d proofs

d.1 security

in the following, we write \( \Gamma \vdash C \) to indicate that the program being executed on the processor is well-typed according to the transient-flow type-system.

non-speculative projection of observations. function \( O_1 \) computes the non-speculative projection of observations \( O \). To do that, it applies function \( C(o, ps) \) pointwise. Function \( C(o, ps) \) takes as input a single observation \( o \) and \( ps \), a set of identifiers of mispredicted guards. the function then removes prediction identifiers from observations correctly speculated and replaces mispredicted load, store and rollbacks with the empty observation \( \epsilon \). function \( R(O) \) collects the identifiers of rollbacked guards from events \( \text{rollback}(p) \).

\[ R(O) = \{ p \mid \text{rollback}(p) \}(p) \in O \]

\[ C(\text{load}(n, ps_1), ps_2) \]

\[ C(\text{store}(n, ps_1), ps_2) \]

\[ C(\text{rollback}(p)), ps) = \epsilon \]

\[ C(o, \text{...}) = 0 \]

\[ O_1 = \{ C(o, R(O)) \mid o \in O \} \]

definition 2 (l-equivalence). two configurations \( C_1 = \langle i_1, cs_1, \mu_1, p_1 \rangle \) and \( C_2 = \langle i_2, cs_2, \mu_2, p_2 \rangle \) are l-equivalent, if and only if \( i_1 = i_2, cs_1 = cs_2, \mu_1 \approx L \mu_2, \text{ and } p_1 \approx L p_2 \).

lemma d.1 (l-equivalence 1-step preservation). let \( ps \) be the set of guard identifiers rollbacked in the rest of the execution of well-typed configurations \( \Gamma \vdash C_1 \) and
 two configurations execute the process the same instruction
in the same stage. All the instructions that generate
the empty observation \( e \) or \( \text{fail} \) are trivial. This include
all the rules in the \textit{fetch} and in the \textit{retire} stage. The
interesting cases that can leak occur speculatively and out-of-order, i.e., the \( \rightarrow \) stage. By inspecting rule \textit{EXECUTE},
we notice both configurations execute the same \( n \)-th
instruction from the attacker supplied directive and
with \( L \)-equivalent transient variable maps, \((\phi(is,p_1) \approx_L \phi(is,p_2)) \) from \( p_1 \approx_L p_1 \). Then, we consider the instructions
that can leak during the execute stage: guards \textit{guard} \( (e^b,cs,p) \), loads \textit{load} \( (e) \), and stores \textit{store} \( (n,e) \).
The guard instruction can result in a rollback (rule \textit{EXEC-Branch-Mispredict}) or resolved successfully (rule \textit{EXEC-Branch-Ok}) If the two execution differ the attacker
gains information by observing a \textit{rollback} \( (p) \)
or not (e.g., through the data cache). We show that this
cannot happen because the guard expression is typed \( S \).
We need to prove $\|e\|^{\rho_1} = \|e\|^{\rho_2}$, with $\rho_1 \approx_\mu \rho_2$. If $e$ contains a secret variable, then $\|e\|^{\rho_1} \neq \|e\|^{\rho_2}$, however the secret value would be leaked during sequential execution as well, i.e., it contradicts the hypothesis $C(o1, ps) = C(o2, ps)$. If $e$ contains only public variables, the outcome of the two conditions may still differ. In particular transient secrets may taint public variables and from there transmitted to the condition through the transient function map. However, by the rules of our type system $\Gamma \vdash e : S$, which means that there must be a protect(·) in between the transient source and the stable sink. Since protect(·) forbids values forwarding, the value of the condition is undefined $e(\rho_1) = e(\rho_2) = \text{bot}$ and this case is void.

The reasoning for rules [Exec-Load] and [Exec-Store-Value] is similar.

□

**Theorem D.2 (Soundness).** For all programs $c$, if $\Gamma \vdash c$ then $c$ satisfies speculative non-interference.

**Proof.** Let $\mu_1$ and $\mu_2$ be memories such that $\mu_1 \approx_\mu \mu_2$ and similarly $\rho_1$ and $\rho_2$ variables map such that $\rho_1 \approx_\mu \rho_2$. Let $C_i = ([], [c], \mu_i, \rho_i)$ for $i \in \{1, 2\}$ and let $D$ be a valid schedule such that $C_1 \Downarrow_D C_1'$ and $C_2 \Downarrow_D C_2'$. We now assume $O_1 = O_2$ and show that $O_1 = O_2$ by induction on the typing judgment. The base case ([Done]) is trivial. In the inductive case, we two pairs of small and big steps: reductions. A pair of small-step reductions $\langle is_i, cs_i, \mu_i, \rho_i \rangle \xrightarrow{\alpha} \langle is'_i, cs'_i, \mu'_i, \rho'_i \rangle$ and a pair of big-step reductions $\langle is_i', cs_i', \mu'_i, \rho'_i \rangle \Downarrow_D \langle is''_i, cs''_i, \mu''_i, \rho''_i \rangle$ for $i \in \{1, 2\}$. Assuming that the program does not leak sequentially, we have $o1 = o2$ and $O_1 = O_2$. By induction hypothesis on the big-step we obtain $O_1 = O_2$ and derive $o1 = o2$ by Lemma D.1 applied to small-step reductions and the set of mispredicted guard identifiers $R(O_1)$.

□