# Verifying that a compiler preserves concurrent value-dependent information-flow security

#### 



Data61, CSIRO, Australia UNSW Sydney, Australia Robert.Sison@data61.csiro.au

### Toby Murray

University of Melbourne, Australia toby.murray@unimelb.edu.au

#### Abstract

It is common to prove by reasoning over source code that programs do not leak sensitive data. But doing so leaves a gap between reasoning and reality that can only be filled by accounting for the behaviour of the compiler. This task is complicated when programs enforce value-dependent information-flow security properties—in which classification of locations can vary depending on values in other locations—and complicated further when programs exploit shared-variable concurrency.

Prior work has formally defined a notion of concurrency-aware refinement for preserving valuedependent security properties. However, that notion is considerably more complex than standard refinement definitions typically applied in the verification of semantics preservation by compilers. To date it remains unclear whether it can be applied to a realistic compiler, because there exist no general decomposition principles for separating it into smaller, more familiar, proof obligations.

In this work, we provide such a decomposition principle, which we show can almost halve the complexity of proving secure refinement. Further, we demonstrate its applicability to secure compilation, by proving in Isabelle/HOL the preservation of value-dependent security by a proof-ofconcept compiler from an imperative While language to a generic RISC-style assembly language, for programs with shared-memory concurrency mediated by locking primitives. Finally, we execute our compiler in Isabelle on a While language model of the Cross Domain Desktop Compositor, demonstrating to our knowledge the first use of a compiler verification result to carry an informationflow security property down to the assembly-level model of a non-trivial concurrent program.

2012 ACM Subject Classification Security and privacy → Logic and verification; Security and privacy  $\rightarrow$  Information flow control; Software and its engineering  $\rightarrow$  Compilers

Keywords and phrases Secure compilation, Information flow security, Concurrency, Verification

Supplement Material The Isabelle/HOL theories are available at https://covern.org/itp19.html.

Funding Robert Sison: Australian Government RTP Scholarship & Data61 Research Project Award

Acknowledgements We would like to thank our anonymous reviewers, as well as Carroll Morgan, Kai Engelhardt, Gerwin Klein, Christine Rizkallah, Matthew Brecknell, Johannes Åman Pohjola, and Qian Ge, for their very helpful feedback on earlier versions of this paper.

#### 1 Introduction

It is well known that program translations of the kind carried out by compilers can in principle break security properties like confidentiality [12, 2]. Yet source level reasoning about confidentiality remains common [20, 19, 18]. Existing verified compilers like CompCert [15] and CakeML [14] preserve semantics, but semantics preservation alone may be insufficient to preserve confidentiality, especially for shared memory concurrent programs whose threads must guard against timing leaks in order to prevent them manifesting as storage leaks [22].

```
coupling-inv-pres \mathcal{B} \mathcal{R} \mathcal{I} \equiv \forall lc_{1A} \ lc_{1C}. \ (lc_{1A}, lc_{1C}) \in \mathcal{R} \longrightarrow abstract \ \mathcal{B} \ n \ \mathcal{B}_1
(\forall lc'_{1C}. \ lc_{1C} \sim_{\mathbb{C}} \ lc'_{1C} \longrightarrow (lc'_{1A}, lc_{1A} \sim_{\mathbb{A}}^n \ lc'_{1A} \ \wedge \ (lc'_{1A}, lc'_{1C}) \in \mathcal{R} \wedge (lc_{1A}, lc_{1A} \sim_{\mathbb{A}}^n \ lc'_{1A}, lc_{1A}, lc_{2A}) \in \mathcal{B} \wedge lc_{1A} =_{\mathsf{mds}} \ lc_{2A} \wedge (lc_{2A}, lc_{2C}) \in \mathcal{R} \wedge (lc_{1C}, lc_{2C}) \in \mathcal{I} \wedge \mathcal{R} 
(lc_{2A}, lc_{2C}) \in \mathcal{R} \wedge (lc_{1C}, lc_{2C}) \in \mathcal{I} \wedge \mathcal{R} 
(lc_{1C} =_{\mathsf{mds}} \ lc_{2C} \wedge lc_{2A} \sim_{\mathbb{A}}^n \ lc'_{2A} \wedge lc'_{1A} =_{\mathsf{mds}} \ lc'_{2A} \wedge (lc'_{1C} =_{\mathsf{mds}} \ lc'_{2A} \wedge (lc'_{1C}, lc'_{2C}) \in \mathcal{I} \wedge \mathcal{R} 
(lc'_{2A}, lc'_{2C}) \in \mathcal{R} \wedge (lc'_{1C}, lc'_{2C}) \in \mathcal{I}))))
(lc'_{2A}, lc'_{2C}) \in \mathcal{R} \wedge (lc'_{1C}, lc'_{2C}) \in \mathcal{I}))))
```

**Figure 1** Definition, graphical depiction of refinement preservation for secure-refinement (Def. 6)

Supporting secure compilation of programs that must enforce value-dependent security policies poses an additional challenge, because in such policies the sensitivity of a memory location can depend on the values held in other memory locations. Thus, unlike prior work on secure compilation [4], preserving security under refinement requires a refinement relation that is strong enough to preserve those memory contents on which the policy depends.

In prior work [22], we presented a definition for a notion of value-dependent security-preserving refinement that is compositional for concurrent programs: by applying it to each thread individually, one can derive a secure refinement of the concurrent composition.

The essence of this notion of security-preserving refinement (presented fully in Section 2.2) is in its refinement preservation obligation (coupling-inv-pres in Figure 1). Here, the usual square-shaped commuting diagram that is commonly used to depict (semantics-preserving) refinement (Figure 4a) has been replaced by a *cube* (Figure 1). The additional dimension of this cube reflects that it preserves a 2-safety hyperproperty [6] that compares two executions rather than examining a single one. As such, it is significantly more complicated to prove than standard notions of semantics-preserving refinement typical in verified compilation [15, 14].

To date there exist no verified compilers for shared-variable concurrent programs proved to preserve value-dependent information-flow security. We argue that without a *decomposition* principle the cube-shaped refinement notion is too cumbersome to prove for realistic compilers.

In this paper, we tackle the central problem of making our notion of secure refinement applicable to verified secure compilation. Firstly, we present a decomposition principle that makes the cube-shaped notion more tractable. Secondly, we demonstrate its tractability with our major contribution: a machine-checked formal proof of concurrent value-dependent security preservation, for a proof-of-concept compiler.

In Section 3 we present our decomposition principle, which decomposes the cube (Figure 1) into three separate obligations (Figure 4). The first of these is akin to semantics-preserving refinement, while the second and third essentially ensure together that the refinement has not introduced any termination- and timing-leaks.

In Section 4 we show how the decomposition principle can almost halve the effort to prove secure refinement – in this case, of a program that is especially prone to introduced timing leaks because it branches on secrets (a feature not yet allowed by our compiler). There, we present a side-by-side comparison of the proof effort, both with and without the decomposition principle. We find that using it reduces the proof's complexity by 44%.

In Section 5, we present our compiler and its formal verification, as an application of the decomposition principle. This compiler translates concurrent programs written in an imperative While language, with locking primitives for mediating access to shared memory, into a RISC-style assembly language. It does so by compiling each thread individually, and in doing so preserves a formal security property that remains compositional between threads. Furthermore, our compiler demonstrates a way of formalising and proving when it is safe for a compiler to perform optimisations in the presence of concurrency. To ensure that the contents of shared memory locations are preserved under compilation despite potential interference from other threads, our compiler tracks which shared memory locations are stable (free from any such interference). It then makes use of this tracking to avoid redundant loads from stable shared variables safely, that would otherwise be considered unsafe to omit.

All results are mechanised in Isabelle/HOL,<sup>1</sup> and in Section 6 we explain how, in order to validate our theory, we instantiated it so that we could execute our compiler in Isabelle. This enabled us to execute it over a While language model of the Cross Domain Desktop Compositor [5] (CDDC), a concurrent program that enforces information flow control over value-dependently classified input. To our knowledge this is the first proof of information flow security for an assembly-level model of a non-trivial concurrent program, demonstrating the power of verified secure compilation for deriving security properties of compiled code.

# 2 Background and example

We begin by introducing with an illustrative example (Figure 2) the challenges of verifying value-dependent information-flow security in the presence of shared-variable concurrency.

Consider the task of verifying a multithreaded system that manages the user interface (UI) for a dual-personality smartphone, a phone that provides clearly distinguished user contexts (personalities), typically for work versus leisure. Specifically, our task is to verify that it does not leak sensitive information intended only for one of those personalities, which we classify High (Figure 2b), to locations belonging to the other, which we classify Low (Figure 2c).

Here and generally, our *attacker model* is an entity that can read from the system's *untrusted sinks*: some subset of permanently Low-classified locations not subject to synchronisation. In our example, this may include WLAN device registers in a hostile environment.

The smartphone's UI system consists of a number of threads running concurrently with a shared address space, and we aim to verify that as a whole it satisfies the security requirement. But to avoid a state space explosion that is exponential in the number of threads, we must do this *compositionally*: one thread at a time, then combining the results of these analyses.

We focus on a particular worker thread (Figure 2a), the one responsible for sending touchscreen input from the *source* variable to its intended destination.

The first challenge is that the destination depends on which personality the phone is currently providing, which is indicated by the value of *domain*. This is reflected by the classification of *source* being dependent on the value of *domain*: *source* is classified Low exactly when *domain* = LOW (where LOW is a designated constant), and is classified High otherwise. Due to this dependency, *domain* is known as a *control variable* of *source*.

The second challenge is the worker thread runs in a shared address space that might be accessed or modified by other threads, for various purposes. One of these threads may be responsible for maintaining that domain = LOW exactly when the phone indicates it is providing the Low personality (Figure 2c), so the user knows not to type in anything sensitive. Another thread may be responsible for assigning suspended := TRUE when the user turns the phone's screen off, to make the worker stop processing touchscreen input. We may then wish

The wr-compiler totals ~7k lines, and verification + compilation of the 2-thread CDDC model totals ~1.6k lines of Isabelle proof script, excluding whitespace and comments. See "Supplement Material".

```
while TRUE do
    lock(workspace_lock);
    while !suspended do
        lock(source_lock);
        workspace := source;
        /* ... operations on workspace ... */
        if domain = LOW then
             low\_sink := workspace
        else
             high\_sink := workspace;
             workspace := 0
        fi:
        unlock(source_lock)
    od:
    unlock(workspace_lock);
    while suspended do skip od
(a) Input processing worker thread program
```



(b) The phone providing the High personality:  $domain \neq LOW$ , and source is classified High to reflect that the user might type in secrets.



(c) The phone displaying visual indicators that it is providing the Low personality: domain = LOW, and source is classified Low to reflect that we trust the user not to type in secrets.

Figure 2 Example: Touchscreen input processing for a dual-personality smartphone

for workspace to be usable by some other thread—e.g. processing input from a fingerprint scanner—in such a way that it can assume workspace no longer contains any sensitive values.

When we analyse one thread like this worker in terms of our compositional security property (Section 2.1), all of the other threads in the system are trusted to do two things:

- 1. They follow a *synchronisation scheme*: here, if read- or write-access to a certain variable is governed by a lock, they must hold it in order to access the variable in that manner.
- 2. They themselves do not leak values from High-classified locations (we refer to such values themselves as High) to Low-classified locations that are read-accessible to other threads. Note we are proving that the thread we are analysing can be trusted in the same way.

Even under these assumptions, the concurrency gives rise to some tricky considerations.

Firstly, it is important that no thread in the system (including the thread under analysis) modifies any control variables carelessly. For example, writing domain := LOW immediately after the worker reads a High value from source, will cause it to leak to  $low\_sink$ . To prevent this, the worker uses  $source\_lock$ , granting it  $exclusive\ write-access$  to  $source\ and\ domain$ .

Furthermore as noted above, we may want to ensure that a non-attacker-observable location is nevertheless cleared of any sensitive values before being used by another thread. In our example, we classify workspace Low for the analysis to enforce this when the worker is suspended, but as the worker sometimes uses it to process High values, it is important to know workspace is accessible only to the worker during that time. To ensure this, the worker uses workspace\_lock, granting it exclusive read- and write-access to workspace. It is then responsible for clearing it of any High values by the time it releases exclusive read-access.

# 2.1 Concurrent value-dependent noninterference (CVDNI)

Having illustrated the challenges with an example, we now focus on the formalisation of our information-flow security property CVDNI, which we target with our per-thread analysis, and which our compiler preserves. It is defined in terms of two main elements:

- 1. a binary strong low-bisimulation (modulo modes) relation  $\mathcal{B}$  between program configurations, that establishes the required information-flow security property. Like Goguen & Meseguer-style noninterference [10], any states it relates must agree on their "low" portions, and it demands that lock-step execution preserve that correspondence. This section will explain how it is specialised further for shared-variable concurrency.
- 2. a classification function  $\mathcal{L}$  that determines the "low" portion of a program configuration, thus affecting  $\mathcal{B}$ 's requirements. Unlike [10] however,  $\mathcal{L}$  here can depend on values in the program configuration itself, thus expressing dynamic and not just static classifications.

We now present definitions from Section III-2b of our previous work [22] simplified as noted. The theory is parameterised over the type of values Val, a finite set of shared variables Var, and a deterministic evaluation step semantics  $\rightsquigarrow$  between local configurations (of a thread in a concurrent program) each denoted by a triple  $\langle tps, mds, mem \rangle$ :

- tps is the thread-private state, which is permanently inaccessible to the attacker and the other threads. Note that due to this inaccessibility, we allow the user of the theory to parameterise the type of tps, and do not impose any particular structure.
- mds:: Mode ⇒ Var set is the (access) mode state, which is ghost state associating each Mode = {AsmNoW, AsmNoRW, GuarNoW, GuarNoRW} with a set of shared variables. Intuitively, it identifies the set of variables for which the thread currently possesses (or respects) a kind of exclusivity of access granted (or obligated) by a synchronisation scheme. This facilitates compositional, assume-guarantee [11] style reasoning. For example, when our worker thread holds source\_lock, it assumes no other threads write to source or its control variable ({source, domain} ⊆ mds AsmNoW), otherwise it guarantees it does not write to them (GuarNoW). Similarly, holding workspace\_lock it assumes no other threads read or write to workspace (workspace ∈ mds AsmNoRW), and at all other times it makes the corresponding guarantee (GuarNoRW).
- mem :: Mem is shared memory considered potentially accessible to the attacker and other threads. In order to make what is accessible amenable to analysis, we impose the structure  $Mem = Var \Rightarrow Val$ , a total map from shared variable names to their values.

The theory is then further parameterised by the value-dependent classification function  $\mathcal{L} :: Mem \Rightarrow Var \Rightarrow \{ High, Low \}$ , and a function  $\mathcal{C}vars :: Var \Rightarrow Var \ set$  that returns all the control variables of a given variable. In our worker thread example,  $\mathcal{L} \ mem \ x$  gives:

- High when x is high\_sink, meaning high\_sink is classified High at all times.
- when x is source: Low if mem domain = LOW, and High otherwise.
- Low for all other variables x, meaning they are classified Low at all times.

The set  $C = \{y \mid \exists x. \ y \in C \text{vars } x\}$  is then defined to contain all control variables in the system. Thus in our worker thread example,  $C \text{vars } source = \{domain\} \text{ and } C = \{domain\}.$ 

To support compositionality for concurrent programs, the "low" portion demanded to be equal by the analysis is tightened up to be  $modulo\ modes\ -$  it includes non-control variables only if they are assumed to be readable by other threads according to the mode state: readable  $mds\ x \equiv x \notin mds\ AsmNoRW$ . Thus intuitively, the user of the theory should model permanent untrusted output sinks of the whole concurrent program, as variables for which  $\mathcal L$  always returns Low, ungoverned by any synchronisation scheme that the attacker cannot be trusted to follow. (In our example,  $low\_sink$  is untrusted permanently in this way, but workspace is untrusted only when unlocked.) The notion of observational indistinguishability used for the noninterference property is then defined over memories as follows.

▶ **Definition 1** (Low-equivalent memories modulo modes).

```
mem_1 = ^{\mathsf{Low}}_{mds} mem_2 \equiv \\ \forall x. \ x \in \mathcal{C} \ \lor \ \mathcal{L} \ mem_1 \ x = \mathsf{Low} \ \land \ \mathsf{readable} \ mds \ x \longrightarrow mem_1 \ x = mem_2 \ x
```

For this paper, we will use notation  $lc_1 =_{\mathsf{mds}}^{\mathsf{Low}} lc_2$  to lift  $=_{\mathit{mds}}^{\mathsf{Low}}$  to local program configurations, asserting also that  $lc_1$  and  $lc_2$  are modes-equal (have the same mode state). Additionally, we will use notation  $lc_1 =_{\mathsf{mds}} lc_2$  to denote (alone) that  $lc_1$  and  $lc_2$  are modes-equal.

The per-thread compositional security property com-secure asserts the existence of a witness relation  $\mathcal{B}$  for every possible observationally equivalent pair of starting configurations:

▶ **Definition 2** (Per-thread compositional CVDNI property).

```
com-secure (tps, mds) \equiv \forall mem_1 \ mem_2. \ mem_1 =_{mds}^{\mathsf{Low}} mem_2 \longrightarrow (\exists \mathcal{B}. \ \mathsf{strong\text{-low-bisim-mm}} \ \mathcal{B} \ \land \ (\langle tps, mds, mem_1 \rangle, \langle tps, mds, mem_2 \rangle) \in \mathcal{B})
```

where all such witness relations  $\mathcal{B}$  must be a strong low-bisimulation (modulo modes):

```
strong-low-bisim-mm \mathcal{B} \equiv \text{cg-consistent } \mathcal{B} \land \text{sym } \mathcal{B} \land (\forall lc_1 \ lc_2. \ (lc_1, lc_2) \in \mathcal{B} \land \ lc_1 =_{\mathsf{mds}} lc_2 \longrightarrow \ lc_1 =_{\mathsf{mds}}^{\mathsf{Low}} lc_2 \land (\forall lc_1'. \ lc_1 \leadsto lc_1' \longrightarrow \ (\exists lc_2'. \ lc_2 \leadsto lc_2' \land \ lc_1' =_{\mathsf{mds}} lc_2' \land \ (lc_1', lc_2') \in \mathcal{B})))
```

That is,  $\mathcal{B}$  must maintain observational indistinguishability by requiring that all configuration pairs it relates that have the same mode state, are low-equivalent modulo modes.

Furthermore, it must be a *bisimulation* by being symmetric and *progressing to itself*: any step taken by one of the configurations must be able to be matched by a step taken by the configuration related to it, such that the destinations remain related by  $\mathcal{B}$  (and modes-equal).

Finally—and the most crucial element ensuring the property's compositionality for concurrent programs—is the condition that  $\mathcal{B}$  must be **cg-consistent**: closed under globally consistent changes made to memory by other threads, which is to say, changes that preserve low-equivalence and are permitted by the current mode state mds. Specifically, the environment (of other threads) is permitted to change either of variable x's value or its classification only when x is writable: writable mds  $x \equiv x \notin mds$  AsmNoW  $\land x \notin mds$  AsmNoRW.

▶ **Definition 3** (Closedness under globally consistent changes).

```
cg-consistent \mathcal{B} \equiv \forall tps_1 \ mem_1 \ tps_2 \ mem_2 \ mds. (\langle tps_1, mds, mem_1 \rangle, \langle tps_2, mds, mem_2 \rangle) \in \mathcal{B} \longrightarrow (\forall mem_1' \ mem_2'. \ (\forall x. \ (mem_1 \ x \neq mem_1' \ x \ \lor \ mem_2 \ x \neq mem_2' \ x \ \lor  \mathcal{L} \ mem_1 \ x \neq \mathcal{L} \ mem_1' \ x) \longrightarrow \text{writable} \ mds \ x) \ \land \ mem_1' = \underset{mds}{\text{Low}} \ mem_2' \longrightarrow (\langle tps_1, mds, mem_1' \rangle, \langle tps_2, mds, mem_2' \rangle) \in \mathcal{B})
```

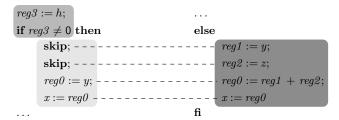
Theorem 3.1 of our prior work [22] then gives us that the parallel composition of com-secure programs is itself a program that enforces a system-wide value-dependent noninterference property (sys-secure, for whose details we refer the reader to Section III-2(a) of [22]).

### 2.2 CVDNI-preserving refinement

Having described the formal security property that we wish to be preserved under refinement (and compilation), we now define formally a suitable notion of secure refinement that preserves it. The proof of CVDNI-preserving refinement for a thread of a concurrent program relies on two binary relations (illustrated by Figure 3) to be nominated by the user of the theory:



(a) Abstract if-conditional. Relation  $\mathcal{R}$  pairs configurations of this program with configurations of the program in Figure 3b that are of the same-shaded region.



(b) Concrete if-conditional. Relation  $\mathcal{I}$  pairs configurations of this program as shown by the dashed lines.

**Figure 3** Excerpts from refinement example [22] that was used to compare proof effort (Section 4).

- 1. a refinement relation  $\mathcal{R}$  relating local configurations of the abstract program to local configurations of the concrete program: abstract must simulate concrete, in a sense typical of much other work on program refinement, including compiler verification efforts.
- 2. a concrete coupling invariant  $\mathcal{I}$  that allows us to use  $\mathcal{B}$  and  $\mathcal{R}$  to build a new strong low-bisimulation (modulo modes) for the concrete program, by discarding unreachable pairs of local configurations after the refinement. It thereby witnesses that any changes a refinement (or compiler) makes to execution time, do not introduce any timing channels.

The essence of the proof technique is to require that a number of conditions—analogous to those for strong-low-bisim-mm—be imposed on the nominated  $\mathcal{R}$  and  $\mathcal{I}$  in relation to a given witness relation  $\mathcal{B}$  establishing CVDNI for the abstract program. The definitions to follow are adapted from Murray et al. [22] Section V. For better readability, we present a simplified version in which no new shared variables are added by the refinement. Consequently we introduce the notation  $=_{\text{mds}}^{\text{mem}}$  to denote that two local configurations have equal mode state and memory, regardless of whether relating configurations of the same or differing languages.

Regarding the maintenance of modes- and observational-equivalence across the relation, the restrictions on refinement are tighter than those that applied to strong-low-bisim-mm. The refinement relation  $\mathcal{R}$  is required to preserve the shared memory in its entirety:

▶ **Definition 4** (Preservation of modes and memory).

```
preserves-modes-mem \mathcal{R} \equiv \forall lc_A \ lc_C. \ (lc_A, lc_C) \in \mathcal{R} \longrightarrow lc_A =_{\mathsf{mds}}^{\mathsf{mem}} \ lc_C
```

Regarding the closedness under changes by other threads that ensures compositionality for concurrency, on  $\mathcal{I}$  we again impose cg-consistent (Definition 3) from Section 2.1. However in the case of  $\mathcal{R}$ , we instead impose closed-others, a simplification of cg-consistent considering only environmental actions that affect the memories on both sides of the relation identically. Furthermore it ensures equality of *all* shared variables, not just those judged observable:

▶ **Definition 5** (Closedness of refinements under changes by others).

```
closed-others \mathcal{R} \equiv \forall tps_A \ tps_C \ mds \ mem \ mem'. (\langle tps_A, mds, mem \rangle_{\mathsf{A}}, \langle tps_C, mds, mem \rangle_{\mathsf{C}}) \in \mathcal{R}) \wedge (\forall x. \ (mem \ x \neq mem' \ x \ \lor \ \mathcal{L} \ mem \ x \neq \mathcal{L} \ mem' \ x) \longrightarrow \text{writable} \ mds \ x) \longrightarrow (\langle tps_A, mds, mem' \rangle_{\mathsf{A}}, \langle tps_C, mds, mem' \rangle_{\mathsf{C}}) \in \mathcal{R})
```

The final major requirement for CVDNI-preservation is then to prove  $\mathcal{R}$  and  $\mathcal{I}$  closed simultaneously under the pairwise executions of the concrete and abstract programs, using

the aforementioned cube-shaped diagram (coupling-inv-pres, Figure 1) whose edges are pairs in  $\mathcal{B}$ ,  $\mathcal{R}$ , and  $\mathcal{I}$ . All that then remains is for the nominated concrete coupling invariant  $\mathcal{I}$  to be symmetric, and the predicate secure-refinement puts together all the requirements:

▶ **Definition 6** (Requirements for secure refinement of the per-thread CVDNI property).

```
secure-refinement \mathcal{B} \ \mathcal{R} \ \mathcal{I} \equiv \text{preserves-modes-mem} \ \mathcal{R} \ \land \ \text{closed-others} \ \mathcal{R} \ \land \ \text{cg-consistent} \ \mathcal{I} \ \land \ \text{sym} \ \mathcal{I} \ \land \ \text{coupling-inv-pres} \ \mathcal{B} \ \mathcal{R} \ \mathcal{I}
```

Theorem 5.1 of our prior work [22] gives us that under the aforementioned conditions,

$$\mathcal{B}_{\mathsf{C}} \text{ of } \mathcal{B} \ \mathcal{R} \ \mathcal{I} \equiv \{ (lc_{1C}, lc_{2C}) \mid \exists lc_{1A} \ lc_{2A}. \ (lc_{1A}, lc_{1C}) \in \mathcal{R} \ \land \ (lc_{2A}, lc_{2C}) \in \mathcal{R} \ \land \ (lc_{1A}, lc_{2A}) \in \mathcal{B} \ \land \ lc_{1C} = \underset{\mathsf{mds}}{\overset{\mathsf{Low}}{\mathsf{pds}}} \ lc_{2C} \ \land \ (lc_{1C}, lc_{2C}) \in \mathcal{I} \}$$

is a witness strong-low-bisim-mm for the concrete program:

strong-low-bisim-mm  $\mathcal{B} \land$  secure-refinement  $\mathcal{B} \mathrel{\mathcal{R}} \mathcal{I} \Longrightarrow$  strong-low-bisim-mm ( $\mathcal{B}_{\mathsf{C}}$  of  $\mathcal{B} \mathrel{\mathcal{R}} \mathcal{I}$ )

# 3 Decomposition principle for CVDNI-preserving refinement

Having presented our previous work [22]'s formalisation of our security property CVDNI and its preservation by refinement, we now present our first contribution: an alternative way of proving secure-refinement (Definition 6) that does away with the use of the cube-shaped, two-sided refinement obligation coupling-inv-pres  $\mathcal{B}$   $\mathcal{R}$   $\mathcal{I}$  (depicted by Figure 1), by decomposing its concerns into (1) proving  $\mathcal{R}$  closed under the pairwise executions of the concrete and abstract programs alone using a square-shaped diagram (depicted by Figure 4a, which is akin to ordinary semantics-preserving refinement), and (2) a number of smaller and more separable obligations gathered together under the side-condition predicate decomp-refinement-safe.

▶ **Definition 7** (Decomposed requirements for CVDNI-preserving secure refinement).

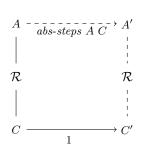
```
secure-refinement-decomp \mathcal{B} \ \mathcal{R} \ \mathcal{I} \ abs\text{-}steps \equiv preserves-modes-mem \mathcal{R} \ \land \ \text{closed-others} \ \mathcal{R} \ \land \ \text{cg-consistent} \ \mathcal{I} \ \land \ \text{sym} \ \mathcal{I} \ \land decomp-refinement-safe \mathcal{B} \ \mathcal{R} \ \mathcal{I} \ abs\text{-}steps \ \land \ (\forall lc_A \ lc_C. \ (lc_A, lc_C) \in \mathcal{R} \longrightarrow (\forall lc'_C. \ lc_C \leadsto_{\mathsf{C}} lc'_C \longrightarrow (\exists lc'_A. \ lc_A \leadsto_{\mathsf{A}}^{(abs\text{-}steps \ lc_A \ lc_C)} \ lc'_A \ \land \ (lc'_A, lc'_C) \in \mathcal{R})))
```

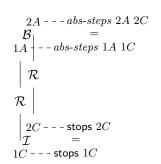
The decomposition requires the provision of a new refinement parameter that we will call abs-steps or the pacing function, whose role is to dictate the pace of the refinement by returning the number of abstract steps that ought to be taken for a single concrete step, for a given abstract-concrete local configuration pair related by  $\mathcal{R}$ . The side-conditions on all of the refinement parameters (depicted by Figures 4b, 4c) are then defined as follows:

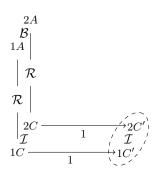
▶ **Definition 8** (Side-conditions for CVDNI-preserving refinement decomposition).

$$\begin{array}{l} \operatorname{decomp-refinement-safe} \ \mathcal{B} \ \mathcal{R} \ \mathcal{I} \ abs\text{-}steps \ \equiv \ \forall lc_{1A} \ lc_{2A} \ lc_{1C} \ lc_{2C}. \ (lc_{1A}, lc_{2A}) \in \mathcal{B} \ \land \\ lc_{1A} =_{\mathsf{mds}} lc_{2A} \ \land (lc_{1A}, lc_{1C}) \in \mathcal{R} \ \land (lc_{2A}, lc_{2C}) \in \mathcal{R} \ \land (lc_{1C}, lc_{2C}) \in \mathcal{I} \land \ lc_{1C} =_{\mathsf{mds}} lc_{2C} \\ \longrightarrow \ \operatorname{stops} \ lc_{1C} = \operatorname{stops} \ lc_{2C} \ \land \ abs\text{-}steps \ lc_{1A} \ lc_{1C} = abs\text{-}steps \ lc_{2A} \ lc_{2C} \ \land \\ (\forall lc'_{1C} \ lc'_{2C}. \ lc_{1C} \leadsto_{\mathsf{C}} lc'_{1C} \ \land \ lc_{2C} \leadsto_{\mathsf{C}} lc'_{2C} \longrightarrow \ (lc'_{1C}, lc'_{2C}) \in \mathcal{I} \ \land \ lc'_{1C} =_{\mathsf{mds}} lc'_{2C}) \end{array}$$

On the intuitive meaning of the side-conditions in Definition 8:







- (a) Refinement preservation for relation  $\mathcal{R}$  under program execution paced by *abs-steps*
- **(b)** Consistency of pacing and stopping behaviour, to prevent timing and termination leaks
- (c) Closedness of the coupling invariant relation  $\mathcal{I}$  under lock-step program execution

**Figure 4** Graphical depictions of refinement decomposition obligations

- stops  $lc_{1C} = \text{stops } lc_{2C}$  ensures that the refinement has not introduced any termination leaks, by asserting *consistent stopping behaviour* for  $\mathcal{I}$ -related concrete program configurations, which we know to be observationally indistinguishable.
- abs-steps  $lc_{1A}$   $lc_{1C} = abs$ -steps  $lc_{2A}$   $lc_{2C}$  ensures that the refinement has not introduced any timing leaks, by asserting consistency of the pace of the refinement for  $\mathcal{R}$ -related program configurations, which we again know to be observationally indistinguishable.
- The final  $\forall$ -quantified clause asserts  $\mathcal{I}$ 's suitability as a coupling invariant, in that it must remain closed under lockstep evaluation of the concrete program configurations it relates. Furthermore it must maintain mode state equality with each lockstep evaluation, which ensures that the refinement has not introduced any inconsistencies in the memory access assumptions and guarantees needed for the concurrent compositionality of the property.

Note the  $\mathcal{B}$ - and  $\mathcal{R}$ -edges in Figure 4c may capture useful facts about a particular program verification technique and compiler, so their availability as assumptions is intended to reduce greatly the effort needed to specify a coupling invariant  $\mathcal{I}$  and prove it satisfies the condition.

Assuming the fulfilment of all of the decomposed requirements, we obtain that they are a sound method for establishing secure refinement of the per-thread CVDNI property:

▶ **Theorem 9** (Soundness of secure-refinement-decomp).

secure-refinement-decomp  $\mathcal{B} \ \mathcal{R} \ \mathcal{I} \ abs\text{-}steps \implies \text{secure-refinement} \ \mathcal{B} \ \mathcal{R} \ \mathcal{I}$ 

In the interests of brevity we relegate proof sketches for all results to Appendices C and D, and for fuller details we refer the reader to our Isabelle/HOL formalisation.

We now devote our attention to two instantiations of this new decomposition principle: (Section 4) for a proof of CVDNI-preservation for the refinement of a program that branches on a secret, and (Section 5.5) for the proof of CVDNI-preservation by a compiler.

# 4 Proof effort comparison

To demonstrate how the decomposition principle reduces proof complexity and effort, we returned to the example refinement discussed in Section V-E of our previous work [22], an excerpt of which is shown in Figure 3. The abstract program (9 imperative commands) branches on a sensitive value, and executes a single atomic expression assignment in each branch. Its refinement (to 16 commands) models expansion of the expressions into multiple steps, resolving a timing disparity between the two branches by padding with **skip**.

We use proof size as a proxy for proof effort, since the former is known to be strongly linearly correlated with the latter [28]. Formalised in Isabelle/HOL as EgHighBranchRevC.thy [21], the proof line count for that theory stood at about 4.6K lines of definitions and proof, of which approx. 3.6K line were proofs. Adapting the proof instead to use the decomposition principle secure-refinement-decomp (Definition 7), the proof line count drops from 3.6K to approx. 2K, a 44% reduction. Regarding definition changes, the new proof makes <10 lines of adaptations to a coupling invariant and pacing function used by the old proof, and adds about 30 lines worth of new helper definitions, for use with the decomposition principle. The rest of the theory and its external dependencies remain in common between the two versions.

As would be expected, the bulk of the deletions are from the full cube-shaped refinement diagram proof (Figure 1) of secure-refinement (Definition 6) for the refinement relation. The surviving parts of that proof just become the square-shaped refinement diagram proof (Figure 4a) of secure-refinement-decomp without much modification. The deletions are replaced by newly added proofs of the three sub-obligations of decomp-refinement-safe (Definition 8).

# 5 The COVERN wr-compiler

Having presented our new decomposition principle for CVDNI-preserving refinement, we now turn to our compiler, whose most notable features for formal proof of secure refinement are:

- 1. Its implementation tracks variable stability (Section 5.4) responsive to use of locking primitives, to know when accesses to shared variables are safe to optimise, and when register contents can be still be considered consistent with shared variable contents.
- 2. Its verification uses a pacing function (Section 5.5.2) and coupling invariant (Section 5.5.3) as the decomposition demands, to ensure it does not introduce timing leaks.

First, we describe its source and target languages, and parameters to the compilation.

#### 5.1 Source language

The COVERN wr-compiler—short for *While-to-RISC compiler*—takes the simple imperative language with while-looping and lock-based synchronisation targeted by the COVERN program logic [20], which we will refer to as While, consisting of the commands *cmd*:

```
exp ::= n \mid v \mid exp \oplus exp

exp := \mathbf{skip} \mid exp \mid \mathbf{if} \ exp \ \mathbf{then} \ exp \ \mathbf{delse} \ exp \ \mathbf{fi} \mid \mathbf{skip} \mid \mathbf{exp} \ \mathbf{do} \ exp \ \mathbf{do} \ \mathbf{do} \mid v := exp \mid \mathbf{dock}(k) \mid \mathbf{unlock}(k)
```

The language is parameterised over a type of values Val, and binary operators  $\oplus :: Val \Rightarrow Val \Rightarrow Val$ . Constants n :: Val; v :: Var and k :: Lock are (resp.) shared program- and lock-variables. The semantics of the locking primitives lock(k) and unlock(k) is informed by a locking discipline provided by the user of the theory as a parameter (see Section 5.3). We leave for future work adding support for pointers and arrays, which we believe will be straightforward because our assume-guarantee framework already provides the means to encode the memory footprint of a command in a way that depends on values in memory.

We assume that the underlying concurrent execution model (e.g. operating system, scheduler) for the While language prevents threads from seeing each others' current program location, and thus (as in previous work [22, 19]) the While program command c :: cmd being executed we model as thread-private state:  $\langle c, mds, mem \rangle_{\mathsf{w}}$ . In contrast, all program variables v :: Var and lock variables k :: Lock reside in the shared memory mem.

### 5.2 Target language

The wr-compiler's target is a generic RISC-style assembly language like that of Tedesco et al. [29] but with lock-based synchronisation primitives added, which we will refer to as RISC:

```
\begin{split} I &:= [l :] B \\ B &:= \mathbf{Load} \ r \ v \mid \mathbf{Store} \ v \ r \mid \mathbf{Jmp} \ l \mid \mathbf{Jz} \ l \ r \mid \mathbf{Nop} \\ \mathbf{MoveK} \ r \ n \mid \mathbf{MoveR} \ r \ r \mid \mathbf{Op} \ \oplus \ r \ r \\ \mathbf{LockAcq} \ k \mid \mathbf{LockRel} \ k \end{split}
```

The language is parameterised over the same value type Val and binary operators  $\oplus$ , shared program variables v::Var and shared lock variables k::Lock as the While language. Presently, direct-addressing **Load** and **Store** instructions (referring to registers r::Reg) are adequate for RISC to implement all existing While features, and we expect adding indirect addressing to RISC to be as straightforward as adding pointer and array support to While.

RISC program texts P are just lists of binary instructions I, each optionally associated with a label l::Lab. We assume that the underlying concurrency model for the RISC language (e.g. OS, scheduler etc.) prevents one thread from reading the program code (instructions) of another,<sup>2</sup> as well as another's registers (including the program counter). Thus, we model the distinguished program counter register's value pc::nat, program text P, and register bank  $regs::Reg \Rightarrow Val$  as thread-private state:  $\langle ((pc, P), regs), mds, mem \rangle_r$ . Apart from this adaptation to our triple format, evaluation semantics follows that of the RISC target of [29].

Finally, like Tedesco et al. [29] we generalise over the (user-supplied) register allocation scheme, and assume there are enough registers to service the maximum depth of expressions in the source program. (More details are available in Appendix D.1.) We leave for future work the modelling and analysis of a compiler phase that spills register contents to memory, in order to make this assumption unnecessary.

#### 5.3 Locking discipline

Like the COVERN logic [20], we assume that the While language program being compiled follows a certain locking discipline, about which the compiler has knowledge, so as to ensure that the RISC program it produces follows the same discipline.

The user of the theory provides the details of the locking discipline in the form of a *lock interpretation* parameter: lock-interp ::  $Lock \Rightarrow (Var\ set \times Var\ set)$ , which for each lock gives the two non-overlapping sets of program variables over which acquiring the lock grants exclusive permission to write, (resp.) read and write. These permissions are then reflected in the way the semantics of the While and RISC locking primitives act on the mode state.

Regarding lock interpretations and the way they interact with the user-provided valuedependent classification function  $\mathcal{L}$  (see Section 2.1), we inherit a few cleanliness conditions from that earlier work [20], chief of which are that lock variables k cannot be control variables,

As is usual for program analyses, we omit any explicit modelling of the microarchitectural state used by superscalar processors (like CPU caches, and state relied on by speculative and out-of-order execution, on whose behaviour attacks like Spectre [13] and Meltdown [16] relied). We argue however that our present assumptions are reasonable under two circumstances: when there is no such state (e.g. on microcontrollers like AVR [7]), or when such state is correctly partitioned by the underlying hardware [30] or the OS [8] – if the hardware allows it [9]! In the latter case, our analysis assumes that microarchitectural state footprints are partitioned according to thread (for memory containing program text) and according to classification by  $\mathcal{L}$  (for shared memory), and furthermore that each value-dependently classified region is given a distinct partition that is flushed on reclassification.

a lock variable k governing access to a program variable v must govern the same kind of access to all of v's control variables, and  $\mathcal{L}$  must classify all lock variables as Low.

### 5.4 Compiler implementation and tracking of shared variable stability

We chose as a starting point the compilation scheme of [29], on the basis of their preserving a noninterference property that like ours exhibits resilience to changes made by an environment—in their case, intended for fault-resilience. Aiming to repurpose that for shared-variable concurrency, we adapted it to Isabelle, implementing it as a primitive recursive function:

compile-cmd :: 
$$CompRec \Rightarrow Lab \ option \Rightarrow Lab \Rightarrow cmd \Rightarrow$$
  
 $(I \times CompRec) \ list \times Lab \ option \times Lab \times CompRec \times bool$ 

where we choose Lab = nat for RISC instruction labels, and the *compilation record* type CompRec is bookkeeping maintained by the compiler that we will describe further below.

A typical invocation to compile a While program c :: cmd takes the form:

$$(PCs, l', nl', C', failed) = \text{compile-cmd } C \ l \ nl \ c$$
 (1)

Here, compile-cmd takes an *initial compilation record* C, an optional *entry label* l, and the *next available label nl*, and for the benefit of the next invocation returns an optional *exit label* l' if one is used by the program just compiled, the *new next available label nl'*, and a *final compilation record* C'. We leave details of label allocation and its impact on achieving sequential composability for compiled RISC programs to Appendix D.2.

In addition to the output RISC program  $P :: I \ list$  itself, a call to compile-cmd also outputs every CompRec associated with the state of the program just before executing every instruction in P. These are returned zipped up together with P as the CompRec-annotated  $RISC \ program \ PCs :: (I \times CompRec) \ list. (P \ can trivially be recovered as map fst <math>PCs$ .) Finally, compile-cmd may return True for failed to reject the input program, such as when it detects a data race (see below), or if expression depth exceeds the assumed limit (Section 5.2).

In the style of the compilation scheme on which it was based [29], the wr-compiler maintains a register record  $\Phi :: reg \rightharpoonup exp$ , i.e. a partial map of registers to expressions on shared variables. In addition to using it to compile away any unnecessary loads from variables in shared memory, we also use it to ensure that an expression calculated by RISC in registers is equal to the value of the expression as if it had all been calculated by While in one step. This is especially important when writing the result of an expression back to shared memory, because the refinement is required to maintain all shared memory values.

New to the wr-compiler is the responsibility of maintaining an assumption record, which it uses primarily to detect and reject programs with data races on shared memory, and to rule out the introduction of any new ones. Each assumption record  $S::(Var\ set \times Var\ set)$  is a pair tracking the set of variables on which (resp.) **AsmNoW**, **AsmNoRW** assumptions are currently active at a given point in the program being compiled. As a secondary concern we also use it to assert that the two sides of any if-conditional branches act consistently on the mode state, and that while-loops restore the original mode state on termination.

A compilation record  $C = (\Phi, \mathcal{S}) :: CompRec$  is then just a register/assumption record pair. For readability, we use regrec, asmrec to denote (resp.) a CompRec's fst, snd projections.

To explain how the compilation record is used to rule out data races, and to ensure consistency of expression evaluation between source and target program, firstly we must introduce the concept of *stability* of a variable v according to an assumption record S:

var-stable 
$$S v \equiv v \in (\text{fst } S \cup \text{snd } S) \land (\forall v' \in C \text{vars } v. \ v' \in (\text{fst } S \cup \text{snd } S))$$

In short, this means that the variable and all its control variables (Cvars v) are recorded as having either of AsmNoW or AsmNoW active on them.

For register record entries to be of any help in ensuring consistency of While and RISC expression evaluation, we exclude expression evaluation on data race-prone variables by lifting the concept of stability to register records. The following predicate asserts internal consistency of the compilation record C created by compile-cmd, in the sense that the register record may only map to expressions that mention variables that are recorded as stable by the assumption record accompanying it. (Here, ran denotes the range of a map.)

```
regrec-stable C \equiv \forall e \in \text{ran (regrec } C). (\forall v \in \text{exp-vars } e. \text{ var-stable (asmrec } C) \ v)
```

To ensure that an input While program maintains register record stability, we define the predicate no-unstable-exprs c C to capture the requirement that a program c, if started with a configuration consistent with compilation record C, will never access a lock-protected variable without holding the relevant lock. (It also checks the secondary, mode-state consistency concerns of the assumption record mentioned earlier.) We implement it as a simple static check carried out by a primitive recursive function on the structure of While programs.

Together, regrec-stable and no-unstable-exprs make up the main two requirements of a predicate compile-cmd-input-reqs  $C\ l\ nl\ c$  imposed on the input arguments to compile-cmd, which gives us enough information to prove a lemma that compile-cmd only ever outputs stable register records. Full details of these we leave to Appendix D.3.

#### 5.5 Proof of CVDNI-preserving compilation

Having covered the most significant aspects of the COVERN wr-compiler's parameters and machinery, we can now present the refinement relation  $\mathcal{R}_{wr}$  (Section 5.5.1), pacing function abs-steps<sub>wr</sub> (Section 5.5.2), and coupling invariant  $\mathcal{I}_{wr}$  (Section 5.5.3) that we use with our new decomposition principle (of Section 3) to prove that it preserves CVDNI (Section 5.5.4).

#### 5.5.1 Refinement relation $\mathcal{R}_{wr}$ and its invariants

Just like our example  $\mathcal{R}$  of Figure 3,  $\mathcal{R}_{wr}$  pairs abstract with concrete configurations.

Here, we will focus on  $\mathcal{R}_{wr}$ 's most notable characteristics for understanding why it is suitable to describe a CVDNI-preserving compilation.<sup>3</sup> We focus on the case if\_expr of  $\mathcal{R}_{wr}$ , which relates the expression evaluation part of the While program if e then  $c_1$  else  $c_2$  fi, with the corresponding part (including the conditional jump Jz after expression evaluation) of the RISC program obtained by running compile-cmd on it. (Variables ignored are in gray.)

#### **Example 10** (Introduction rule for case if\_expr of $\mathcal{R}_{wr}$ ).

```
c = \text{if } e \text{ then } c_1 \text{ else } c_2 \text{ fi} \qquad \text{compile-cmd-input-reqs } C \ l \ nl \ c \\ (PCs, l', nl_2, C', \mathsf{False}) = \text{compile-cmd } C \ l \ nl \ c \\ (P_e, r, C_1, \mathsf{False}) = \text{compile-expr } C \varnothing \ l \ e \\ (P_1, l_1, nl_1, C_2, \mathsf{False}) = \text{compile-cmd } C_1 \text{ None } (\mathsf{Suc } nl)) \ c_1 \\ pc \leq \mathsf{length } P_e \\ (P_2, l_2, nl_2, C_3, \mathsf{False}) = \mathsf{compile-cmd } C_1 \text{ (Some } nl) \ nl_1 \ c_2 \\ \mathsf{compiled-cmd-config-consistent } C_pc = (\mathsf{map } \mathsf{snd } PCs \ ! \ pc) \\ \mathsf{compiled-cmd-config-consistent } C_pc \ regs \ mds \ mem \\ \mathsf{regrec-stable } C_pc \\ \forall mds' \ mem' \ regs'. \ \mathsf{compiled-cmd-config-consistent } C_1 \ regs' \ mds' \ mem' \land \mathsf{regrec-stable } C_1 \\ \longrightarrow ((\langle c_1, mds', mem' \rangle_{\mathsf{w}}, \langle ((0, \mathsf{map } \mathsf{fst } P_1), regs'), mds', mem' \rangle_{\mathsf{r}}) \in \mathcal{R}_{\mathsf{wr}} \land \\ (\langle c_2, mds', mem' \rangle_{\mathsf{w}}, \langle ((0, \mathsf{map } \mathsf{fst } P2), regs'), mds', mem' \rangle_{\mathsf{r}}) \in \mathcal{R}_{\mathsf{wr}} \\ (\langle c_1, mds, mem \rangle_{\mathsf{w}}, \langle ((pc, \mathsf{map } \mathsf{fst } PCs), regs), mds, mem \rangle_{\mathsf{r}}) \in \mathcal{R}_{\mathsf{wr}} \\ \end{pmatrix}
```

<sup>&</sup>lt;sup>3</sup> We provide an informal description of all of the cases, their purpose, and the invariants they maintain, along with a code listing from compile-cmd relevant to the part that will be presented, in Appendices A and B (respectively). For full details, we refer the reader to the Isabelle formalisation.

This is a fairly typical case of  $\mathcal{R}_{wr}$  in a number of respects:

Firstly, there is a direct reference to the call to compile-cmd for the given While program. Secondly, various guards (compiled-cmd-config-consistent introduced below, and regrec-stable defined in Section 5.4) are asserted in order to restrict the scope of  $\mathcal{R}_{wr}$  only to consider wellformed local program configurations that line up with the conditions captured by the compilation record. Thirdly, the inductive references to  $\mathcal{R}_{wr}$  for  $P_1$  and  $P_2$ , the branches of the conditional that have not been reached yet, are quantified over all configurations that obey the guards compiled-cmd-config-consistent and regrec-stable relative to  $C_1$ , the initial compilation record for each of the sub-calls to compile-cmd for those sub-programs.

The guard compiled-cmd-config-consistent mentioned above asserts that the compilation record C is consistent with the registers regs, memory mem and mode state mds.

```
compiled-cmd-config-consistent C regs mds mem \equiv (\forall r \ e. \ (regrec \ C) \ r = Some \ e \longrightarrow regs \ r = ev_{exp} \ mem \ e) \land asmrec \ C = (mds \ AsmNoW, \ mds \ AsmNoRW)
```

Firstly, for all entries in register record mapping some register r to some expression e, the value held in r of the register bank regs must match the value of e if evaluated under memory mem. Secondly, the assumption record must consist exactly of the program variables the mode state mds says have  $\mathbf{AsmNoW}$ ,  $\mathbf{AsmNoRW}$  on them respectively.

As we will see in Theorem 17, compiled-cmd-config-consistent also serves as *initial config*uration requirements for compiled programs: only configurations obeying them may be used to initialise a RISC program compiled by the wr-compiler with initial compilation record C.

With  $\mathcal{R}_{wr}$  specified, we then prove the two requirements for secure-refinement-decomp that pertain to  $\mathcal{R}_{wr}$  alone: preserves-modes-mem (Definition 4) and closed-others (Definition 5).

- **Lemma 11** ( $\mathcal{R}_{wr}$  preserves modes and memory). preserves-modes-mem  $\mathcal{R}_{wr}$
- ▶ Lemma 12 ( $\mathcal{R}_{wr}$  is closed under changes by others). closed-others  $\mathcal{R}_{wr}$

# 5.5.2 Refinement pacing function abs-steps<sub>wr</sub>

We now nominate an *abs-steps* function, determining the pace at which While programs progress in comparison to the RISC programs that they are compiled to by the wr-compiler.

To assist here and elsewhere, we define a primitive recursive helper leftmost-cmd that given a sequence of ;-separated While commands, strips all but the first: given  $c_1$ ;  $c_2$  it returns leftmost-cmd  $c_1$ , and given any other While program  $c_1$  it returns  $c_2$ .

Our pacing function abs-steps<sub>wr</sub> primarily looks at the form of the RISC program instruction about to be executed. The RISC instructions are divided into three categories:

- Instructions output by compile-expr: Load, Op, and MoveK. For these, abs-steps<sub>wr</sub> returns 1 if the leftmost-cmd of the While program is while e do c od, to allow it to step to if e then (c; while e do c od) else stop fi concurrently with the first RISC step of the compiled expression itself. Otherwise, abs-steps<sub>wr</sub> returns 0 to indicate the While program standing still while the RISC program takes new steps to evaluate the expression.
- "Epilogue" steps: **Jmp** and **Nop** when used for control flow at the end of a smaller compiled program in the context of a larger one. For these, abs-steps<sub>wr</sub> returns 0.
- All other RISC instructions are assumed to proceed at a lockstep pace with the While command they were compiled from, and for these abs-steps<sub>wr</sub> returns 1.

Having nominated abs-steps<sub>wr</sub> and  $\mathcal{R}_{wr}$ , we now have the parameters over which we are obliged to prove refinement preservation (Figure 4a) as demanded by secure-refinement-decomp

(Definition 7). To this end, we prove firstly (elided to Appendix D.3) that every step of execution of a RISC program produced by the wr-compiler from a While program, maintains the consistency demanded by compiled-cmd-config-consistent between configurations and compilation records. Also, we must prove a correctness lemma for the expression compiler:

▶ **Lemma 13.**  $(PCs, r, C', \mathsf{False}) = \mathsf{compile-expr}\ C\ A\ l\ e \implies (\mathsf{regrec}\ C')\ r = \mathsf{Some}\ e$ 

Armed with these facts, we can now prove the main refinement preservation result:

▶ **Lemma 14** ( $\mathcal{R}_{wr}$  is a refinement paced by abs-steps<sub>wr</sub>).

$$\forall lc_w \ lc_r. \ (lc_w, lc_r) \in \mathcal{R}_{\mathsf{Wr}} \longrightarrow \ (\forall lc'_r. \ lc_r \leadsto_{\mathsf{r}} lc'_r \longrightarrow (\exists lc'_w. \ lc_w \leadsto_{\mathsf{w}}^{(\mathsf{abs-steps}_{\mathsf{wr}} \ lc_w \ lc_r)} \ lc'_w \ \land \ (lc'_w, lc'_r) \in \mathcal{R}_{\mathsf{Wr}}))$$

# 5.5.3 Concrete coupling invariant $\mathcal{I}_{wr}$

The next element needed is the concrete coupling invariant  $\mathcal{I}_{wr}$ , which we define as follows:  $\mathcal{I}_{wr} \equiv \{(\langle ((pc,P), regs), mds, mem \rangle_r, \langle ((pc',P'), regs'), mds', mem' \rangle_r) \mid (pc,P) = (pc',P') \}$  In other words,  $\mathcal{I}_{wr}$  asserts that we only need compare local configurations that are at the same location pc = pc' of the same RISC program P = P'. When used in concert with a no-high-branching  $\mathcal{B}$  (see Section 5.5.4), the effect of  $\mathcal{I}_{wr}$  is to ensure that the wr-compiler has not introduced any new branching on sensitive values.

# 5.5.4 Successful compilations are CVDNI-preserving refinements

We are ready to prove preservation. First we qualify that we allow only strong-low-bisim-mm  $\mathcal{B}$  that describe only While-programs with no branching on High-classified values, as follows:

```
no-high-branching \mathcal{B}\equiv
```

$$\forall c\ c'\ mds\ mem\ mem'.\ (\langle c, mds, mem \rangle_{\mathsf{w}}, \langle c', mds, mem' \rangle_{\mathsf{w}}) \in \mathcal{B} \longrightarrow c = c' \land (\forall e\ c_1\ c_2.\ \text{leftmost-cmd}\ c = \mathbf{if}\ e\ \mathbf{then}\ c_1\ \mathbf{else}\ c_2\ \mathbf{fi} \longrightarrow \mathsf{ev}_{\mathsf{exp}}\ mem\ e = \mathsf{ev}_{\mathsf{exp}}\ mem'\ e)$$

That is, it refuses to relate configurations at different program locations. Furthermore if it is at a conditional branching point, the expression e determining which branch will be taken evaluates to the same boolean value for both configurations' memories. When imposed on a relation that already ensures Low-equivalent memory modulo modes, this effectively disallows any present or past branching on sensitive values. Then, for such programs:

► Lemma 15. 
$$\frac{\text{strong-low-bisim-mm }\mathcal{B}}{\text{secure-refinement-decomp }\mathcal{B} \ \mathcal{R}_{wr} \ \mathcal{I}_{wr} \ \text{abs-steps}_{wr}}$$

From this it follows immediately via Theorem 9 that  $\mathcal{R}_{wr}$  with the help of  $\mathcal{I}_{wr}$  describes a CVDNI-preserving refinement for non-High-branching While programs:

▶ Corollary 16 ( $\mathcal{R}_{wr}$  is a CVDNI-preserving refinement for non-High-branching programs).

```
strong-low-bisim-mm \mathcal{B} \wedge \mathsf{no}-high-branching \mathcal{B} \implies \mathsf{secure}-refinement \mathcal{B} \mathcal{R}_\mathsf{wr} \mathcal{I}_\mathsf{wr}
```

Finally, we prove that successful compilation produces a RISC program related by  $\mathcal{R}_{wr}$  to its input While program, when started with corresponding and reasonable initial configurations:

▶ **Theorem 17** (Successful compilations are refinements in  $\mathcal{R}_{wr}$ ).

$$(PCs, l', nl', C', failed) = {\sf compile-cmd} \ C \ l \ nl \ c$$
 compile-cmd-input-reqs  $C \ l \ nl \ c$  failed = False compiled-cmd-config-consistent  $C \ regs \ mds \ mem$   $P = {\sf map} \ {\sf fst} \ PCs$ 

$$(\langle c, mds, mem \rangle_{\mathsf{w}}, \langle ((0, P), regs), mds, mem \rangle_{\mathsf{r}}) \in \mathcal{R}_{\mathsf{wr}}$$

# 6 Case study: the wr-compiler in action

To test the theory, we instantiated it and applied the wr-compiler to a While-language model of the Cross Domain Desktop Compositor [5] (CDDC), a non-trivial concurrent program that facilitates a trusted user's interaction with multiple desktop machines of differing clearance.

The CDDC model to which we applied the compiler is a 2-thread program that was a precursor to the 3-thread model that was verified using the COVERN program logic [20].<sup>4</sup> Each of the threads of the CDDC program (together about 150 lines of While) we proved satisfy the compositional security property com-secure (Definition 2), using a precursor to the COVERN logic that yields CVDNI-witness bisimulations that are non-High-branching.

The resulting compiler is *executable* in Isabelle, meaning that compile-cmd can be executed on the While program text for each of the two threads to obtain their compilations (together totalling about 250 RISC instructions) using the Isabelle tactic eval. The secure compilation theorems (Section 5.5.4), together with strong-low-bisim-mm preservation and compositionality for com-secure (Theorems 5.1, 3.1 of [22], mentioned in Section 2) then allow us to derive that the compiled program is secure when its threads are run concurrently.

To our knowledge this is the first proof of source-level information-flow security being carried by a verified compiler to an assembly-level model of a non-trivial concurrent program.

# 7 Related work

The following three works, like ours, focus on compilation preserving a form of noninterference. Tedesco et al. [29] present a type-directed compilation scheme that preserves a fault-resilient noninterference property. The compilation scheme of our wr-compiler was inspired by theirs. Like our com-secure CVDNI security property that wr-compiler preserves, Tedesco et al.'s security property is also strong bisimulation-based [27]. But where our property accounts (via mode states) for controlled interference by other threads, theirs instead quantifies over all possible interference by the environment with the memory contents. While this simplifies their task of proving that their security property is preserved under compilation—as it need not require the compiler to preserve the contents of memory—it means their security property cannot capture value-dependent noninterference. In contrast, our wr-compiler must obey our secure-refinement notion's requirement that memory contents are preserved.<sup>5</sup>

Barthe et al. [2] consider the problem of preserving cryptographic constant-time policies, a class of noninterference properties similar to CVDNI in its explicit consideration for capturing timing-sensitivity. Barthe et al. consider a wider scope of common categories of compile-time optimisations (than those performed by our wr-compiler), and mechanise proofs in Coq that such optimisations preserve various constant-time security properties. The sharing of variables in our setting severely limits the scope of our optimisations, to those that the compiler can perform knowing that a shared variable is stable because it has been locked. At present, our wr-compiler avoids redundant loads during expression compilation, but other optimisations like loop hoisting and constant folding we are yet to implement. Their preservation proof technique, constant-time simulation was developed independently to our original cube-shaped secure refinement definition [22]. Like ours, theirs is also a cube-shaped obligation and makes

<sup>&</sup>lt;sup>4</sup> We leave for future work an adaptation of the refinement theory and wr-compiler in order to support the *shared data invariants* added by the COVERN logic, required to verify the 3-thread CDDC model.

<sup>&</sup>lt;sup>5</sup> Consequently, we found and fixed a bug in their expression compiler (acknowledged privately) whereby registers in use were incorrectly reallocated. Expressions like v + (v + 1) were thus compiled incorrectly to programs yielding (v + 1) + (v + 1) instead, causing a violation of memory contents preservation.

use of a pacing function analogous to our *abs-steps*. Unlike our work here, Barthe et al. do not give a general method for decomposing their cube-shaped simulation diagrams.

Neither of the above consider per-thread compositional compilation of concurrent, shared memory programs, nor value-dependent noninterference policies – the focus of our theory and compiler. Barthe et al. [4] however did aim to preserve noninterference of multithreaded programs by compilation, extending a prior (security) type-preserving compilation approach [3]. Their noninterference property however was termination- and timing-insensitive, so preventing internal timing leaks relied on the scheduler disallowing certain interleavings between threads. Also, their type-preservation argument was derived from a big-step semantics preservation property for their compiler. Here we instead rely on preservation of a small-step semantics (specifically memory contents), which is necessary for us to preserve value-dependent security under compilation, as well as to avoid imposing non-standard requirements on the scheduler.

Other recent works have improved on fully abstract compilation (surveyed [23]) by mapping out the spectrum [1] or developing specific forms [25] of robust property preservation, concerned with robustness of source program (hyper)properties to concrete adversarial contexts. Like Tedesco et al. [29], these works differ from ours in quantifying over a wider range of hostile interference. They also focus prominently on changes to data types, which we do not support. Thus, as a 2-safety hyperproperty quantifying over a lesser range of interference, we expect CVDNI-preservation to be implied by R2HSP (robust 2-hypersafety preservation), but do not expect it to imply any other secure compilation criterion on Abate et al.'s [1] spectrum.

While recently Patrignani and Garg [25] instantiated their robustly safe compilation for shared-memory fork-join concurrent programs, it only preserves (1-)safety properties. Previously however, Patrignani et al. [24] proved their trace-preserving compilation preserves k-safety hyperproperties [6], including noninterference properties. However, it disallows the removal or addition of trace entries, which would be necessary to change the passage of time as seen in the observable trace events. Thus it excludes optimisations carried out by our compiler (when it permits changes to pacing regulated by abs-steps) and studied by the two other works [29, 2] on timing-sensitive security-preserving compilation mentioned above.

Finally, there has been much work on large-scale verified compilation [15, 14] some of which has also treated compilation of shared-memory concurrent programs [17] including taking weak-memory consistency into account [26]. Our work here does not consider the effects of weak-memory models. However, it differs to prior work on verified concurrent compilation, in that it formalises and proves a compiler's ability to use information about the application's locking protocol, to exclude unsafe access to shared variables, and conversely to know when it is safe to allow optimisations that would typically be excluded (see Section 5.4).

#### 8 Conclusion

To our knowledge, we have presented the first mechanised verification that a compiler preserves concurrent, value-dependent noninterference. To this end, we provided a general decomposition principle for compositional, secure refinement. Although our compiler is a proof-of-concept targeting simple source and target languages, we nevertheless applied it to produce a verified assembly-level model of the CDDC [5], a non-trivial concurrent program.

This work serves to demonstrate that verified security-preserving compilation for concurrent programs is now within reach, by augmenting traditional proof obligations for verified compilation (e.g. square-shaped semantics preservation) with those specific to security (e.g. absence of termination- and timing-leaks) as depicted in Figure 4. We hope that this work paves the way for future large-scale verified security-preserving compilation efforts.

#### References

- 1 Carmine Abate, Roberto Blanco, Deepak Garg, Catalin Hritcu, Marco Patrignani, and Jérémy Thibault. Exploring robust property preservation for secure compilation. *CoRR*, abs/1807.04603, 2018. URL: http://arxiv.org/abs/1807.04603.
- 2 G. Barthe, B. Grégoire, and V. Laporte. Secure compilation of side-channel countermeasures: The case of cryptographic "constant-time". In 2018 IEEE 31st Computer Security Foundations Symposium (CSF), pages 328–343, July 2018.
- 3 Gilles Barthe, Tamara Rezk, and Amitabh Basu. Security types preserving compilation. Comput. Lang. Syst. Struct., 33(2):35-59, July 2007. URL: http://dx.doi.org/10.1016/j.cl.2005.05.002.
- 4 Gilles Barthe, Tamara Rezk, Alejandro Russo, and Andrei Sabelfeld. Security of multithreaded programs by compilation. *ACM Trans. Inf. Syst. Secur.*, 13(3):21:1–21:32, July 2010. URL: http://doi.acm.org/10.1145/1805974.1805977.
- 5 Mark Beaumont, Jim McCarthy, and Toby Murray. The cross domain desktop compositor: Using hardware-based video compositing for a multi-level secure user interface. In *Annual Computer Security Applications Conference (ACSAC)*, pages 533–545, 2016.
- 6 Michael R. Clarkson and Fred B. Schneider. Hyperproperties. J. Comput. Secur., 18(6):1157–1210, September 2010. URL: http://dl.acm.org/citation.cfm?id=1891823.1891830.
- 7 Florian Dewald, Heiko Mantel, and Alexandra Weber. AVR processors as a platform for language-based security. In Computer Security ESORICS 2017 22nd European Symposium on Research in Computer Security, Oslo, Norway, September 11-15, 2017, Proceedings, Part I, pages 427-445, 2017. URL: https://doi.org/10.1007/978-3-319-66402-6\_25.
- **8** Qian Ge, Yuval Yarom, Tom Chothia, and Gernot Heiser. Time protection: the missing OS abstraction. In *Eurosys19*, Dresden, Germany, March 2019. ACM.
- 9 Qian Ge, Yuval Yarom, and Gernot Heiser. No security without time protection: We need a new hardware-software contract. In Asia-Pacific Workshop on Systems (APSys), Korea, August 2018. ACM SIGOPS.
- Joseph Goguen and José Meseguer. Security policies and security models. In Proceedings of the IEEE Symposium on Security and Privacy, pages 11–20, Oakland, California, USA, April 1982. IEEE Computer Society.
- 11 Cliff B. Jones. Development Methods for Computer Programs including a Notion of Interference. D.Phil. thesis, University of Oxford, June 1981.
- Thierry Kaufmann, Hervé Pelletier, Serge Vaudenay, and Karine Villegas. When constant-time source yields variable-time binary: Exploiting curve25519-donna built with msvc 2015. In Cryptology and Network Security, pages 573–582, Cham, 2016. Springer International Publishing.
- Paul Kocher, Jann Horn, Anders Fogh, , Daniel Genkin, Daniel Gruss, Werner Haas, Mike Hamburg, Moritz Lipp, Stefan Mangard, Thomas Prescher, Michael Schwarz, and Yuval Yarom. Spectre attacks: Exploiting speculative execution. In 40th IEEE Symposium on Security and Privacy (S&P'19), 2019.
- 14 Ramana Kumar, Magnus Myreen, Michael Norrish, and Scott Owens. CakeML: A verified implementation of ML. In ACM SIGPLAN-SIGACT Symposium on Principles of Programming Languages, pages 179–191, San Diego, January 2014. ACM Press.
- Xavier Leroy. A formally verified compiler back-end. J. Autom. Reason., 43(4):363-446, December 2009. URL: http://dx.doi.org/10.1007/s10817-009-9155-4.
- Moritz Lipp, Michael Schwarz, Daniel Gruss, Thomas Prescher, Werner Haas, Anders Fogh, Jann Horn, Stefan Mangard, Paul Kocher, Daniel Genkin, Yuval Yarom, and Mike Hamburg. Meltdown: Reading kernel memory from user space. In 27th USENIX Security Symposium (USENIX Security 18), 2018.
- Andreas Lochbihler. Mechanising a type-safe model of multithreaded java with a verified compiler. *Journal of Automated Reasoning*, 61(1):243–332, Jun 2018. URL: https://doi.org/10.1007/s10817-018-9452-x.

- 18 Luísa Lourenço and Luís Caires. Dependent information flow types. In ACM SIGPLAN-SIGACT Symposium on Principles of Programming Languages, pages 317–328, Mumbai, India, January 2015. ACM.
- 19 Heiko Mantel, David Sands, and Henning Sudbrock. Assumptions and guarantees for compositional noninterference. In *IEEE Computer Security Foundations Symposium*, pages 218–232, Cernay-la-Ville, France, June 2011. IEEE.
- 20 Toby Murray, Robert Sison, and Kai Engelhardt. COVERN: A logic for compositional verification of information flow control. In European Symposium on Security and Privacy, London, United Kingdom, April 2018. IEEE.
- 21 Toby Murray, Robert Sison, Edward Pierzchalski, and Christine Rizkallah. Compositional security-preserving refinement for concurrent imperative programs. Archive of Formal Proofs, June 2016. http://isa-afp.org/entries/Dependent\_SIFUM\_Refinement.shtml, Formal proof development.
- 22 Toby Murray, Robert Sison, Edward Pierzchalski, and Christine Rizkallah. Compositional verification and refinement of concurrent value-dependent noninterference. In *IEEE Computer Security Foundations Symposium*, pages 417–431, Lisbon, Portugal, June 2016.
- Marco Patrignani, Amal Ahmed, and Dave Clarke. Formal approaches to secure compilation: A survey of fully abstract compilation and related work. *ACM Comput. Surv.*, 51(6):125:1–125:36, February 2019. URL: http://doi.acm.org/10.1145/3280984.
- 24 Marco Patrignani and Deepak Garg. Secure Compilation and Hyperproperty Preservation. In IEEE 30th Computer Security Foundations Symposium, CSF 2017, Santa Barbara, USA, August 21 - 25, 2017, CSF'17, 2017.
- 25 Marco Patrignani and Deepak Garg. Robustly safe compilation. In *Programming Languages* and Systems, pages 469–498, Cham, 2019. Springer International Publishing.
- Anton Podkopaev, Ori Lahav, and Viktor Vafeiadis. Bridging the gap between programming languages and hardware weak memory models. *Proc. ACM Program. Lang.*, 3(POPL):69:1–69:31, January 2019. URL: http://doi.acm.org/10.1145/3290382.
- Andrei Sabelfeld and David Sands. Probabilistic noninterference for multi-threaded programs. In *Proceedings of the 13th IEEE Workshop on Computer Security Foundations*, CSFW '00, pages 200-, Washington, DC, USA, 2000. IEEE Computer Society. URL: http://dl.acm.org/citation.cfm?id=794200.795151.
- 28 Mark Staples, Ross Jeffery, June Andronick, Toby Murray, Gerwin Klein, and Rafal Kolanski. Productivity for proof engineering. In *Empirical Software Engineering and Measurement*, page 15, Turin, Italy, September 2014.
- 29 F. Del Tedesco, D. Sands, and A. Russo. Fault-resilient non-interference. In 2016 IEEE 29th Computer Security Foundations Symposium (CSF), pages 401–416, June 2016.
- Danfeng Zhang, Yao Wang, G. Edward Suh, and Andrew C. Myers. A hardware design language for timing-sensitive information-flow security. In Proceedings of the Twentieth International Conference on Architectural Support for Programming Languages and Operating Systems, ASPLOS '15, pages 503-516, New York, NY, USA, 2015. ACM. URL: http://doi.acm.org/10.1145/2694344.2694372.

# **A** Informal descriptions of the cases of refinement relation $\mathcal{R}_{\mathsf{wr}}$

#### A.1 Base cases

- **stop**: This case relates a terminated While program with a terminated RISC program (i.e. one where the program counter is at the length of the program text).
- skip\_nop: This case relates the While program skip with the configuration where the program counter is at the start of the RISC program [Nop].

- **assign\_expr**: This case relates the expression evaluation part (for the expression e) of the While program v := e with the corresponding part of the RISC program obtained by compiling it with the wr-compiler.
- **assign\_store**: As for assign\_expr, but for the very last **Store** instruction that commits the result of the expression evaluation back to shared memory variable v.
  - It asserts additionally that v must be stable if lock-governed, and non-lock-governed otherwise. This prevents threads from violating the locking discipline (see Section 5.3).
- **lock\_acg**: This case relates lock(k) with LockAcg k.
- $\blacksquare$  lock\_rel: This case relates unlock(k) with LockRel k.

#### A.2 Inductive cases

■ seq: This case relates the While program  $c_1$ ;  $c_2$  with the concatenation  $P_1@P_2$  of the RISC programs  $P_1$  and  $P_2$  that are respectively the outputs of successful consecutive compilation of  $c_1$  and  $c_2$  by the wr-compiler. It is intended for cases where the While (resp. RISC) program is currently in  $c_1$  (resp.  $P_1$ ).

It is an inductive case of  $\mathcal{R}_{wr}$ , in that:

- $c_1$  is required to be related by  $\mathcal{R}_{wr}$  to the present location in  $P_1$ .
- For all local configurations that obey the compiled-cmd-config-consistent requirements,  $c_2$  is required to be related by  $\mathcal{R}_{wr}$  to the first instruction of  $P_2$ . This quantification ensures that  $\mathcal{R}_{wr}$  remains closed when execution progresses from the first program to the second program.

It asserts that  $P_1$  and  $P_2$  are joinable (Section D.2), particularly relevant here to ensure that  $P_1$  can only jump to locations within or at the end of itself (i.e. the start of  $P_2$ ).

- join: This case relates a While program c with an offset  $pc > \text{length } P_1$  into a RISC program  $P_1@P_2$ , assuming the inductive hypothesis that c is related by  $\mathcal{R}_{wr}$  with the offset  $pc \text{length } P_1$  into the RISC program  $P_2$  alone.
  - It is intended primarily for cases where the While (resp. RISC) program is currently in the  $c_2$  (resp.  $P_2$ ) of some consecutively compiled  $c_1$ ;  $c_2$  (resp.  $P_1$  concatenated with  $P_2$ ) but applies more broadly to allow any prepend of dead, unreachable instructions onto the front of a RISC program without breaking  $\mathcal{R}_{wr}$ .
  - It also asserts that  $P_1$  and  $P_2$  are joinable, which is important here to ensure that  $P_2$  cannot jump back into  $P_1$ .
- if\_expr: This case relates the expression evaluation part (for the expression e) of the While program if e then  $c_1$  else  $c_2$  fi with the corresponding part (including the conditional jump Jz at the end of expression evaluation) of the RISC program obtained by compiling it with the wr-compiler.
  - It relies on both  $c_1$  and  $c_2$  being related by  $\mathcal{R}_{wr}$  to its compiled RISC counterparts when started with initialisation states judged valid by compiled-cmd-config-consistent.
- if\_c1: This case relates some While program  $c'_1$  reachable from  $c_1$  with the corresponding part within the  $c_1$  part of the RISC program obtained by compiling if e then  $c_1$  else  $c_2$  fi with the wr-compiler.
  - It relies on  $c_1$  being related by  $\mathcal{R}_{wr}$  to its compiled RISC counterpart at the appropriate program counter offset.
- **=** if\_c2: As for if\_c1, but for  $c_2$ .
- epilogue\_step: This case relates a terminated While program to the silent control flow steps navigating to the end of a RISC program from the end of the "then" and "else" branches of a compiled if-conditional.

It works only for epilogue step forms (see Section 5.5.2).

- It is inductive in that it asserts closedness of  $\mathcal{R}_{wr}$  over pairwise reachability from the pair currently under consideration the only case to do so directly.
- while\_expr: This case relates the While program (while e do c od)'s initial intermediate step to if e then (e; while e do e od) else stop fi, and its expression evaluation part, with the expression evaluation and conditional jump of the RISC program that while e do e od was compiled to by compile-cmd.
  - It relies on c being related by  $\mathcal{R}_{wr}$  to its compiled RISC counterpart when started with initialisation states judged valid by compiled-cmd-config-consistent.
- while\_inner: This case relates some program  $c_I$ ; while e do c od reachable from c; while e do c od to the loop body part of the RISC program compiled from while e do c od. It relies on  $c_I$  being related by  $\mathcal{R}_{wr}$  to its compiled RISC counterpart at the appropriate program counter offset.
  - It also carries around the same reliance on c being related by  $\mathcal{R}_{\mathsf{wr}}$  to its compiled RISC counterpart for all initialisation states judged valid by compiled-cmd-config-consistent.
- while\_loop: This case handles epilogue steps for the inner loop body program, and the final jump back to the beginning of the While-loop.
  - It requires  $\mathcal{R}_{wr}$  to relate the terminated While program to the end of the compiled loop body, and furthermore also carries around the same reliance on c being related by  $\mathcal{R}_{wr}$  to its compiled RISC counterpart for all initialisation states judged valid by compiled-cmd-config-consistent.

# B Code listing for the case of compile-cmd for if-conditionals

This code listing has been adapted slightly to improve the clarity of the presentation.  $\Phi \sqcap_R \Phi'$  denotes the subset of mappings on which  $\Phi$  and  $\Phi'$  agree.

#### **Listing 1** Implementation of compile-cmd case for if e then $c_1$ else $c_2$ fi

```
compile_cmd C l nl (If e c<sub>1</sub> c<sub>2</sub>) = 
 (let (P<sub>e</sub>, r, C<sub>1</sub>, fail<sub>e</sub>) = (compile_expr C {} l e);
  (br, nl') = (nl, Suc nl); (ex, nl'') = (nl', Suc nl');
  (P<sub>1</sub>, l<sub>1</sub>, nl<sub>1</sub>, C<sub>2</sub>, fail<sub>1</sub>) = (compile_cmd C<sub>1</sub> None nl'' c<sub>1</sub>);
  (P<sub>2</sub>, l<sub>2</sub>, nl<sub>2</sub>, C<sub>3</sub>, fail<sub>2</sub>) = (compile_cmd C<sub>1</sub> (Some br) nl<sub>1</sub> c<sub>2</sub>);
  (* Pre-compilation check ensures asmrec C<sub>2</sub> = asmrec C<sub>3</sub> *)
  C' = (regrec C<sub>2</sub> \sqcap_R regrec C<sub>3</sub>, asmrec C<sub>2</sub>)
  in (P<sub>e</sub> @ [((if P<sub>e</sub> = [] then l else None, Jz br r), C<sub>1</sub>)] @
   P<sub>1</sub> @ [((l<sub>1</sub>, Jmp ex), C<sub>2</sub>)] @ P<sub>2</sub> @ [((l<sub>2</sub>, Nop'), C<sub>3</sub>)],
   Some ex, nl<sub>2</sub>, C', fail<sub>e</sub> \vee fail<sub>1</sub> \vee fail<sub>2</sub>))
```

### C Proof sketch for decomposition principle soundness result

▶ **Theorem 18** (Soundness of secure-refinement-decomp).

```
secure-refinement-decomp \mathcal{B} \ \mathcal{R} \ \mathcal{I} \ abs\text{-}steps \implies secure-refinement \mathcal{B} \ \mathcal{R} \ \mathcal{I}
```

**Proof.** The only obligation for secure-refinement (Definition 6) not obtained immediately from secure-refinement-decomp (Definition 7) is the cube-shaped coupling-inv-pres (Figure 1).

The front face of the cube is just ordinary square-shaped refinement preservation (depicted in Figure 4a), given to us by secure-refinement-decomp. This gives us that a single concrete step from  $lc_{1C}$  is simulated by n abstract steps  $lc_{1A}$ , where n is given by abs-steps.

We are then obliged to prove a simulation in the other direction (the back face of the cube), that n abstract steps from all configurations  $lc_{2A}$  related by  $\mathcal{B}$  to  $lc_{1A}$  are simulated by some concrete step from  $lc_{2C}$  related by  $\mathcal{R}$  to  $lc_{2A}$  and by  $\mathcal{I}$  to  $lc_{1C}$ .

Here, we lean on the determinism of the abstract program's evaluation semantics (required by the theory) to flip the direction of simulation, knowing that n abstract steps from  $lc_{2A}$ , simulating a single concrete step from  $lc_{2C}$ , could only be the very same n abstract steps from  $lc_{2A}$  that we were required to consider. This allows us to use once again the square-shaped refinement preservation (Figure 4a) given to us by secure-refinement-decomp.

Consistency of refinement pacing and stopping behaviour (depicted in Figure 4b) given by decomp-refinement-safe (Definition 8) then respectively ensure that n (via abs-steps) is the correct number of abstract steps to consider, and that there will indeed be a concrete step from  $lc_{2C}$  to drive the matching simulation step.

Finally, the remainder of decomp-refinement-safe (depicted in Figure 4c) discharges the requirement of closedness and modes-equality maintenance of  $\mathcal{I}$  under lockstep execution, demanded by the bottom face of the cube.

# D More details on the COVERN wr-compiler

## D.1 Register allocation scheme model

We model the (user-supplied) register allocation scheme by two functions  $reg\_alloc$  and  $reg\_alloc\_cached$  on the register record  $\Phi$  (see Section 5.4) and the set A of registers whose contents are needed to evaluate the current expression. In order to avoid loading from memory unnecessarily, the compiler may first call  $reg\_alloc\_cached$   $\Phi$  A v to identify a register that  $\Phi$  records as already containing the variable v. When the compiler needs a fresh register, it will call  $reg\_alloc$   $\Phi$  A. Neither function is allowed to allocate a register in A, so the allocator is permitted to fail if it cannot find any suitable register. As mentioned in Section 5.2 we assume there are enough registers to service the expressions in the source program. Also, registers typically become available again as expression evaluation is resolved.

#### D.2 Label allocation and sequential composability

For allocating natural numbers to use as labels for RISC instructions the wr-compiler ensures freshness merely by using the highest number reached so far on a "next label" counter (nl in the invocation example (1)), incrementing the counter before passing it along to subsequent calls, and outputting the next available unused label on return (as nl' in the example).

We define two RISC programs  $P_1, P_2$  to be joinable if they are both:

- $\blacksquare$  joinable-forward:  $P_1$  only ever jumps to labels that are either
  - $\blacksquare$  labelling an instruction in  $P_1$  itself, or
  - $\blacksquare$  the label of the very first instruction in  $P_2$ .
- **–** joinable-backward:  $P_2$  does not jump to any of the labels of instructions in  $P_1$ .

We prove a lemma that says that two RISC programs that were compiled by the wr-compiler consecutively—in the sense that the relevant outputs from the first call are fed directly into the second call—are joinable.

#### D.3 More detail on compile-cmd-input-regs and the wr-compiler proofs

The first two requirements to the predicate compile-cmd-input-reqs  $C\ l\ nl\ c$  were given in Section 5.4. Its other two requirements reflect that the terminated While program stop has

no valid compilation, and that the initial label (if provided) must be valid (see Section D.2 for more information on label allocation).

▶ **Definition 19** (Requirements on inputs to compile-cmd).

```
compile-cmd-input-reqs C\ l\ nl\ c \equiv c \neq stop \land (\forall x.\ l = \mathsf{Some}\ x \longrightarrow x < nl) \land \mathsf{no-unstable-exprs}\ c\ C\ \land \mathsf{regrec-stable}\ C
```

These input conditions give us enough information to prove that every instruction of a CompRec-annotated RISC program output by a successful run of compile-cmd is annotated by a stable register record, and that the output CompRec's register record is also stable:

▶ **Lemma 20** (Successful compilations output only stable register records).

```
\frac{(\mathit{PCs}, \mathit{l'}, \mathit{nl'}, \mathit{C'}, \mathsf{False}) = \mathsf{compile-cmd} \ \mathit{C} \ \mathit{l} \ \mathit{nl} \ \mathit{c}}{(\forall \mathit{pc} < \mathsf{length} \ \mathit{PCs}. \ \mathsf{regrec\text{-stable}} \ (\mathsf{map} \ \mathsf{snd} \ \mathit{PCs} \ ! \ \mathit{pc})) \ \land \ \mathsf{regrec\text{-stable}} \ \mathit{C'}}
```

**Proof.** By induction on the structure of the While language program c, making reference to the implementation of compile-cmd. For cases that must compile expressions, we furthermore prove and make use of a lemma by induction on the structure of expressions, making reference to the implementation of the expression compiler function compile-expr called by compile-cmd. In essence, we prove that (sub)expressions that appear in register records must be stable, for two reasons. Firstly, they are always only ever subexpressions over variables that must have been stable in the input program when their contents were first loaded into registers. Furthermore, when compiling an  $\mathbf{unlock}()$ , the  $\mathbf{wr-compiler}$  will always flush all register records that make reference to any variables that the  $\mathbf{unlock}()$  makes unstable.

Before proceeding, we name the parts of compiled-cmd-config-consistent more explicitly:

▶ **Definition 21** (Configuration consistency requirements for compiled commands).

```
compiled-cmd-config-consistent C regs mds mem \equiv regrec-mem-consistent (regrec C) regs mem \land asmrec-mds-consistent (asmrec C) mds
```

▶ Definition 22 (Consistency between a register record, register bank, and shared memory).

```
regrec-mem-consistent \Phi regs mem \equiv \forall r e. \Phi r = Some e \longrightarrow regs r = \operatorname{ev}_{\operatorname{exp}} mem e
```

▶ **Definition 23** (Consistency between an assumption record and a mode state).

```
asmrec-mds-consistent \mathcal{S} mds \equiv \mathcal{S} = (mds \ \mathbf{AsmNoW}, \ mds \ \mathbf{AsmNoRW})
```

▶ **Lemma 24** ( $\mathcal{R}_{wr}$  preserves modes and memory). preserves-modes-mem  $\mathcal{R}_{wr}$ 

**Proof.** By induction on the structure of  $\mathcal{R}_{wr}$ . For all cases of  $(lc_w, lc_r) \in \mathcal{R}_{wr}$ ,  $lc_w =_{mds}^{mem} lc_r$  is either asserted directly by the guards or obtainable from the inductive hypothesis.

**Lemma 25** ( $\mathcal{R}_{wr}$  is closed under changes by others). closed-others  $\mathcal{R}_{wr}$ 

**Proof.** By induction on the structure of  $\mathcal{R}_{wr}$ . Changes by others (Definition 5) only modify writable variables the same way for both configurations, so preservation of  $=^{mem}_{mds}$  is immediate. Also, regrec-mem-consistent is unaffected because compile-cmd only creates regrec-stable records (referring to no writable variables). No other  $\mathcal{R}_{wr}$  guards mention shared memory.

▶ Lemma 26 (Successfully compiled programs maintain config consistency requirements).

```
(PCs, l', nl', C', failed) = \text{compile-cmd} \ C \ l \ nl \ c compile-cmd-input-reqs C \ l \ nl \ c failed = False pc < \text{length} \ PCs P = \text{map} \ \text{fst} \ PCs Cs = \text{map} \ \text{snd} \ PCs compiled-cmd-config-consistent} (Cs \ l \ pc) \ regs \ mds \ mem \langle ((pc, P), regs), mds, mem \rangle_r \rightsquigarrow_r \langle ((pc', P), regs'), mds', mem' \rangle_r)
```

compiled-cmd-config-consistent (if  $pc' < \text{length } P \text{ then } (Cs \mid pc') \text{ else } C') \ regs' \ mds' \ mem'$ 

**Proof.** We in fact prove it separately for regrec-mem-consistent and asmrec-mds-consistent, in both cases by induction on the structure of the While program c. In each case, we use the simplifiers for the compile-cmd implementation to yield the corresponding RISC program fragment in question, and then prove the lemma for each of the possible locations of pc in the compiled program. For both proofs, there is some trickiness in accounting for (and ruling out) which destination pc' must be considered for each of these cases of pc, particularly for those While programs that compile to RISC programs that may have jumps in them.

Control flow trickiness aside, the intuition for regrec-mem-consistent is that it tests the correctness of the compilation of expressions, and so for this we must prove a sub-lemma for maintenance of compiled-cmd-config-consistent by induction on the structure of expressions e that are encountered in the While programs if e then  $c_1$  else  $c_2$  fi, while e do e od, e v:= e. Additionally, unlock() flushes register record entries mentioning variables that are to become unstable, and while e do e od conservatively flushes entries to force evaluation of the loop condition expression. This is safe trivially because flushing entries can never make a consistent register record inconsistent. The rest of the cases for e are straightforward because they do not touch the register record.

Then for asmrec-mds-consistent, the substantial part of the proof is as a test of the correctness of the compiler's bookkeeping of assumptions being consistent with the semantics of lock() and unlock(). The other cases for c do not touch the mode state.

▶ Lemma 27 (Correctness of the expression compiler).

$$(PCs, r, C', \mathsf{False}) = \mathsf{compile-expr}\ C\ A\ l\ e \implies (\mathsf{regrec}\ C')\ r = \mathsf{Some}\ e$$

**Proof.** By induction on the structure of expressions e, using the simplification rules for the implementation of compile-expr, and also relying on assumptions of correctness of the register allocation scheme supplied by the instantiator of the theory.

▶ **Lemma 28** ( $\mathcal{R}_{wr}$  is a refinement paced by abs-steps<sub>wr</sub>).

$$\forall lc_w \ lc_r. \ (lc_w, lc_r) \in \mathcal{R}_{\mathsf{wr}} \longrightarrow \ (\forall lc'_r. \ lc_r \leadsto_{\mathsf{r}} lc'_r \longrightarrow \\ (\exists lc'_w. \ lc_w \leadsto_{\mathsf{w}}^{(\mathsf{abs-steps}_{\mathsf{wr}} \ lc_w \ lc_r)} \ lc'_w \ \land \ (lc'_w, lc'_r) \in \mathcal{R}_{\mathsf{wr}}))$$

**Proof.** By induction on the structure of  $\mathcal{R}_{wr}$ .

The base case stop is immediate, because it pertains to a terminated While and RISC program. The base cases that proceed in one step to a terminating program configuration (skip\_nop, assign\_store, lock\_acq, lock\_rel) are fairly straightforward because after dealing with the single step, the resulting obligation can then be handled by the stop case. This leaves the last remaining base case assign\_expr, which proceeds in one step either to itself, or to assign\_store. In all of these cases, we use Lemma 26 to obtain the preservation of the guards demanded by the  $\mathcal{R}_{wr}$  introduction rule for the destination configuration of the step. Particularly, the assign\_store case must make use of regree-mem-consistent and the

correctness of compile-expr (Lemma 27) in order to ensure that once the expression evaluation result is written back to shared memory,  $lc'_w =_{\mathsf{mds}}^{\mathsf{mem}} lc'_r$  holds as demanded by the stop case.

The inductive cases that concern expression evaluation (if\_expr, while\_expr) are much like assign\_expr in that they have the possibility of progressing in one step to themselves. Unlike assign\_expr however, their other possibility is a conditional jump based on the result of that expression. Again we use Lemma 27 to obtain that the result is an accurate calculation of the expression, and this time we prove by the two different cases whether if\_expr ends up in if\_c1 or if\_c2, or if while\_expr ends up in while\_inner or at stop (having jumped to the exit label). In these cases, the guards over which the inductive references to  $\mathcal{R}_{wr}$  have been quantified are versatile enough to discharge themselves (when \*\_expr steps to itself), or to discharge any reachable initial starting state for the nested compiled RISC program, given that Lemma 26 ensures the invariance of these guards.

This just leaves the inductive cases that pertain to configurations inside a nested compiled RISC program (if\_c1, if\_c2, while\_inner), or at the end of one (epilogue\_step, while\_loop). In these cases, the inductive hypotheses obtained from the inductive reference to  $\mathcal{R}_{wr}$  are always enough to satisfy the guards demanded by the possible destination cases. Like in the proof of Lemma 26, the trickiness mostly comes from accounting for all the possible cases of control flow (ruling out spurious destinations) that need to be considered.

▶ Lemma 29. 
$$\frac{\text{strong-low-bisim-mm }\mathcal{B}}{\text{decomp-refinement-safe }\mathcal{B} \ \mathcal{R}_{\text{wr}} \ \mathcal{I}_{\text{wr}} \ \text{abs-steps}_{\text{wr}}}$$

**Proof.** Definition 8 gives us the following obligations.

For consistent stopping behaviour, we prove a lemma that RISC programs stop if and only if their pc is outside the program text P, i.e. pc > length P. Because  $\mathcal{I}_{wr}$  equates pc and P for the two configurations, then clearly both have identical stopping behaviour.

For consistency of change in timing behaviour,  $abs\text{-steps}_{wr}$  depends only on While and RISC program locations, and no-high-branching and  $\mathcal{I}_{wr}$  forces them (resp.) to be equal for the local configurations under consideration.

For closedness of  $\mathcal{I}_{wr}$  under lockstep execution, the only non-straightforward cases to consider are conditional branching, and the locking primitives. For conditional branching, we use no-high-branching for  $\mathcal{B}$  with memory preservation via  $\mathcal{R}_{wr}$  (Lemma 11) to ensure that the conditional branching outcome is the same on both sides.

Finally, as the only operations that touch mode state, the locking primitives are the only non-straightforward cases for mode state equality maintenance under lockstep execution. As all lock memory is classified Low (see Section 5.3), we use strong-low-bisim-mm for  $\mathcal{B}$  with memory preservation via  $\mathcal{R}_{wr}$  to ensure the RISC configurations behave consistently.

▶ Lemma 30. 
$$\frac{\text{strong-low-bisim-mm }\mathcal{B}}{\text{secure-refinement-decomp }\mathcal{B} \; \mathcal{R}_{wr} \; \mathcal{I}_{wr} \; \text{abs-steps}_{wr}}$$

**Proof.** Referring to Definition 7, the obligations pertaining only to  $\mathcal{R}_{wr}$  and abs-steps<sub>wr</sub> are discharged by Lemma 14, Lemma 12, and Lemma 11. Pertaining to  $\mathcal{I}_{wr}$ : clearly  $\mathcal{I}_{wr}$  is symmetric, and furthermore it is cg-consistent (Definition 3) because the actions over which  $\mathcal{I}_{wr}$  must be closed modify only the shared memory, and  $\mathcal{I}_{wr}$  places only restrictions on the program text and current location. The final obligation is discharged by Lemma 29.

▶ **Theorem 31** (Successful compilations are refinements in  $\mathcal{R}_{wr}$ ).

```
\frac{(PCs, l', nl', C', failed) = \mathsf{compile-cmd} \ C \ l \ nl \ c}{(c, mds, mem)_{\mathsf{w}}, \langle ((0, P), regs), mds, mem)_{\mathsf{r}}) \in \mathcal{R}_{\mathsf{wr}}} = \mathsf{compile-cmd} \ C \ l \ nl \ c}
```

**Proof.** By induction on the structure of While. The compiler input and initial configuration conditions we impose allow us to have each of  $\mathbf{skip}$ , cmd; cmd,  $\mathbf{if}$  exp then cmd else cmd fi, while exp do cmd od, v := exp,  $\mathbf{lock}(k)$ , and  $\mathbf{unlock}(k)$  and their compiled output meet the guards of the introduction rules for the cases  $\mathbf{skip}$ ,  $\mathbf{seq}$ ,  $\mathbf{if}_{expr}$ , while\_expr,  $\mathbf{assign}_{expr}$ ,  $\mathbf{lock}_{acq}$ , and  $\mathbf{lock}_{rel}$  of  $\mathcal{R}_{wr}$  that were designed for them respectively.