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Formally Verified Tag-Based Enforcement of Control Flow Integrity

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The purpose of this diploma thesis is to present a novel, hardware-assisted, formally verified implementation of low-level security policies, such as Control-Flow Integrity and Call Stack Protection. Contrary to existing

Keywords

concolic testing, Erlang, software testing, dynamic symbolic execution, SMT solving

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Chapter 1

Introduction

1.1 Motivation

Computer hardware and software continuously grow in size and complexity and as a result ensuring the absence of exploitable behaviors is becoming increasingly difficult. In the era when (**Feedback:** where?)computer systems are used extensively to carry important information (e.g. credit card numbers, national security documents), it has been widely accepted that security of these systems is a priority. Researchers have identified a number of potential vulnerabilities which arise from the violation of known but in-practice unenforceable safety and security policies.

So far, computer security has been delegated mostly to software, while the hardware is being almost completely controlled by the software. High-level languages are becoming more widely used, due to features such as strong type systems with type inference and automatic memory management, making programming less error prone and reducing the number of exploitable bugs. Furthermore, in order to strengthen the security of computing systems a variety of low-level mitigation techniques (**TODO:** reference some? stack canaries, ASLR, $W \oplus X$) have been proposed, however these are mostly ad-hoc solutions designed to prevent specific known attacks, rather than enforcing a security policy by preventing a well-defined class of attacks, thus making it hard to reason about their effectiveness. In fact most of these mitigation techniques can be circumvented by attackers, (**TODO:** reference; Overcoming CFI; Eternal War in Memory) which has lead to a continuous "chase" between attackers and security researchers.

One common attack technique is to exploit some low-level vulnerability such as a buffer overflow to redirect the control flow to attacker injected code. This attack can be stopped by a simple protection scheme known as $W \oplus X$, which enforces that a memory page is either executable or writable but not both. Unfortunately, clever attack techniques can bypass $W \oplus X$. In particular, attackers have been using code-reuse attacks (e.g. return/jump - oriented programming) that allows them to chain together existing pieces of code to achieve malicious behavior without directly introducing new code. Abadi *et al.* [1] introduced a property called Control Flow Integrity (CFI), which provides effective protection against control-flow hijacking attacks. CFI enforces that any execution of a program will respect a statically computed control flow graph (CFG). (**Feedback:** missing references throughout)

The main contribution of this thesis is the formalization and verification of a dynamic monitor for CFI, based on a generic hardware-software security mechanism. We provide a precise attacker model and prove in Coq that the monitor enforces a variant of the CFI

property proposed by Abadi *et al.* [1]. To obtain this result we prove refinement between a concrete machine running a monitor satisfying our Coq specification and an abstract machine having CFI by construction. We conclude the proof using a novel generic result stating that under certain assumptions CFI is preserved by refinement. (**Feedback:** Is there anything missing here?)

1.2 Thesis Outline

Map 1. Intro 2a. Safety and Security Policies 2b. Micropolicies 3. CFI description 4. CFI formalization 5. Conclusions and Future work 6. Related work

Chapter 2 of this thesis briefly describes the basic requirements a security policy must satisfy and puts into context the framework we utilize in order to formalize and enforce a Control-Flow Integrity (CFI) policy. Chapter 3 discusses the current state of research on Control-Flow Integrity and clarifies our goals and contributions to it. Chapter 4 describes in detail the design of a fine-grained CFI policy and how we used the framework from Chapter 2 in order to enforce the policy and formally reason about it's security properties. Chapter 5.. conclusions, future work? Chapter 6.. related work and bibliography? Appendix with code and/or step relations etc.?

Chapter 2

Micro-policies: Verified, Hardware-Assisted Monitors

Currently the hardware provides very limited security mechanisms (**TODO:** name some; 4 protection rings, page-level memory protection via virtual memory), leaving most of the work to the software. This requires that the software performs various sanity-checks during an execution and that it carefully maintains various safety and security invariants, a tedious and error-prone task that results in high runtime performance overheads.

Many potentially effective mitigation techniques are not deployed because of the performance overhead they incur. Another requirement for deployment of a protection mechanism is the compatibility with existing executables and the degree of intervention required by a human. Usually even making slight changes to a code and redistributing has high cost and the protection mechanism is likely to see very low adoption.

The lack of efficient and effective generic ways to enforce security policies, forces programmers to protect their own code, a task which is not trivial even for the small and simply programs. As a result most, if not all, software carries weaknesses which can be exploited by an attacker. "Safe" languages, automate some of the checks required and eases the work of the programmer, for example by implementing array bounds checking or by disallowing pointer-arithmetic. However these solutions only reduce the chance of introducing exploitable bugs in a program and do not enforce stricter, more effective policies such as Control Flow Integrity or complete Memory Safety (spatial/temporal protection for heap and stack). In addition, we still need effective and efficient protection mechanisms for a plethora of software written in unsafe languages such as C.

2.1 Micro-Policies

(**Feedback:** Can the main idea of the CFI micro-policy be introduced here already? See grant proposal.)

A wide range of security policies can be enforced by associating metadata to the data being processed (e.g., this is an instruction, this is from the network, this is private, etc.), propagating the metadata as instructions are executed and using a set of rules on the metadata to check whether a policy is violated and how the tags should be propagated.

Abstractly, these rules form a partial function from a set of input tags to a set of output tags

$$(opcode, tag_{pc}, tag_{instr}, tag_{arg1}, tag_{arg2}, tag_{arg3}) \rightarrow (tag_{pc'}, tag_{result})$$

informally read as, "if the next instruction to be executed is opcode, the current tag of

the program counter is pc_{tag}, the current tag on the instruction location is tag_{instr} and the tags on the operands of the instruction are tag_{arg1}, tag_{arg2} and tag_{arg3} then if execution of the instruction is allowed the tag on the program counter should be set to $tag_{pc'}$ and any new data created by the instruction should be tagged tag_{result}".

More specific, a micro-policy is made up of the following elements:

1. a set of *metadata tags* used to tag the contents of the memory and all the registers as well as the pc.
2. a *transfer function* that implements the checks on the tags and the tag propagation as described above.
3. a tagging scheme for the initial state of the machine.
4. for some micro-policies, a set of *monitor services* (i.e., privileged code) that can be invoked by user code.

Furthermore, as we will see in 2.4, a software-hardware mechanism that allows for implementation of micro-policies without sacrificing flexibility (in terms of the policies that can be enforced) has already been designed. Simulations and benchmarks show that the runtime overhead is very low compared to dedicated software solutions thus making it a realistic and appealing way to deploy a wide range of security policies in future computing systems.

2.2 Example: Non-Writable Code & Non-Executable Data

(CH: Make it clearer that this is informal and you will return to the formalization later on)

(CH: Symbolic vs concrete rules ... should introduce symbolic rules first, although this is a quite trivial example; ALT: write these as pseudo-Coq functions?)

(CH: Rename Instr to Code?)

Consider the set of tags $\mathcal{T} = \{Data, Instr\}$. If we initially tag all executable regions in memory as *Instr*and all non-executable as *Data* then we can enforce *NWC* by two rules of the form

$$\frac{}{Store : \{CI{=}Instr, MR{=}Data\} \to \{PC'{=}-, RES{=}Data\}} \text{(Store/Data)}$$

$$\frac{opcode \notin \{Store\}}{opcode : \{CI{=}Instr\} \to \{PC'{=}-, RES{=}-\}} \text{(Rest)}$$

Figure 2.1: Rules enforcing *NWC* and *NXD*

The dashes in the result vector, represent *don't care* values, meaning we will not use their values for anything, so any tag (usually a default tag set by the policy designer) can be used. Furthermore, we are omitting from the input vector the fields that are unused by the transfer function. For this simply policy, the transfer function only uses the tag on the current instruction (CI) and in the case of a Store instruction the tag on the memory (MR), i.e., the tag on the memory location we are trying to write. Additionally, if no rule applies, execution of the instruction is disallowed. Informally the above rules can be read

as "If the tag on the current instruction is *Instr* , then if the opcode of the instruction is Store, execution should be allowed only if the tag of the third operand is *Data*. In that case the tag on the new data on the memory should remain *Data*. If the opcode of the instruction is not Store, then it's allowed and the result tags are indifferent for the enforced policy."

2.3 Generic Verification Framework for Micro-Policies

Unsurprisingly, designing a security policy, reasoning about it's effectiveness against potential attackers and encoding it as a micro-policy can become a complex task. Azevedo *et al.* [8] built a generic framework for defining micro-policies on top of a simple machine modeling a RISC processor augmented with the PUMP hardware (referred to as concrete machine), formalized this framework in Coq and used it to define and formally verify micro-policies for dynamic sealing, control-flow integrity, memory safety, compartmentalization and protecting the enforcement mechanism (monitor) itself.

The framework offers a high-level machine, called the symbolic machine, that abstracts away from unnecessary low-level implementation details and can be used as an interface to a concrete machine, simplifying the work of the micro-policy designer and allowing him to define and reason about a micro-policy using structured mathematical objects than low-level machine words.

In order to implement the micro-policy at the concrete machine level, one needs to additionally provide machine code that implements the transfer function, an encoding of tags to words and machine code for any monitor services that the micro-policy may use. The relation between the symbolic and the concrete machine is formally defined as a two-way refinement (forward and backward). This is a generic refinement proof, parameterized by the encoding of the symbolic tags to words and a proof of correctness of the monitor code for a micro-policy. The designer of a micro-policy can use this two-way refinement simply by providing these two parameters.

2.3.1 Correctness of micro-policies

For each micro-policy an abstract machine, which serves as a specification to the invariants the policy designer wants to enforce, is defined. The abstract machine is "correct" by construction, meaning that it's designed to respect those invariants. Using the symbolic machine as an intermediate step to simplify the proofs, by proving a refinement between the symbolic and the abstract machine and by utilizing the generic refinement between the symbolic and the concrete machine, we can prove a refinement between the abstract and the concrete machine, thus showing that every valid step for the concrete machine is also a valid step for the abstract machine.

All the machines introduced in the original paper by Azevedo *et al.* [8], as well as this thesis, have a similar structure. In particular, they share a common RISC-based instruction set (with a few - uninteresting for the scope of this thesis - exceptions) and they have a fixed number of general-purpose registers, along with a pc register. Of course the abstract machine defined by the policy designer can differ in various ways, but more similarities with the symbolic machine implies easier proofs of correctness.

(**Feedback:** Introduce the (names of the) various machines and how they relate to each-other. Nice diagram?)

(**TODO:** Write instruction set)

2.3.2 Symbolic Machine

As mentioned above, the symbolic machine enables us to abstract away from various low-level details of the concrete machine. We can express and reason about policies in terms of mathematical objects written in Gallina rather than machine code and the corresponding proofs for the concrete machine comes for free under some assumptions. In essence, the symbolic machine is parameterized by a micro-policy as it was defined in 2.1, with the addition of an internal state that can be used by monitor services.

The states of the symbolic machine consists of the memory, the registers, the *pc* register and the internal state. The memory and register contents, as well as the *pc*, are all tagged with a symbolic tag *t*. We name their contents *symbolic atoms* referred to with the notation $w@t$, where w is the value (word) and t is the tag.

At each step, a record named *mvector* is formed. It consists of the current opcode, the tag on the *pc*, the tag on the current instruction and optionally up to three tags depending on the opcode of the instruction. The *mvector* is passed to the transfer function which decides whether the step violated the policy enforced by the *transfer* function and in this case halts the machine, or if no violation occurred returns a tag for the new *pc* and a tag for any results the instruction execution produced.

In fig. 2.2 we give, in form of inference rules, the stepping relation for the Symbolic machine, demonstrating how the transfer function and the tag propagation works at each step.

Notice for example, that when a store instruction executed, the tag on the memory location to be overwritten is fetched, allowing the *transfer* function to know what kind of data we are trying to overwrite. Returning to the example micro-policy in 2.2 we can define the transfer function that is used by the symbolic machine as a Coq function.

(**TODO:** Actually define it - but wait for beautified coq code to kick in)

2.4 A Programmable Unit for Metadata Processing

(CH: Could consider moving this one level up (turn it into chapter))

2.4.1 Hardware Architecture

The Programmable Unit for Metadata Processing (PUMP) architecture [11] allows us to efficiently implement a wide range of micro-policies [10], using software to define the rules enforcing the policy, while the hardware provides efficiency by undertaking the propagation of the tags and a rules cache.

On the hardware level, the PUMP is an extension to a conventional RISC architecture. Every word of data in the machine - whether in memory or a register, is extended with a word-sized metadata tag. These tags are not interpreted by hardware, instead the interpretation of the tags is left to the software, thus making it easy to implement new policies on the metadata. Since tags are word-sized, they can be pointers to complex data-structures of tags, such as tuples of tags, allowing for complex policies to be expressed and multiple orthogonal policies to be enforced in parallel.

The hardware undertakes the correct propagation of tags from operands to results according to the rules defined by the software. A hardware rule cache mapping sets of input tags to sets of output tags is used for common case efficiency. On each instruction dispatch, in parallel with the usual behavior of an instruction (e.g., execution of an addition in the ALU), the hardware forms the set of input tags and a lookup is performed on the

$$\frac{mem[pc] = i@t_i \qquad decode\ i = Nop \qquad Nop : \{PC{=}t_{pc}, CI{=}t_i\} \to \{PC'{=}t'_{pc}, RES{=}{-}\}}{(mem, reg, pc@t_{pc}, int) \to (mem, reg, pc + 1@t'_{pc}, int)} \tag{Nop}$$

$$\frac{\begin{array}{c} mem[pc] = i@t_i \qquad decode\ i = Const\ n\ r \qquad reg[r]{=}w_{old}@t_{old} \\ Const : \{PC{=}t_{pc}, CI{=}t_i, OP1{=}t_{old}\} \to \{PC'{=}t'_{pc}, RES{=}t_{res}\} \\ reg' = reg[r{\leftarrow}n@t_{res}] \end{array}}{(mem, reg, pc@t_{pc}, int) \to (mem, reg', pc + 1@t'_{pc}, int)} \tag{Const}$$

$$\frac{\begin{array}{c} mem[pc] = i@t_i \qquad decode\ i = Mov\ r_p\ r_s \\ reg[r_p]{=}w@t_p \qquad reg[r_s]{=}w_{old}@t_{old} \\ Mov : \{PC{=}t_{pc}, CI{=}t_i, OP1{=}t_p, OP2{=}t_{old}\} \to \{PC'{=}t'_{pc}, RES{=}t_{res}\} \\ reg' = reg[r_s{\leftarrow}w@t_{res}] \end{array}}{(mem, reg, pc@t_{pc}, int) \to (mem, reg', pc + 1@t'_{pc}, int)} \tag{Mov}$$

$$\frac{\begin{array}{c} mem[pc] = i@t_i \qquad decode\ i = Binop\ op\ r_p\ r_s\ r_t \\ reg[r_p]{=}w_p@t_p \qquad reg[r_s]{=}w_s@t_s \qquad reg[r_t]{=}w_{old}@t_{old} \\ Binop\ op : \{PC{=}t_{pc}, CI{=}t_i, OP1{=}t_p, OP2{=}t_s, MR{=}t_{old}\} \to \{PC'{=}t'_{pc}, RES{=}t_{res}\} \\ reg' = reg[r_t{\leftarrow}(w_p op w_s@t_{res}] \end{array}}{(mem, reg, pc@t_{pc}, int) \to (mem, reg', pc + 1@t'_{pc}, int)} \tag{Binop}$$

$$\frac{\begin{array}{c} mem[pc] = i@t_i \qquad decode\ i = Load\ r_p\ r_s \\ reg[r_p]{=}w_p@t_p \qquad mem[w_p]{=}w@t_{mem} \qquad reg[r_s]{=}w_{old}@t_{old} \\ Load : \{PC{=}t_{pc}, CI{=}t_i, OP1{=}t_p, OP2{=}t_{mem}, MR{=}t_{old}\} \to \{PC'{=}t'_{pc}, RES{=}t_{res}\} \\ reg' = reg[r_s{\leftarrow}(w@t_{res}] \end{array}}{(mem, reg, pc@t_{pc}, int) \to (mem, reg', pc + 1@t'_{pc}, int)} \tag{Load}$$

$$\frac{\begin{array}{c} mem[pc] = i@t_i \qquad decode\ i = Store\ r_p\ r_s \\ reg[r_p]{=}w_p@t_p \qquad reg[r_s]{=}w_s@t_s \qquad mem[w_p]{=}w_{old}@t_{old} \\ Store : \{PC{=}t_{pc}, CI{=}t_i, OP1{=}t_p, OP2{=}t_s, MR{=}t_{old}\} \to \{PC'{=}t'_{pc}, RES{=}t'_d\} \\ mem' = mem[w_p{\leftarrow}w_s@t'_d] \end{array}}{(mem, reg, pc@t_{pc}, int) \to (mem', reg, pc + 1@t'_{pc}, int)} \tag{Store}$$

$$\frac{mem[pc] = i@t_i \qquad decode\ i = Jump\ r \qquad reg[r]{=}w@t_w \qquad Jump : \{PC{=}t_{pc}, CI{=}t_i, OP1{=}t_w\} \to \{PC'{=}t'_{pc}, RES{=}{-}\}}{(mem, reg, pc@t_{pc}, int) \to (mem, reg, w@t'_{pc}, int)} \tag{Jump}$$

$$\frac{mem[pc] = i@t_i \qquad decode\ i = Bnz\ r\ n \qquad reg[r]{=}w@t_w \qquad Bnz : \{PC{=}t_{pc}, CI{=}t_i, OP1{=}t_w\} \to \{PC'{=}t'_{pc}, RES{=}{-}\}}{(mem, reg, pc@t_{pc}, int) \to (mem, reg, w@t'_{pc}, int)} \tag{Bnz}$$

$$\frac{\begin{array}{c} mem[pc] = i@t_i \qquad decode\ i = Jal\ r \\ reg[r]{=}w@t_w \qquad reg[ra]{=}w_{old}@t_{old} \\ Jal : \{PC{=}t_{pc}, CI{=}t_i, OP1{=}t_w, OP2{=}t_{old}\} \to \{PC'{=}t'_{pc}, RES{=}{-}\} \\ reg' = reg[ra{\leftarrow}pc + 1@t_{res}] \end{array}}{(mem, reg, pc@t_{pc}, int) \to (mem, reg', w@t'_{pc}, int)} \tag{Jal}$$

$$\frac{\begin{array}{c} mem[pc] = \varnothing \qquad get\_service\ pc = (t_i, f) \\ Service : \{PC{=}t_{pc}, CI{=}t_i\} \to \{PC'{=}t'_{pc}, RES{=}{-}\} \\ f\ (mem, reg, pc, int,) = (mem', reg', pc', int,) \end{array}}{(mem, reg, pc@t_{pc}, int) \to (mem', reg', pc'@t'_{pc}, int')} \tag{Service}$$

Figure 2.2: Stepping relation for the symbolic machine

rule cache. If the lookup is successful a set of output tags is returned and combined with the results of the normal execution of the instruction a new state is produced. On the other hand, if the lookup failed, the hardware invokes a trusted piece of system software - the fault handler - which checks the input tags and decides whether the execution should be allowed or not. In the first case, the fault handler returns a set of result tags, a pair of set of input and output tags is formed and inserted into the rules cache, while the faulting instruction is restarted and will now hit the cache. Otherwise, execution of this instruction violated some rules of the enforced policy and execution should not continue normally (e.g., should be halted).

As described in the original PUMP paper by Dehon *et al.* [11] and in [10] a rich set of effective security policies can be efficiently implemented using the architecture mentioned above. In particular, implementations of dynamic typing, memory safety for heap-based data, control flow integrity and taint tracking are described, evaluated against a specific threat model and benchmarked. The benchmarks are done using a simulation of the described hardware and the two papers claim low overhead (3% on average) for each of the policies named above.

Compared to other software solutions for enforcing security policies, the PUMP offers significantly lower overhead, thanks to dedicated hardware assistance, while the fact that interpretation of the metadata is done by software offers flexibility with regard to the policies that can be implemented, compared to hardware solutions implementing a specific policy.

While the PUMP offers flexibility at a low runtime performance overhead, there are more overheads associated to such a mechanism. For example adding metadata to all the data in the machine, would result in a 100% memory overhead. In addition, the extra hardware and the rule cache along with potentially larger memories could result into a 400% overhead on energy usage. [10] The authors claim that a careful and well-optimized implementation can reduce these numbers, resulting in a 50% energy overhead. (**Feedback:** use optimized numbers)

2.4.2 Concrete Machine Modeling PUMP Architecture

The concrete machine is a model of the PUMP architecture, modeling a RISC machine with a rules *cache* and the software *miss handler*. The instruction set has been extended with four additional instructions that are meant to be used by monitor code only, a restriction that is enforced by the monitor self-protection micro-policy.

The state of the concrete machine consists of the memory, the registers, the *pc* register, the *epc* register a special purpose register that holds the address of the faulting instruction after a cache miss and a rules cache. The cache works as a key-value store where a key is an *input vector* that contains an instruction opcode, the tag of the current instruction, the tag of the pc and up to three operand tags, and a value is an *output vector* which contain a tag for the new pc and a tag for any results from the execution of the instruction. In the context of the concrete machine a tag is the encoding into a word of a symbolic tag. Lifting this encoding relation to vectors, we get that a concrete vector is the encoding of a symbolic vector (*mvector*). Similar to the symbolic machine the contents of the memory, the registers, the pc and the epc are concrete atoms w@t where w is a word and t is the encoding of a tag into a word.

The stepping relation for the concrete machine is a bit more complicated than the one for the symbolic machine. In particular, on each step the machine forms the *input vector* and looks it up in the cache. If the lookup succeeds then the instruction is allowed, an

output vector is returned by the cache and the next state is tagged according to it. If the lookup fails, then the *input vector* is saved in memory, the current *pc* is stored in *epc* and the machine traps to the *miss handler*. The above are demonstrated in the two example rules in fig. 2.3.

$$\frac{\begin{array}{c} mem[pc] = i@t_i \qquad decode\ i = Store\ r_p\ r_s \\ reg[r_p]{=}w_p@t_p \qquad reg[r_s]{=}w_s@t_s \qquad mem[w_p]{=}w_{old}@t_{old} \\ cache \vdash (Store, t_{pc}, t_i, t_p, t_s, t_{old}) \mapsto (t'_{pc}, t'_d) \\ mem' = mem[w_p{\leftarrow}w_s@t'_d] \end{array}}{(mem, reg, pc@t_{pc}, epc, cache) \rightarrow (mem', reg, (pc{+}1)@t'_{pc}, epc, cache)} \quad \text{(STORE)}$$

$$\frac{\begin{array}{c} mem[pc] = i@t_i \qquad decode\ i = Store\ r_p\ r_s \\ reg[r_p]{=}w_p@t_p \qquad reg[r_s]{=}w_s@t_s \qquad mem[w_p]{=}w_{old}@t_{old} \\ cache \vdash (Store, t_{pc}, t_i, t_p, t_s, t_{old}) \not\mapsto \\ mem' = mem[0..5 \leftarrow (Store, t_{pc}, t_i, t_p, t_s, t_{old})] \end{array}}{(mem, reg, pc@t_{pc}, epc, cache) \rightarrow (mem', reg, trapaddr@Monitor, pc@t_{pc}, cache)} \quad \text{(STORE-MISS)}$$

Figure 2.3: Concrete step rules for Store instruction

Addresses 0 to 5 are used to store the *input vector* and 6 to 7 are used by the miss handler to store the *output vector*. As a side-note, cache eviction is not modeled (an infinite cache is assumed).

2.4.3 Concrete Policy Monitor

Unlike the symbolic machine, where the user cannot cannot change the *transfer* function, enforcing a micro-policy on the concrete machine requires that we are able to protect the memory of the policy monitor and that privileged instructions are not executed by user code. This self-protection policy can be easily composed with another micro-policy and enforced by the infrastructure described above.

Using tags of the form, *User st*, *Entry st*, *Monitor* we can distinguish between user memory, monitor memory and monitor services. In particular *User st* is used to tag a user-level atom, where *st* is the word-encoding of a symbolic tag. *Monitor* is used to tag the monitor memory and a few reserved registers. The *pc* is tagged with *Monitor* when a monitor execution takes place and *User st* when user-code is executed. The tag *Entry st* is used to tag the first instruction of a monitor service and serves as an indication that execution will continue under the privileged *Monitor* mode.

The miss handler is a composed policy monitor that protects itself from *User* code and that enforces a desired micro-policy. One important thing to note is that the miss handler for the concrete machine can take an arbitrary number of steps before deciding that no violation occurred and returning to *User* mode, unlike the symbolic *transfer* function that does not need to take any steps.

Chapter 3

Control-Flow Integrity

Restricting the control-flow of a program in some way is a technique widely spread among security researchers. For example non-executable data (NXD) can be considered as a form of (very) coarse-grained *CFI* where control-flow is not allowed to reach any memory region that holds non-executable data. Other mitigation techniques such as protecting return addresses on the stack enforce a form of coarse-grained *CFI*.

Moreover it is common that security properties are enforced dynamically by code that is statically injected to the program (e.g., Inlined Reference Monitors (IRM) [12] follow that approach), thus some form of *CFI* is required in order to ensure that these checks are not circumvented.

(**TODO:** Think about title)

3.1 Related Work

3.1.1 Balancing between performance and security

Abadi *et al.* first proposed a technique to enforce *CFI* based on IRMs. In particular, they proposed to mark all valid targets of *indirect* control transfers with a unique identifier and inject checks before all indirect jumps (including return instructions). However they assume that any two destinations are equivalent, in the sense that they share the same identifier, if the CFG contains edges from the same set of sources, which may significantly reduce the precision of the CFG. The authors also note that a 2-ID approach where one identifier is used for calls and another for returns could provide adequate security in many cases.

The work of Abadi *et al.* sparked interest of researchers who tried to improve some of the weaknesses of the initial implementation, usually by choosing between performance against precision and vice-versa.

Bletsch *et al.* [4] followed the work of Abadi *et al.*, but changed their checking mechanism to perform the check after the control flow transfer has occurred which, as the authors claim, reduced the cache pressure and resulted in better performance. Precision remains the same with the implementation of Abadi *et al.*.

Zhang *et al.* [21] proposed *Compact Control Flow Integrity and Randomization* (CCFIR), a new efficient way to enforce coarse-grained *CFI*. CCFIR collects all valid targets of indirect control-transfers and stores them in a random order, in a protected section called "Springboard section". Indirect control-transfers are only allowed to addresses that are in the Springboard. Their implementation uses a 3-ID approach where one identifier is used for calls and the two other identifiers are for returns, separating them between returns

to sensitive and non-sensitive functions. Their implementation also supports interaction between protected and un-protected modules, which makes it an attractive solution to coarse-grained *CFI*.

The above techniques are evaluated in [13] where the authors demonstrate code-reuse attacks against binaries protected by coarse-grained *CFI*. These attacks illustrate the need for fine-grained *CFI* which however incurs a high runtime-overhead penalty making deployment of such a mechanism unlikely.

Standard assumptions for effective *CFI* Most -if not all- *CFI* implementations also come with a set of assumptions under which *CFI* holds. Two standard assumptions for all mechanisms that attempt to enforce *CFI* are:

- Non-Executable Data (*NXD*), a security mechanism that disallows execution of data.
- Non-Writable Code (*NWC*). Changing the code of a program would allow an attacker to circumvent dynamic checks.

Both assumptions are fairly standard for modern computers and are enforced through hardware or software. In some cases *NXD* can be lifted, but additional security risks and complexity is not worth the minor advantages offered by such an action.

Many implementations that attempt to do fine-grained *CFI* also require that identifiers used to mark nodes in the CFG are unique.

3.1.2 Coarse-grained CFI Micro-Policy

(CH: consider moving to appendix, or related work section)

We can use the PUMP to implement the coarse-grained *CFI* mechanisms described earlier. Suppose we want to implement 1-ID *CFI*, we tag all indirect flow destinations and sources with a tag *Marked* and the rest of the instructions as *Unmarked*. Executing instructions that are sources of indirect flows, propagates their instruction tag to the *pc*. We then have to check that the tag on the destination matches the tag on the tag on the *pc*.

$$\frac{op \in \{Jump, Jal\}}{op : \{CI{=}Marked\} \rightarrow \{PC'{=}Marked, RES{=}-\}} \quad (\textsc{Mark})$$

$$\frac{op \notin \{Jump, Jal\}}{op : \{PC{=}Marked, CI{=}Marked\} \rightarrow \{PC'{=}Unmarked, RES{=}-\}} \quad (\textsc{Check})$$

$$\frac{op \notin \{Jump, Jal\}}{op : \{PC{=}Unmarked, CI{=}Unmarked\} \rightarrow \{PC'{=}Unmarked, RES{=}-\}} \quad (\textsc{NoCheck})$$

Figure 3.1: Rules enforcing fine-grained *CFI*, *NXD* and *NWC*

Rule *Mark* is used in the case the opcode is Jump or Jal (the only indirect jumps in the RISC machine we examine) and propagates the *Marked* tag on the tag of the new *pc*. Rule *Check* applies when the tag on the *pc* is set to *Marked* and corresponds to a legal destination and rule *NoCheck* corresponds to any instruction that is not a jump source or target.

We do not further study this coarse-grained approach as we consider it ineffective since attacks against it has already been demonstrated in [13]. Instead we are going to focus on implementing and formalizing a fine-grained *CFI* micro-policy.

3.1.3 Formal verification of Control-Flow Integrity

In [2] Abadi *et al.* extended their original paper, with -among other things- a more detailed formal study of *CFI*. Their formalization regarded a much simpler machine than the x86 omitting all the complexity in modern systems. The machine has a few instructions, a separate data memory and instruction memory which by the operational semantics of the machine are non-executable and non-writable (enforcing *NXD* and *NWC* by construction), and a small set of registers. Moreover, their attacker model permits arbitrary changes to the data memory, arbitrary changes to all the registers but a few distinguished ones that are used during the dynamic checks and no changes to the instruction memory. The authors proof that under some assumptions *CFI* is preserved for every step even in the presence of an attacker as powerful as the one described above. Their formal study served as a guideline for the implementation, but as it is done on paper their proofs cannot be machine checked. Furthermore, their formalization omits less interesting but important details such as instruction encoding and decoding which as shown in [15] are far from trivial for the x86.

Machine-checked formal verification efforts include [22], which is a SFI formalization for the ARM architecture that also enforces *CFI*. Their formalization was developed using the HOL theorem prover and a program logic framework they created. However their benchmarks report a 240% runtime overhead. The authors of [7] claim partial proofs for a *CFI* enforcement mechanism focused on the kernel of an operating system. Their runtime overhead can also reach 100%.

3.2 Fine-Grained Control-Flow Integrity Micro-Policy

The PUMP hardware allows us to avoid taking the difficult decision between performance and security. As shown in [10], we can enforce a *fine-grained CFI* policy with an average runtime overhead of less than 3% (maximum overhead of less than 10%), on the SPEC2006 benchmarks.

(CH: Shrink and polish this)(NG: done) We follow the standard approach, by designing a composed micro-policy that enforces *NXD*, *NWC* and *CFI*. We considered designs that lifted the *NXD* and *NWC* restrictions but we rejected them, for the time being, as there did not seem to be any considerable advantages (i.e., compatibility with self-modifying programs, JIT compilers, etc.). Moreover unlike other *CFI* enforcement mechanisms we do not have to rely on the CPU or the operating system to enforce *NXD* and *NWC*, therefore lifting these restrictions would not reduce our assumptions and consequently would not increase our confidence in the robustness of our approach.

Our approach uses unique identifiers to tag the contents of the memory that correspond to sources and potential destinations of indirect flows according to a binary relation (on the identifiers) $\mathcal{CFG}$.

Consider the set of tags $\mathcal{T} = \{Data, Instr\ id, Instr\ \bot\}$ where id is a unique identifier (i.e., used to tag the contents of only one location in the memory). One simply way to achieve this is to use the address of the instruction as it's id, for example an instruction stored at address 100 would be tagged *Instr* 100. This is the approach we take in our development. Adapting the rules from 2.2, we shall use *Data* to tag all contents in memory

that are considered non-executable data, *Instr id* to tag all contents in memory that are considered executable instructions and are sources or targets of indirect control flows and *Instr* $\perp$ to tag all other instructions. The rules to enforce *NWC* and *NXD* are intuitively the same and only change to account for the splitting of the *Instr* tag.

We follow the same idea as with coarse-grained *CFI*, propagating the instruction tag of instructions that are sources of indirect flows to the tag on the *pc* of the next state and upon execution of the next instruction, checking that the tag on the *pc* and on the instruction are in some relation. In the case of coarse-grained *CFI* we required that they match but for fine-grained *CFI* we require that they are in the $\mathcal{CFG}$ relation.

$$\frac{op \in \{Jump, Jal\} \qquad (src, dst) \in \mathcal{CFG}}{op : \{PC{=}Instr\ src, CI{=}Instr\ dst\} \rightarrow \{PC'{=}Instr\ dst, RES{=}-\}} \text{(Flow/Check)}$$

$$\frac{op \in \{Jump, Jal\}}{op : \{PC{=}Data, CI{=}Instr\ dst\} \rightarrow \{PC'{=}Instr\ dst, RES{=}-\}} \text{(Flow/NoCheck)}$$

$$\frac{(src, dst) \in \mathcal{CFG}}{Store : \{PC{=}Instr\ src, CI{=}Instr\ dst, MR{=}Data\} \rightarrow \{PC'{=}Data, RES{=}Data\}} \text{(Store/Check)}$$

$$\frac{ti \in \{Instr\ dst, Instr\ \perp\}}{Store : \{PC{=}Data, CI{=}ti, MR{=}Data\} \rightarrow \{PC'{=}Data, RES{=}Data\}} \text{(Store/NoCheck)}$$

$$\frac{opcode \notin \{Jump, Jal, Store\} \qquad (src, dst) \in \mathcal{CFG}}{opcode : \{PC{=}Instr\ src, CI{=}Instr\ dst\} \rightarrow \{PC'{=}Data, RES{=}-\}} \text{(Rest/Check)}$$

$$\frac{opcode \notin \{Jump, Jal, Store\} \qquad ti \in \{Instr\ dst, Instr\ \perp\}}{opcode : \{PC{=}Data, CI{=}ti\} \rightarrow \{PC'{=}Data, RES{=}-\}} \text{(Rest/NoCheck)}$$

Figure 3.2: Rules enforcing fine-grained *CFI*, *NXD* and *NWC*

We note in the above rules that the tag on the *pc* is *Data* when no check for a control-flow violation is required and *Instr src* where *src* is some id, when an indirect flow instruction was executed and a check for a control-flow violation is required. An important observation is that the rules above allow for one control-flow violation to occur, but disallow the next step and therefore the machine will certainly halt after a violation.

If the PUMP hardware fetched the tag on the memory address the machine is jumping to and passed it as an argument to input vector, as it does in the case of a Store instruction, we would be able to enforce *CFI* with no violations at all. (**TODO:** It can't do that for efficiency reasons?)

Chapter 4

Formally Verified Control-Flow Integrity Micro-Policy

Using the micro-policies framework described in 2.3 we proved that the concrete machine instantiated with a *CFI* micro-policy like the one described in 3.2 *simulates* an abstract machine that has *CFI* by construction. We do this proof by using the symbolic machine as an intermediate step, to prove backwards simulation between the symbolic and the abstract machine and afterwards by leveraging the framework of section 2.3 we obtain a backwards refinement between the concrete and the abstract machine.

In addition, we provide an attacker model for all the machines used and we prove that a property capturing the notion of *CFI* holds even when the attacker tampers with the machine, similarly to what is proposed in [2], but adapted to the setting of our machines. We do this by first proving this property for the abstract machine and then by using a generic preservation theorem we developed we prove that this property is *preserved* by backwards refinement and thus transferring the property to the symbolic and consequently to the concrete machine.

4.1 Representing control-flow graphs

Our approach for enforcing *CFI*, as explained in 3.2, requires that we encode the nodes in the control-flow graph in terms of identifiers, which in turn are used to tag all sources and targets of indirect control-flows.

At this point we take a detour, to point out an important design point of the micro-policies framework and our *CFI* micro-policy. Throughout both developments, a heavily parametric and modular approach was taken. This parametric design is enabled by the use of the *Section* and *Type Classes* mechanisms of Coq. As an example, the node identifiers, along with a number of properties we require of them are expressed by the following interface (defined in terms of a type class):

The *Context* command on the top of the code above, allows us to *assume* that there exists an instance of this interface. In fact, the *machine_types* argument is just another type class, serving as a specification of the various types of the machine (e.g., the word size). This approach allowed us to abstract away from various details and structure our proofs in a clean way. In addition, we can easily instantiate a different machine with minimal changes in our proofs and definitions (e.g., instantiate the machine with a different word size).

However, one drawback is that one wrong specification in a type class would disallow

```
Context {t : machine_types }.

Class cfi_id := {
  id          : eqType ;

  word_to_id  : word t → option id;
  id_to_word  : id → word t;

  id_to_wordK  : forall x, word_to_id (id_to_word x) = Some x;
  word_to_idK  : forall w x, word_to_id w = Some x → id_to_word x = w

}.
```

Listing 4.1: Interface of node identifiers

us to instantiate it and would require that we go back and change all parts that used this wrong specification (e.g., in our case, the *machine_types* class was widely used). Therefore one should be careful when doing heavy use of such mechanisms.

Returning to the identifiers, looking at the definition in listing 4.1, we require that the type of the identifiers *id* is an Eqtype (has decidable boolean equality) and that there exists conversion functions between elements of type *word* and *id*, satisfying some constraints.

(**Question:** Should I give some intuition, as to why word_to_id is partial? Is it obvious?)

As mentioned in section 3.2, we check for violations of the control-flow with respect to a binary relation (on the identifiers) $\mathcal{CFG}$ which represents the set of allowed (indirect) jumps. We can extend this relation to precisely describe the control-flow of a program, by lifting $\mathcal{CFG}$ to a relation $\mathcal{SUCC}_{\mathcal{CFG}}$ on machine states, that includes the set of allowed targets for the rest of the instructions. Throughout this thesis we use $\mathcal{CFG}$ to refer to the set of allowed jumps. In our Coq development we assumed a translation of the allowed jumps in form of a function on two identifiers.

```
Variable cfg : id → id → bool.
```

Listing 4.2: Function on ids representing the set of allowed jumps

In addition we defined a function valid_jmp (referred to with the notation $\mathcal{J}$) that expresses the set of allowed jumps between words, by using the *word_to_id* function.

```
Definition valid_jmp w1 w2 :=
  match word_to_id w1, word_to_id w2 with
    | Some id1, Some id2 ⇒ cfg id1 id2
    | _, _ ⇒ false
  end.
```

Listing 4.3: Function on words representing the set of allowed jumps

4.2 Control-Flow Integrity Property

Our formalization includes a definition of *CFI*, similar to the one found in [2], which we prove to be true of all our machines. The need for a new definition arises from fundamental differences between our enforcement mechanism on the concrete mechanism and the one used by Abadi *et al.*. In particular, our enforcement-mechanism does not prevent a violation, instead it can detect it after it has occurred by taking an arbitrary number of "protected" (monitor mode) steps before eventually bringing the machine to a halt. This does not have any impact on the security effectiveness of our mechanism, it does however lead to a more complex definition and therefore more complex proofs.

The definition of *CFI* is further parameterized by an attacker model. We model the attacker as a step relation ($\rightarrow_a$). Intuitively the attacker is allowed to change any *user-level* data but not the code of the program and the *pc*, as well as the tags in the case of a tagged machine. This limitations ensures that an attacker cannot directly circumvent the monitor protection mechanism and our user-level policies (*NWC*, *NXD* and *CFI*). To account for attacker steps, the stepping relation is extended as the union of the normal step relation ($\rightarrow_n$), as defined by the machine semantics, and the attacker step relation ($\rightarrow_a$), as defined by the attacker model.

$$\frac{s \rightarrow_n s'}{s \rightarrow s'} \qquad \frac{s \rightarrow_a s'}{s \rightarrow s'}$$

Figure 4.4: Step relation definition

We define a predicate *initial s*, where s is a machine state, that states that s is an initial state. We use this predicate to express some invariants that are preserved through execution (e.g., the initial tagging scheme for the memory). Finally we define a stopping predicate on an execution trace that states that the machine is coming to a halt with respect to normal steps.

Since we want to instantiate the above parameters in a different way for each of our machines, it makes sense to wrap them in a type class which we will instantiate for each machine to get the corresponding definition of *CFI*.

```
Class cfi_machine := {
  state : Type;
  initial : state → Prop;

  step : state → state → Prop;
  step_a : state → state → Prop;

  succ : state → state → bool;
  stopping : list state → Prop
}.
```

Listing 4.5: Interface of a cfi_machine

For a machine of type cfi_machine we give the following definitions:

Definition 4.6 (Trace has CFI)**.** *We say that an execution trace* $s_0 \rightarrow s_1 \rightarrow \ldots \rightarrow s_n$ has CFI *if for all* $i \in [0, \ldots, n)$ *if* $s_i \rightarrow_n s_{i+1}$ *then* $(s_i, s_{i+1}) \in \mathcal{SUCC}_{\mathcal{CFG}}$

(**Question:** The word relation for succ and cfg is strange since they are booleans, is it ok, or does it confuse you, making you believe they are Props?)

The above definition corresponds to the one found in [1], however it is stronger in the sense that it requires that steps that are in the intersection of normal and attacker steps respect the control-flow. If we did not allow for any violations then the above definition would be enough, but since our enforcement mechanism allows for one violation we have to resort to a weaker definition.

Definition 4.7 (CFI)**.** *We say that the machine* $(State, initial, \rightarrow_n, \rightarrow_a, \mathcal{SUCC}_{\mathcal{CFG}}, stopping)$ *has* CFI *with respect to the set of allowed indirect jumps* $\mathcal{CFG}$ *if, for any execution starting from initial state* s_0 *and producing a trace* $s_0 \rightarrow \ldots \rightarrow s_n$*, either*

1. *The whole trace has* CFI *according to definition 4.6, or else*
2. *There is some* i *such that* $s_i \rightarrow_n s_{i+1}$*, and* $(s_i, s_{i+1}) \notin \mathcal{SUCC}_{\mathcal{CFG}}$*, where the sub-traces* $s_0 \rightarrow \ldots \rightarrow s_i$ *and* $s_{i+1} \rightarrow \ldots \rightarrow s_n$ *both have CFI and the sub-trace* $s_{i+1} \rightarrow \ldots \rightarrow s_n$ *is stopping.*

4.3 The Abstract Machine

The abstract machine has *CFI*, *NXD* and *NWC* by construction and will serve as a specification for the symbolic and eventually the concrete machine that implement *CFI* through the tag-based system explained in the previous chapter.

Unlike the symbolic and the concrete machine, this abstract machine splits the memory into two disjoint memories, an instruction memory and a data memory. The instruction memory is fixed (non-writable) and the machine uses this memory to fetch instructions to execute, so *NWC* and *NXD* are enforced by construction.

In addition the state of the machine includes an *ok* bit, indicating whether a control-flow violation has occurred or not. The rest of the machine state is completed by a set of registers and a *pc* register. We use a 5-tuple notation for the state (im, dm, reg, pc, ok), where the first field is the instruction memory, the second the data memory, the third the registers, the fourth is the pc register and the fifth is the *ok* bit.

4.3.1 Operational semantics

Below is the step rule for the Store instruction, illustrating both *NWC* and *NXD*. Notice that the instruction is fetched by the instruction memory and the store is done on the data memory.

$$\frac{\begin{array}{c} im[pc] = i \qquad decode\ i = Store\ r_p\ r_s \qquad reg[r_p] = p \\ reg[r_s] = w \qquad dm' = dm[p \leftarrow w] \end{array}}{(im, dm, reg, pc, true) \rightarrow (im, dm', reg', pc + 1, true)} \qquad \text{(STORE)}$$

Figure 4.8: Step rule for Store instruction of abstract machine

In the above rule, the *ok* bit is true for both the starting and the resulting state. In fact, the machine can take a step only when the *ok* bit is set to true. In the above rule, the *ok* bit is set to true in the resulting state, indicating that no control-flow violation has happened, as expected by the execution of a Store instruction. Control-flow violations in the *NWC* setting our machine is executing, can only occur from *indirect* jump instructions,

in our case the Jump and Jal instructions. Upon execution of a Jump or Jal instruction, we consult $\mathcal{J}$ (see listing 4.3) to check whether the change of control-flow is legal. If the jump is not allowed according to $\mathcal{J}$ then the jump is taken but the *ok* bit is set to false, which will halt the machine in the next step, as it is only allowed to step when the *ok* bit is set to true. Otherwise the *ok* bit will remain true.

$$\frac{im[pc] = i \qquad decode\ i = Jal\ r \qquad reg[r] = pc' \qquad reg' = reg[ra \leftarrow pc + 1] \qquad ok = (pc, pc') \in \mathcal{J}}{(im, dm, reg, pc, true) \rightarrow (im, dm, reg', pc', ok)} \quad (\textsc{Jal})$$

$$\frac{im[pc] = i \qquad decode\ i = Jump\ r \qquad reg[r] = pc' \qquad ok = (pc, pc') \in \mathcal{J}}{(im, dm, reg, pc, true) \rightarrow (im, dm, reg', pc', ok)} \quad (\textsc{Jump})$$

Figure 4.9: Step rule for Jump and Jal instruction of abstract machine

As the abstract machine serves as a specification to a machine with *CFI*, a more intuitive definition of it would not include the *ok* bit and would only allow the Jump and Jal instructions to step if they do not violate the control-flow graph. However, this abstract machine would not allow for any violations to occur unlike our enforcement mechanism for the symbolic and the concrete machine and would lead to more complex simulation proofs, therefore we do not favor it.

The abstract machine also allows for monitor services to be included, although the *CFI* enforcement mechanism does not require any. We assume that a monitor service is a privileged action and that it's execution does not violate the control-flow of the program. Execution of a monitor service is done simply by jumping to it's address, there is no separate instruction. As with all other instructions, execution of the monitor service is only allowed if the *ok* bit is set to true.

$$\frac{pc \notin dom(im) \qquad pc \notin dom(dm) \qquad get\_service\ pc = (addr, f) \qquad f\ (im, dm, reg, pc, true) = (im, dm', reg', pc', true)}{(im, dm, reg, pc, true) \rightarrow_n (im, dm', reg', pc', true)} \quad (\textsc{Service})$$

Figure 4.10: Step rule for monitor services of abstract machine

(**TODO:** Put all rules in appendix?)

4.3.2 Attacker model

The attacker for the abstract machine is allowed to change the contents of the data memory and the registers but not the rest of the state.

$$\frac{dom\ dm = dom\ dm' \qquad dom\ reg = dom\ reg'}{(im, dm, reg, pc, ok) \rightarrow_a^A (im, dm', reg', pc, ok)}$$

Figure 4.11: Attacker model for the abstract machine

4.3.3 Allowed control-flows for the abstract machine

We can construct a function $\mathcal{SUCC}_{\mathcal{CFG}}^{\mathcal{A}}$ for the abstract machine that represents the set of allowed control-flows for all instructions, by extending the set of allowed jumps $\mathcal{CFG}$ we

introduced earlier.

Below we give a specification of the $\mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}$ function for the abstract machine, in form of inference rules. A function is defined in the actual Coq development.

$$\frac{im[pc] = i \qquad decode\ i \in \{Jal\ r, Jump\ r\} \qquad (pc, pc') \in \mathcal{J}}{((im, dm, reg, pc, ok), (im, dm', reg', pc', ok)) \in \mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}}\ (\textsc{IndirectFlows})$$

$$\frac{\begin{array}{c} im[pc] = i \qquad decode\ i = Bnz\ r\ imm \\ (pc' = pc + 1) \vee (pc' = pc + imm) \end{array}}{((im, dm, reg, pc, ok), (im, dm', reg', pc', ok)) \in \mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}}\ (\textsc{ConditionalFlows})$$

$$\frac{\begin{array}{c} im[pc] = i \qquad decode\ i \notin \{Jal\ r, Jump\ r, Bnz\ r\ imm, \varnothing\} \\ pc' = pc + 1 \end{array}}{((im, dm, reg, pc, ok), (im, dm', reg', pc', ok)) \in \mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}}\ (\textsc{NormalFlows})$$

$$\frac{\begin{array}{c} im[pc] = \varnothing \qquad dm[pc] = \varnothing \\ get\_service\ pc = (addr, f) \end{array}}{((im, dm, reg, pc, ok), (im, dm', reg', pc', ok)) \in \mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}}\ (\textsc{ServiceFlows})$$

Figure 4.12: Allowed control-flows for instructions of the abstract machine

Notice that a monitor service is allowed to return anywhere. As we mentioned before, monitor services, execute in a protected by the monitor environment where we assume that an attacker who can only tamper the machine at the user level cannot interfere.

4.3.4 Stopping predicate for the abstract machine

Finally, we define what it means for the abstract machine to be "stopping" by defining a predicate on execution traces:

1. All states in the trace are stuck with respect to normal steps ($\rightarrow_n$)
2. All steps in the trace are attacker steps ($\rightarrow_a$)

4.3.5 CFI proof for the Abstract Machine

Regarding initial states, we only require that the *ok* bit is set to true. We can now instantiate the class of the machines defined in listing 4.5, with the abstract machine and that the abstract machine has *CFI* according to definition 4.7. We first prove a helpful lemma.

Lemma 4.13 (Step Intersection)**.** *For all states st st' such that* $st \rightarrow^A_a st'$ *and* $st \rightarrow_n st'$, $(st, st') \in \mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}$.

Proof.

- By the relation $st \rightarrow_n st'$ we know that the *ok* bit of *st* is set to true.
- The relation $st \rightarrow^A_a st'$ retains the *ok* bit of *st*, therefore *st'* has the *ok* bit set to true.
- It trivially follows from the definition of $\mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}$ that $(st, st') \in \mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}$.

□

Theorem 4.14 (Abstract CFI)**.** *The abstract machine has the* CFI *property stated by definition 4.7.*

Proof. The proof proceeds by induction on the execution trace.

- **Base Case** In this case the execution trace is made up of a single step $st \to st'$. We proceed with case analysis on the step.
  - **Attacker Step** By lemma 4.13 we note that an attacker step cannot be a normal step outside the $\mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}$ relation. Thus in this case the whole trace has *CFI* according to definition 4.6.
  - **Normal Step** By case analysis, if $(st, st') \in \mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}$ then trivially the whole trace has *CFI*. Otherwise $(st, st') \notin \mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}$ and the sub-traces *st* and *st'* vacuously have *CFI*. In addition the sub-trace *st'* is stopping, as the *ok* bit of *st'* is set to false and the state is stuck with respect to normal steps.
- **Inductive Case** In this case the execution trace is extended by an additional step at it's beginning $\mathbf{s_0} \to \mathbf{s_1} \to s_2 \to \ldots \to s_n$. By the induction hypothesis either:
  - The trace $s_1 \to s_2 \to \ldots \to s_n$ has *CFI*, by case analysis if $(s_0, s_1) \in \mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}$ the whole trace has *CFI*. Otherwise $(s_0, s_1) \notin \mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}$, the sub-trace s_0 vacuously has *CFI* and the sub-trace $s_1 \to \ldots \to s_n$ has *CFI* by the induction hypothesis. Additionally, the sub-trace $s_1 \to \ldots \to s_n$ is stopping because:
    * The whole trace is made up of attacker steps. Since $(s_0, s_1) \notin \mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}$ the *ok* bit of s_1 will be set to false and a normal step is not allowed by the operational semantics, while attacker steps retain the *ok* bit.
    * The whole trace is stuck with respect to normal steps. Trivial from the above.
  - There exists a step $s_{v1} \to_n s_{v2}$ such that $(s_{v1}, s_{v2}) \notin \mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}$ and the sub-traces $s_1 \to \ldots \to s_{v1}$ and $s_{v2} \to \ldots \to s_n$ both have *CFI* and the later is also a stopping trace.
    * If $(s_0, s_1) \in \mathcal{SUCC}^{\mathcal{A}}_{\mathcal{CFG}}$ then definition 4.7 still holds and the sub-trace $s_1 \to \ldots \to s_{v1}$ is extended by one step to $s_0 \to \ldots \to s_{v1}$.
    * Otherwise the *ok* bit for s_1 is set to false and the rest of the trace is stuck with respect to normal steps. However from the induction hypothesis we know that $s_{v1} \to_n s_{v2}$, which is a contradiction.

□

4.4 The Symbolic Machine

The symbolic machine was described in section 2.3.2. Unlike the abstract machine, the symbolic machine has one memory and the distinction between data and executable instructions is made through tags, in a fashion similar to what was shown in sections 2.2 and 3.2. We instantiate the symbolic machine, according to the aforementioned sections, with a set of tags $\mathcal{T} = \{Data, Instr\ id, Instr\ \bot\}$ where id is drawn from the class of identifiers **??**.

```
Context {ids : · cfi_id t}.

Inductive cfi_tag : Type :=
| INSTR : option id → cfi_tag
| DATA : cfi_tag .
```

Listing 4.15: Coq definition of Symbolic tags

Although enforcement of *CFI* does not require any monitor services we expose the monitor services mechanism and we check whether calls to each monitor service are allowed or not according to the control-flow graph. This is done by assuming a lookup-table of monitor services where each entry has a tag that is used to check for control-flow violations and a semantic function from symbolic state to symbolic state which produces the new machine state after execution of the system call, as shown in fig. 2.2.

We do not need any internal state for this micro-policy therefore, only the transfer function is left to implement.

4.4.1 Transfer Function

We implement the *transfer* function based on the rules found in 3.2, using Gallina to define a function mapping input vectors (mvector) to output vectors (rvector).

(**TODO:** Should I remove the aggressive capitilization above? It may make it less painful on the eye... Thanks to dependent types it also looks super ugly too, probably make it pseudo-code at some point)

Although, the rules in section 3.2 were fairly simply, expressing them using Gallina's pattern matching increased their size. We also experimented, with different ways of writing the transfer function but we decided to stick with the definition above as it's the most straightforward. It's worth to note that bugs in the above definition were easily made apparent when proving theorems involving the transfer function. In fact, an "interesting" experiment was to re-define the above function in a different way and prove the two equivalent. It took two iterations before getting both functions to agree and although for small definitions like the one above, testing or manually reviewing the code will reveal most if not all bugs, the importance of formal verification in software engineering and critical software is made obvious even for definitions that may seem trivial at first. The correctness of the transfer function will come from simulation proofs between the abstract and the symbolic machine.

4.4.2 Attacker model

Similar to the abstract attacker, the symbolic attacker can change all words tagged as *Data* but not the ones tagged as *Instr*. This is expressed by the following relations:

These attacker capabilities on symbolic atoms are lifted to the memory and registers by a pointwise extension. (**TODO:** be more specific about pointwise?)

4.4.3 Allowed control-flows for the Symbolic Machine

Similar to the abstract machine of section 4.3.3, we construct $\mathcal{SUCCS}^{\mathcal{S}}_{\mathcal{CFG}}$ for the symbolic machine (fig. 4.19) by extending the set of allowed jumps $\mathcal{CFG}$.

```
Definition cfi_handler (ivec : Symbolic.IVec cfi_tags) :
           option (Symbolic.OVec cfi_tags (Symbolic.op ivec)) :=
  match ivec return
        option (Symbolic.OVec cfi_tags (Symbolic.op ivec)) with
  | mkIVec (JUMP as op) (INSTR (Some n)) (INSTR (Some m)) _
  | mkIVec (JAL as op) (INSTR (Some n)) (INSTR (Some m)) _ ⇒
    if cfg n m then
      Some (·mkOVec cfi_tags op (INSTR (Some m)) (default_rtag op))
    else
      None
  | mkIVec (JUMP as op) DATA (INSTR (Some n)) _
  | mkIVec (JAL as op) DATA (INSTR (Some n)) _ ⇒
    Some (·mkOVec cfi_tags op (INSTR (Some n)) (default_rtag op))
  | mkIVec JUMP DATA (INSTR None) _
  | mkIVec JAL DATA (INSTR None) _ ⇒
    None
  | mkIVec STORE (INSTR (Some n)) (INSTR (Some m))
[_ ; _ ; DATA] ⇒
    if cfg n m then Some (·mkOVec cfi_tags STORE DATA DATA) else None
  | mkIVec STORE DATA (INSTR _) [_ ; _ ; DATA] ⇒
    Some (·mkOVec cfi_tags STORE DATA DATA)
  | mkIVec STORE _ _ _ ⇒ None
  | mkIVec op (INSTR (Some n)) (INSTR (Some m)) _ ⇒
    (* this includes op = SERVICE *)
    if cfg n m then
      Some (·mkOVec cfi_tags op DATA (default_rtag op))
    else
      None
  | mkIVec op DATA (INSTR _) _ ⇒
    (* this includes op = SERVICE, fall-throughs checked statically *)
    Some (·mkOVec cfi_tags op DATA (default_rtag op))
  | mkIVec _ _ _ _ ⇒ None
  end.
```

Listing 4.16: Transfer function for symbolic machine in Gallina

$$\frac{}{w_1@Data \rightarrow_a^S w_2@Data} \quad \text{(AttackData)}$$

$$\frac{}{w_1@Instr\ id \rightarrow_a^S w_1@Instr\ id} \quad \text{(AttackInstr)}$$

Figure 4.17: Attacker capabilities

$$\frac{mem \rightarrow_a^S mem' \qquad reg \rightarrow_a^S reg'}{(mem, reg, pc@t_{pc}, int) \rightarrow_a^S (mem', reg', pc@t_{pc}, int)}$$

Figure 4.18: Attacker model for the Symbolic machine

$$\frac{\begin{array}{c} mem[pc] = i@t_i \qquad decode\ i \in \{Jal\ r, Jump\ r\} \\ mem[pc'] = i'@t'_i \qquad t_i = Instr\ src \qquad t'_i = Instr\ dst \\ (src, dst) \in \mathcal{CFG} \end{array}}{((mem, reg, pc, int,), (mem, reg, pc', int,)) \in \mathcal{SUCC}^{\mathcal{S}}_{\mathcal{CFG}}} \quad \text{(IndirectFlows)}$$

$$\frac{\begin{array}{c} mem[pc] = i@t_i \qquad decode\ i \in \{Jal\ r, Jump\ r\} \\ mem[pc'] = \varnothing \qquad get\_service\ pc = (t'_i, f) \\ t_i = Instr\ src \qquad t'_i = Instr\ dst \\ (src, dst) \in \mathcal{CFG} \end{array}}{((mem, reg, pc, int,), (mem, reg, pc', int,)) \in \mathcal{SUCC}^{\mathcal{S}}_{\mathcal{CFG}}} \quad \text{(IndirectFlows2)}$$

$$\frac{\begin{array}{c} mem[pc] = i@t_i \qquad decode\ i = Bnz\ r\ imm \\ (pc' = pc + 1) \vee (pc' = pc + imm) \end{array}}{((mem, reg, pc, int,), (mem, reg, pc', int,)) \in \mathcal{SUCC}^{\mathcal{S}}_{\mathcal{CFG}}} \quad \text{(ConditionalFlows)}$$

$$\frac{\begin{array}{c} mem[pc] = i@t_i \qquad decode\ i \notin \{Jal\ r, Jump\ r, Bnz\ r\ imm, \varnothing\} \\ pc' = pc + 1 \end{array}}{((mem, reg, pc, int,), (mem', reg', pc', int,)) \in \mathcal{SUCC}^{\mathcal{S}}_{\mathcal{CFG}}} \quad \text{(NormalFlows)}$$

$$\frac{mem[pc] = \varnothing \qquad get\_service\ pc = (t'_i, f)}{((mem, reg, pc, int,), (mem', reg', pc', int',)) \in \mathcal{SUCC}^{\mathcal{S}}_{\mathcal{CFG}}} \quad \text{(ServiceFlows)}$$

Figure 4.19: Allowed control-flows for instructions of the symbolic machine

4.4.4 Initial states of the Symbolic Machine

For the symbolic machine, we do require that certain tagging conventions are respected initially. Additionally we prove that these initial conditions are invariants of the machine and they are preserved at every (normal or attacker) step.

These invariants are required for backward simulation between the symbolic and the abstract machine. More invariants may be required for forward simulation.

(**TODO:** hopefully try to get to forward simulation and see what's the situation there.)

Definition 4.20 (Instructions Tagged)**.** *For all addresses addr in the memory such that*

$$mem[addr] = i@Instr\ id$$

addr is in the domain of word_to_id *and additionally*

$$word\_to\_id\ addr = id$$

Definition 4.21 (Entry Points Tagged)**.** *For all addresses addr such that*

$$\begin{array}{l} mem[addr] = \varnothing \\ get\_service\ addr = (it, f) \\ it = Instr\ id \end{array}$$

addr is in the domain of word_to_id *and additionally*

$$word\_to\_id\ addr = id$$

Definition 4.22 (Valid Jumps Tagged)**.** *For all addresses saddr, taddr such that*

$$(saddr, taddr) \in \mathcal{J}$$

it holds that

$$\exists i, mem[saddr] = i@Instr\ (word\_to\_id\ saddr)$$

and either

$$\exists i', mem[taddr] = i'@Instr\ word\_to\_id\ taddr$$

or

$$\begin{aligned} &mem[taddr] = \varnothing \\ &\exists (it, f), get\_service\ addr = (it, f) \\ &it = Instr\ (word\_to\_id\ taddr) \end{aligned}$$

We define a predicate *initial* on symbolic states as the conjunction of the above invariants and the proposition that the tag on the *pc* is set to *Data*.

```
Definition invariants st :=
  instructions_tagged (Symbolic.mem st) ∧
  valid_jmp_tagged (Symbolic.mem st) ∧
  entry_points_tagged (Symbolic.mem st) ∧
  registers_tagged (Symbolic.regs st).

Definition initial (s : Symbolic.state t) :=
  (common.tag (Symbolic.pc s)) = DATA ∧
  invariants s.
```

Listing 4.23: Initial Symbolic states predicate

It's straightforward by the semantics of the step relations to prove that both normal and attacker steps preserve each of the invariants. We only need to assume that this holds for monitor services (i.e., if we were to provide some monitor services they would have to preserve these invariants).

Lemma 4.24 (Symbolic Invariants preserved by normal steps)**.** *For all symbolic states (st, st'),*

$$\begin{aligned} &invariants\ st \implies \\ &st \to_n st' \implies \\ &invariants\ st' \end{aligned}$$

Lemma 4.25 (Symbolic Invariants preserved by attacker steps)**.** *For all symbolic states (st, st'),*

$$\begin{aligned} &invariants\ st \implies \\ &st \to_a^S st' \implies \\ &invariants\ st' \end{aligned}$$

4.4.5 Stopping predicate for the Symbolic Machine

Similar to the abstract machine, we say that an execution trace of the symbolic machine is stopping if:

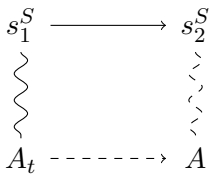
1. All states in the trace are stuck with respect to normal steps ($\to_n$)
2. All steps in the trace are attacker steps ($\to_a$)

4.4.6 Symbolic-Abstract backward simulation

The Symbolic-Abstract backward simulation formally defines the connection between the two machines. We prove a 1-backward simulation theorem for both normal and attacker steps. This means that every step of the symbolic machine is matched by one step of the abstract machine.

Definition 4.26 (1-Backward Simulation)**.** *The symbolic machine simulates the abstract machine with respect to a simulation relation* $\sim$ *between symbolic and abstract states, if* $s_1^A \sim s_1^S$ *and* $s_1^S \rightarrow_n s_2^S$ *implies that there exists* s_2^A *such that,* $s_2^A \sim s_2^S$ *and* $s_1^A \rightarrow_n s_2^A$*.*

We visualize the above definition with the following diagram:



(Plain lines denote premises, dashed ones conclusions.)

Intuitively, backward simulation is enough to capture the desired security property. Our intuition is further strengthened later, when we prove that the *CFI* property given by definition 4.7 is preserved by backward refinement. However, a machine that cannot take any step also enjoys *CFI* vacuously. Forward simulation would guarantee us that if we start from two states in relation ($s_1^A \sim s_1^S$), then a step in the abstract machine ($s_1^A \rightarrow_n s_2^A$) is matched by a step in the symbolic machine ($s_1^S \rightarrow_n s_2^S$) and the resulting states are related by the simulation relation ($s_2^A \sim s_2^S$). We have not proved forward simulation but we believe it holds.

(**TODO:** Try to change that)

Simulation Relation

We define the state simulation relation between the symbolic and abstract machine by defining the simulation relation for each component of the state.

Definition 4.27 (Data Memory Simulation)**.** *An abstract data memory dm is in simulation with a symbolic memory mem, if for all words w, x it holds that*

$$mem[w] = x@Data \iff dm[w] = x$$

Definition 4.28 (Instruction Memory Simulation)**.** *An abstract instruction memory im is in simulation with a symbolic memory mem, if for all words w, x it holds that*

$$(\exists\ it,\ mem[w] = x@(Instr\ it)) \iff im[w] = x$$

Definition 4.29 (Registers Simulation)**.** *An abstract register set areg is in simulation with a symbolic register set sreg, if for all registers r and words x it holds that*

$$(\exists ut \in (id \cup \{\bot\}),\ sreg[r] = x@ut) \iff areg[r] = x$$

Definition 4.30 (PC simulation)**.** *The abstract pc (apc) is in simulation with the symbolic pc (spc@*t_{pc}*), if it holds that*

$$apc = spc \wedge (t_{pc} = Data \vee \exists n \in id,\ t_{pc} = Instr\ n)$$

Definitions 4.27 to 4.30 relate the basic components of the state. What is left to do, is relate the *ok* bit of the abstract machine with the state of the symbolic machine.

Definition 4.31 (Correctness)**.** *The statement of correctness, states that for the symbolic memory (smem), the symbolic pc (spc@t_{pc}) and the ok bit of the abstract machine, it holds that for all words i and tags ti,*

$$
\begin{aligned}
&smem[spc] = i@ti \implies \\
&ok = true \iff \\
&(\forall src \in id,\ t_{pc} = Instr\ src \implies \\
&\quad \exists dst \in id, \\
&\qquad ti = Instr\ dst \wedge (src, dst) \in \mathcal{CFG})
\end{aligned}
$$

Informally definition 4.31 states that if the tag on the current instruction is *ti*, then if the tag on the pc is set to *Instr* src (which means an indirect flow occurred in the previous step), there exists an *id* dst which is used to tag the current instruction and additionally the flow from an instruction with *id* src to one with *id* dst is allowed according to $\mathcal{CFG}$, if and only if the *ok* bit of the abstract machine is set to true. This definition captures the notion that a violation in the abstract machine is also a violation in the symbolic machine and vice-versa.

We give one more definition of correctness, for the case of monitor services. The intuition is the same, but because monitor services live outside the addressable memory of the machines, it's statement needs to be adapted a bit.

Definition 4.32 (Monitor Service Correctness)**.** *Correctness for monitor services, states that for the symbolic memory (smem), the symbolic pc (spc@t_{pc}) and the ok bit of the abstract machine, it holds that for all monitor services sc,*

$$
\begin{aligned}
&smem[spc] = \varnothing \implies \\
&get\_service\ spc = (ti, f) \implies \\
&ok = true \iff \\
&(\forall src \in id,\ t_{pc} = Instr\ src \implies \\
&\quad \exists dst \in id, \\
&\qquad ti = Instr\ dst \wedge (src, dst) \in \mathcal{CFG})
\end{aligned}
$$

The simulation relation ($\sim$) is defined as the conjunction of definitions 4.27 to 4.32 and the invariants 4.20 to 4.22.

Proving 1-backward simulation for normal steps

Proving a 1-backward simulation for normal steps is relatively straight-forward, mostly thanks to the fact that the symbolic machine abstracts away many details of the concrete machine that would make the proofs more tedious. Additionally we do not have to provide such proofs for any monitor service as we did not use any. Therefore we will only have to reason about the small set of instructions that the symbolic and the abstract machine share.

We start with some helpful lemmas about registers and memory updates. These lemmas serve as the basis for proving simulation for instructions that change the registers or the memory. The corresponding Coq definitions and proofs can be found. (**TODO:** cite appendix)

Lemma 4.33 (Registers Update Backward Simulation)**.** *For all symbolic register sets (sreg,sreg'), abstract register sets (areg), registers (r), words (v,v') and tags (tg,tg'),*

$$\begin{aligned}
&areg \sim_{regs} sreg \implies \\
&sreg[r] = v@tg \implies \\
&sreg[r \leftarrow v'@tg'] = sreg' \implies \\
&\quad \exists areg', \\
&\qquad areg[r \leftarrow v'] = areg' \wedge \\
&\qquad areg' \sim_{regs} sreg'
\end{aligned}$$

Lemma 4.34 (Memory Update Backward Simulation)**.** *For all symbolic memories (smem,smem'), abstract data memories (amem) and words (addr,v,v'),*

$$\begin{aligned}
&amem \sim_{dmem} smem \implies \\
&smem[addr] = v@Data \implies \\
&smem[addr \leftarrow v'@Data] = smem' \implies \\
&\quad \exists amem', \\
&\qquad amem[addr \leftarrow v'] = amem' \wedge \\
&\qquad amem' \sim_{dmem} smem'
\end{aligned}$$

With these definitions and lemmas we are able to prove 1-backward simulation for normal steps between the Symbolic and the Abstract machine as defined by definition 4.26, a crucial theorem for the rest of our development.

Proving 1-backward simulation for attacker steps

The same definition as 4.26 of 1-backward simulation is used for the attacker, with the sole difference being that steps now refer to attacker steps.

Definition 4.35 (1-Backward Simulation Attacker)**.** *The symbolic machine simulates the abstract machine with respect to a simulation relation $\sim$ between symbolic and abstract states, if $s_1^A \sim s_1^S$ and $s_1^S \rightarrow_a^S s_2^S$ implies that there exists s_2^A such that, $s_2^A \sim s_2^S$ and $s_1^A \rightarrow_a^S s_2^A$.*

We prove that 1-backward simulation for attacker steps hold, by first showing how we can construct attacker steps at the abstract level from symbolic attacker steps and then showing that this way of building attacker steps preserves the simulation relation ($\sim$).

A step of the symbolic attacker, as mandated by the semantics of the attacker model, can only change the memory and register contents tagged *Data*, formally $mem \rightarrow_a^S mem'$ and $reg \rightarrow_a^S reg'$.

Intuitively, we can construct *areg* by *mapping* a function on the set of registers, that changes a symbolic atom to a word by removing it's tag.

```
Definition untag_atom (a : atom (word t) cfi_tag) := common.val a.
```

Listing 4.36: Untag symbolic atom function

We can trivially prove that the abstract attacker can take a step by *mapping* untag_atom over a symbolic register set. This is trivial because the attacker can arbitrarily change all registers.

Lemma 4.37 (Abstract attacker registers)**.**

$$sreg \rightarrow_a^S sreg' \implies$$
$$areg \rightarrow_a^A map\ untag\_atom\ sreg'$$

However, we still need to prove that the simulation relation between the two machines does not break when attacker steps are taken. We can proof that simulation of registers is preserved by attacker steps. The proof proceeds by using the correctness theorem for the map function.

Theorem 4.38 (Map Correctness instance)**.**

$$(map\ untag\_atom\ sreg')[r] = option\_map\ untag\_atom\ (sreg'[r])$$

where option_map is defined as

```
Definition option_map f x :=
  match x with
  | Some y ⇒ Some (f y)
  | None ⇒ None
  end.
```

Listing 4.39: Option Map function

Lemma 4.40 (Attacker preserves register simulation)**.** *For all abstract register sets (areg) and symbolic register sets (sreg, sreg'),*

$$areg \sim_{regs} reg \implies$$
$$sreg \rightarrow_a^S sreg' \implies$$
$$map\ untag\_atom\ sreg' \sim_{regs} sreg'$$

In order to complete the proof of 1-backward simulation for attacker steps, we also need to construct an abstract memory and to show that the $\sim_{mem}$ relation is preserved by attacker steps. Due to the fact that the abstract machine has split data and instruction memories, in order to follow the same methodology as with registers, we will need to split the symbolic memory. We achieve this, using a filter function.

Firstly we proof that attacker steps do not break simulation of instruction memories. Intuitively this is trivial, as the symbolic attacker can only change memory contents tagged *Data*.

Lemma 4.41 (Attacker preserves instruction memory simulation)**.** *For all abstract instruction memories (imem) and symbolic memories (smem, smem'),*

$$imem \sim_{imem} smem \implies$$
$$smem \rightarrow_a^S smem' \implies$$
$$imem \sim_{imem} smem'$$

Constructing a data memory is more complicated than in the previous cases. Our approach, uses the filter function to create a subset of the symbolic memory that only contains atoms tagged *Data* and then applies the same methodology with registers, mapping the *untag_atom* function over this subset to obtain an abstract data memory.

Again we can prove a few helpful lemmas that ease the final proof.

```
Definition is_data (a : atom (word t) cfi_tag) :=
  match common.tag a with
    | DATA ⇒ true
    | INSTR _ ⇒ false
  end.
```

Listing 4.42: Function that checks if atom is tagged *Data*

Lemma 4.43 (Attacker preserves data memory simulation)**.** *For all abstract data memories (dmem) and symbolic memories (smem, smem'),*

$$\begin{aligned} &dmem \sim_{dmem} smem \implies \\ &smem \rightarrow_a^S smem' \implies \\ &map\ untag\_atom\ (filter\ is\_data\ sreg' \sim_{dmem} dmem' \end{aligned}$$

The proof of lemma 4.43 is slightly more complex than the one for registers, as we now have to invoke the filter correctness theorem as well.

Theorem 4.44 (Filter Correctness instance)**.**

$$(filter\ is\_data\ smem')[addr] = option\_filter\ is\_data\ (smem'[addr])$$

where option_map is defined as

```
Definition option_filter f x :=
  match x with
  | Some x0 ⇒ if f x0 then Some x0 else None
  | None ⇒ None
  end.
```

Listing 4.45: Option Filter function

In all cases, we have to show that the domains of the abstract memories and registers are also preserved. We include here the corresponding lemma for the data memory. It's proof was again more complicated, due to the fact that we had to split the symbolic memory.

Lemma 4.46 (Attacker preserves data memory domains)**.** *For all abstract data memories (dmem, dmem') and symbolic memories (smem, smem'),*

$$\begin{aligned} &dmem \sim_{dmem} smem \implies \\ &smem \rightarrow_a^S smem' \implies \\ &dmem' \sim_{dmem} smem' \implies \\ &\mathcal{D}(dmem) = \mathcal{D}(dmem') \end{aligned}$$

Finally using these lemmas we can prove that, a 1-backward simulation for attacker steps as defined by definition 4.35 holds.

4.5 The Concrete Machine

Assuming the existence of correct code that implements the *CFI* monitor, we can utilize the framework of section 2.3 to instantiate the concrete machine and obtain a refinement between the concrete and the symbolic machines, we need to provide the encoding of symbolic tags. For the concrete machine we only considered a 32-bit architecture, but as already mentioned, we could very easily instantiate the concrete machine with 64-bit words with minimal changes to our proofs.

4.5.1 Concrete tags

In order to obtain the concrete tags, we need to wrap the symbolic tags with the monitor self-protection tags (*User*, *Entry*, *Monitor*) and provide an encoding to words of these tags.

We instantiate the *id* type of cfi_id class (listing 4.1) as bit-fields of size 28-bits. That means, that we can uniquely identify up to 2^{28} instructions. Trying to tag more instructions than this, would break the symbolic invariant 4.20, because by the simulation relation between the concrete and symbolic machines, the two machines follow the same tagging scheme for *User* and *Entry* tags.

Defining the conversion functions [1] between *words* and *ids* is straight forward. We make the simply choice, to convert *words* to *ids* only if they are equal or less than the maximum word our 28-bit *ids* can fit. Note that this does not mean we reduce the addressable space to 28-bits. You can use addresses higher than 2^{28} to place contents tagged as *Data* or *Monitor* or even *Instr* $\perp$ but not instructions with an identifier.

The conversion from *ids* to *words* is trivial by expanding the id to 32-bit words by adding zeros to the high bits.

(**Question:** To Catalin: Do the above make sense to you?)

```
Definition  id_size  := Word.int 27.
Definition  id := [ eqType  of  id_size ].
Definition  bound  :  word  t :=
  Word.repr  (( Word.max_unsigned  27)  + 1)%Z.  (*29 bits*)

Definition  word_to_id  (w : word t) : option id_size :=
  if (Word.ltu w bound) then Some (Word.castu w) else None.

Definition  id_to_word  (x : id) : word t :=
  Word.castu x.
```

Listing 4.47: Coq definitions of conversion functions for ids and words

Finally we prove the two properties required by cfi_id, *id_to_wordK* and *word_to_idK*. We can now instantiate cfi_id with 28-bit sized *ids*.

When using identifiers of 28-bits, we can encode the symbolic tags using 30 bits, with an encoding like the one in table 4.48, where the two least-significant bits are used to distinguish between *Data*, *Instr* $\perp$ and *Instr id*, and the 28 higher-bits are the *id* in the last case and zero otherwise.

[1] Numbers in the Coq definitions are off by one (e.g., 27 means 28), for reasons relating to the underlying words library (**TODO:** cite library)

```
Instance  ids : cfi_id := {|
 id := id;
 word_to_id  := word_to_id ;
 id_to_word  := id_to_word
|}.
Proof .
  - by apply  id_to_wordK .
  - by apply  word_to_idK .
Defined .
```

Listing 4.48: cfi_id instance with 28-bit sized *ids*

Symbolic Tag	Encoding
Data	0
Instr $\perp$	1
Instr id	$4x + 2$

Table 4.48: Encoding of Symbolic Tags

Having an encoding into 30-bits of symbolic tags, we can use the 2-bits left, to wrap the symbolic tags with the monitor self-protection tags. We use the two least-significant bits to distinguish between *User* (01), *Entry* (10) and *Monitor* (00). Only the *User* and *Entry* wrap around symbolic tags. The policy monitor does not use symbolic tags and the corresponding tag *Monitor* does not need to wrap around them. Thus the encoding of the *Monitor* tag has all its bits set to zero.

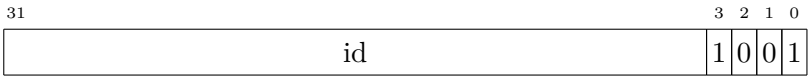


Figure 4.49: Encoding of an instruction with a unique identifier id

Chapter 5

Conclusion

5.1 Future Work

There are many directions still left to explore before we can consider our work done. Some of them include writing the *CFI* monitor code and verifying it, increasing precision by enforcing call-stack protection, scaling to more complex architectures, looking for ways to enforce *CFI* like properties on self modifying programs and more.

5.1.1 Writing and Verifying Monitor Code

5.1.2 Call-Stack Protection/ XFI

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Appendix A

Stuff

A.1 Control-Flow Integrity Micro-Policy

We begin with a micro-policy targeting control-flow hijacking attacks, in which an attacker exploits a low-level vulnerability (e.g. a buffer or integer overflow) to gain full control of a target program [**?**, 20, 3, **?**, **?**, **?**, **?**, **?**, **?**]. As a first line of defense, we can use tags to make code non-writable (NWC) and data non-executable (NXD), preventing the injection and execution of an attacker payload. This useful defense appears in various forms in existing systems. However, it does not prevent code-reuse attacks [16, 19, 14, 18, 6, 5, **?**, 13] such as return- or jump-oriented programming [18, 6], where the attacker chains together existing code snippets ("gadgets") to induce arbitrary malicious behavior. We therefore use tags to enforce fine-grained *control-flow integrity (CFI)* [2, 22, 21, 7, 17, 22, 7] on top of basic NWC and NXD protection. This ensures that all indirect control flows (computed jumps) adhere to a fixed control flow graph (CFG).

We use tags to distinguish between code and data. Tags on memory and the PC are drawn from the set *Data* | *Code addr* | *Code* $\bot$ (registers are always tagged *Data*). To simplify the CFG conformance checks, instructions that are the source or target of indirect control flows are tagged with *Code addr*, where *addr* is the address of that instruction in memory. For example, a *Jump* instruction stored at address 500 is tagged *Code* 500. All other instructions are tagged *Code* $\bot$. *(AAA: Actually, we can't use the instruction's address on the tag if we are to have the same number of bits on words and tags. Maybe change to "id"?)*

We write transfer functions as a collection of *symbolic rules* [9, 11]. (The PUMP hardware uses a lower-level *concrete rule* format, described in **??**.) Each symbolic rule has the form "$opcode : (PC, CI, OP_1, OP_2, OP_3) \to (PC', R')$," which says that the rule matches on the given *opcode* together with the metadata tags on the program counter (PC), the current instruction (CI), and on up to three operands (OP_1 to OP_3). If the rule applies, the right-hand side determines how to update the tags on the PC (PC') and on the result of the operation (R'). We write "$-$" to indicate input or output fields that are ignored ("wildcard"). All instructions that are not explicitly allowed by the symbolic rules are disallowed.*(AAA: We should choose only one of $-$ or _ for our wildcard and use it consistently (cf. the "Store" rule below))*

The CFI transfer function enforces that only memory locations tagged *Data* can be modified (NWC) and only instructions fetched from locations tagged *Code* can be executed (NXD). The symbolic rule for the *Store* instruction illustrates both these points:

$$Store : (Data, Code\ \_, -, -, Data) \to (Data, -)$$

It requires the fetched *Store* instruction to be tagged *Code* and the written location to be tagged *Data*. This rule only applies when the PC is also tagged *Data*, which is the case when the *Store* instruction was reached by direct control flow (not a computed jump). The rule preserves the *Data* tag on the PC, since *Store* is not a computed jump. Performing a computed jump (e.g., using *Jal*, a jump-and-link instruction) requires that the current instruction be tagged *Code src* for some address *src*.

$$Jal : (Data, Code\ src, -, -, -) \to (Code\ src, -)$$

This rule copies *Code src* to the PC tag to indicate that a jump from *src* has just occurred. Only on the next instruction do we have enough information about the destination in the tags to check that the jump is indeed allowed by the CFG. For this we add a second rule for *Store*, dealing with the case where it is the target of a jump and thus the PC is tagged *Code src*.

$$\frac{(src, tgt) \in CFG}{Store : (Code\ src, Code\ tgt, -, -, Data) \to (Data, -)}$$

*(AAA: Maybe we could discuss here a little bit why we verify the jump on the next instruction, as opposed to when the jump is performed. This might get some people confused, since this is not very natural and fundamentally driven by our current design of the PUMP. Even Nick wanted to know if we couldn't do it differently.)*The premise of this rule ensures that the source and target of the just-performed jump are allowed by the CFG. We add a similar rule for each instruction, including jumps (since the target of a computed jump can itself be another computed jump):

$$\frac{(src, tgt\text{-}src) \in CFG}{Jal : (Code\ src, Code\ tgt\text{-}src, -, -, -) \to (Code\ tgt\text{-}src, -)}$$

This micro-policy enforces fine-grained CFI [17, 13, 7], not coarse-grained approximations [2, 21] that are potentially vulnerable to attack [13]. Indeed, we recently proved [9] that this micro-policy enforces a variant of the CFI property introduced by Abadi *et al.* [2], ensuring that all indirect control flows adhere to a fixed CFG. Recent simulations of an optimized PUMP architecture [10] show that the CFI policy can be enforced with around 3% average runtime overhead.