Abstract—Compartmentalization is widely regarded as good security-engineering practice: if we break up a large software system into mutually distrustful components that run with minimal privileges, restricting their interactions to conform to well-defined interfaces, we can limit the damage caused by low-level attacks such as control-flow hijacking. But the formal guarantees provided by such low-level compartmentalization have been surprisingly little investigation.

We propose a new property, secure compartmentalization, that formally characterizes the security guarantees provided by low-level compartmentalization and clarifies its attacker model. We rationally reconstruct the secure compartmentalization property starting from the well-established notion of fully abstract compilation, by identifying and lifting three important limitations that make standard full abstraction unsuitable for compartmentalization. The connection to full abstraction allows us to prove secure compartmentalization for language implementations by adapting established proof techniques; we illustrate this for a simple unsafe imperative language with procedures and a compiler from this language to a compartmentalized abstract machine.

I. INTRODUCTION

Today’s computer systems are distressingly insecure. Visiting a website, opening an email, or serving a client request is often enough to cause a computer to be compromised by a control-hijacking attack. These devastating low-level attacks typically exploit memory-safety vulnerabilities such as buffer overflows, use after frees, or double frees, which are abundant in large software systems. While many techniques have been proposed for guaranteeing memory safety [13], [18], [21], [38], [45]–[49], the challenges of efficiency [48], [49], precision [63], scalability [66], backwards compatibility [16], and effective deployment [13], [18], [20], [21], [38], [45]–[47] have hampered their widespread adoption.

Meanwhile, new mitigation techniques are proposed to deal with the most onerous consequences of the lack of memory safety—for instance, techniques for attempting to prevent control-flow hijacking even in memory-unsafe settings [2], [3], [22]. Unfortunately, these defenses often underestimate the power of the attackers they will face [17], [22], [24], [25], [27], [59]—if, indeed, they have any clear model at all of what they are supposed to protect against. Formally studying and clarifying the precise security properties and attacker models of practical mitigation techniques is thus an important research problem—and a challenging one, since a good model has to capture not only the defense mechanism itself but also the complex world in which low-level attacks occur.

In this paper we focus on low-level compartmentalization [14], [28], [62], a class of strong, practical defense mechanisms whose security has seen surprisingly little formal investigation. The idea is to break up a large software system into mutually distrustful components that run with minimal privileges and can interact only via well-defined interfaces. This provides not only software engineering gains, but also strong security benefits; in particular, control hijacking attacks can compromise only specific components with exploitable vulnerabilities and thus only give the attacker direct control over the privileges held by these components. Moreover, because compartmentalization is a coarse-grained enforcement mechanism, efficiency and backwards compatibility issues are generally smaller than for full-blown memory safety. Indeed, two different compartmentalization technologies are already widely deployed: process-level privilege separation [14], [28], [34] (e.g., used by OpenSSH [55] and for sandboxing plugins and tabs in modern web browsers [56]) and software-fault isolation [61] (e.g., provided by Google Native Client [64]); and many more such technologies are on the drawing boards [13], [29], [54], [62].

So what security guarantees does compartmentalization provide, and what, exactly, is its attacker model? A good starting point for addressing these questions is the familiar notion of fully abstract compilation [1], [4]–[6], [8], [10]–[12], [26], [30], [51]. A fully abstract compiler protects the interactions between a compiled program and its low-level environment, allowing the programmer to reason (soundly) about the behavior of their code when placed in an arbitrary target-language context by considering only its behavior in arbitrary source-language contexts. In particular, if we link the code produced by such a compiler against arbitrary low-level libraries—perhaps written in an unsafe language or even directly in assembly—the resulting executable will not be any less secure than if we had restricted ourselves to library code written in the same high-level language as the calling program. (Why is it useful to restrict attention to attackers written in a high-level language? First, because reasoning about what attackers might do—in particular, what privileges they might exert—is easier in a high-level language. And second, because by phrasing the property in terms of low-level and high-level programs rather than directly in terms of attacker behaviors, specific notions of privilege, etc., we can re-use the same definition for many specific languages.)

Since full abstraction works by partitioning the world into a program and its context, we might expect it to make sense for compartmentalized programs as well: Compartments that are assumed to be subject to control-hijacking attacks would
be grouped into the “low-level context,” while those that are immune to such attacks would constitute the “high-level program.” Full abstraction would then allow us to reason about the possible behaviors of the whole system using the simplifying assumption that the attacker’s injected behavior for the compromised components can be expressed in the same high-level language as the good components. Unfortunately, this intuition does not withstand closer examination. Fully abstract compilation, as previously formulated in the literature, suffers from three important limitations that make it unsuitable for characterizing the security guarantees of low-level compartmentalization.

First, fully abstract compilation assumes that our source language is itself secure, so that it makes sense to define target-level security with respect to the semantics of the source language. However, compartmentalization is often applied to languages like C and C++, which do not have a secure semantics—the C and C++ standards leave most of the security burden to the programmer by calling out a large number of undefined behaviors, including memory-safety violations. Programs in these languages are allowed to do anything—in particular, to do whatever a remote attacker wants—when applied to inputs that lead to undefined behavior. There is no way to tell, statically, whether or not a program may have undefined behavior, and compilers do not check for this situation. (Indeed, not only do they not check: they aggressively exploit the assumption of no undefined behaviors to produce the fastest possible code for well-defined programs, often leading to quite bizarre and easily exploitable behaviors when this assumption is broken.) The point of compartmentalization is to ensure that the effects of undefined behavior are restricted to compromise of the component in which it occurs, and that the other components can only be influenced by compromised components via specified interactions respecting the existing interfaces. To formally characterize the security of low-level compartmentalization we thus need a property that can meaningfully accommodate source languages in which components can be compromised via undefined behavior. Full abstraction as conventionally formulated does not fit the bill, because it requires that all nondeterminism of the source language—including nondeterminism due to undefined behaviors—be preserved by the compilation process. While targetting machine code, no realistic implementation of an unsafe language like C is going to preserve all of this nondeterminism—and removing it would negate the performance and optimization benefits that are the reason for allowing undefined behaviors in the first place—so no realistic C implementation has any hope of being fully abstract.

Second, fully abstract compilation makes an open world assumption about the attacker context. While the context is normally required to be compatible with the protected program, for instance by respecting the program’s typed interface, the structure and privilege of the context are unrestricted. This comes in direct contradiction with the idea of least privilege, which is crucial to compartmentalization, and which relies on the fact that even if a component is compromised, it does not immediately get more privilege. In our static compartmentalization setting the breakup of the application into components is fixed and so are the privileges of all the components. Compromised components cannot change these basic rules of the compartmentalization game and have to play by them. (We consider here the case where the set of compartments is fixed. In a setting where new components can be dynamically created, the details of the game will probably get more complicated; still, we expect a secure compartmentalization property to make a closed world assumption about the compromise of existing components.)

Third, because the definition of full abstraction involves applying the compiler only to the trusted “program” and not to the untrusted context in which it runs, fully abstract compilers can achieve their protection goals by introducing one single barrier around the trusted part to protect it from the untrusted part [8], [39], [50], [51], [53]. Such compilation schemes force the programmer to commit in advance to a single compromise scenario, i.e., to a single static split of their application into a trusted program and an untrusted context from which this program has to be protected. This is not realistic in the setting of low-level compartmentalization, where we generally cannot predict which components may be vulnerable to compromise by control hijacking attacks and where we therefore need to consider multiple compromise scenarios. Fine-grained compartmentalization allows us to build more secure applications that go beyond the blunt trusted/untrusted distinction made by some fully abstract compilers. To describe its guarantees accurately, we need a new property that captures the protection obtained by breaking up applications into multiple mutually distrusting running with least privilege, and that allows us to reason about all compromise scenarios.

Our main contribution is to define such a property, which we call secure compartmentalization (§III). While similar in many respects to full abstraction, secure compartmentalization overcomes the three limitations above. First, secure compartmentalization applies to unsafe source languages with undefined behaviors by introducing a new notion of fully defined sets of components. While undefined behavior is a property of whole programs, full definedness is compositional. Intuitively, a set of components is fully defined if they cannot be blamed for undefined behavior in any context satisfying fixed interfaces. Second, secure compartmentalization makes a closed-world assumption about compromised components, enforcing the basic rules of the compartmentalization game like the fixed breakup into components and the fixed privilege level for each component. Third, secure compartmentalization ensures protection for multiple, mutually distrusting components; it does not assume we know in advance which components are going to be compromised by undefined behavior, but instead it explicitly quantifies over all possible compromise scenarios.

Our second contribution is relating secure compartmentalization to more standard formulations of full abstraction both intuitively and formally (§III). We start from full abstraction and illustrate how the three limitations that make it unsuitable in our setting can be lifted one by one. This results in two
properties we call structured full abstraction and separate compilation, which can be put together and instantiated to obtain secure compartmentalization. While secure compartmentalization directly captures the intuition of our attacker model, reducing it to structured full abstraction is a useful technical step, since the latter is easier to establish for specific examples using a variant of existing proof techniques. Moreover, arriving at the same property by two different paths increases our confidence that we found the right property.

Our third contribution is proving secure compartmentalization for a simple unsafe imperative language with procedures compiling to an abstract machine with protected interacting components (§IV). The proof adapts a standard technique called trace semantics [32], [52], via the reduction to structured full abstraction. In the future, we expect that our secure compartmentalization property can be used to formally assess the security of existing compartmentalization mechanisms [13], [28], [29], [34], [54]–[56], [64]. However, this is nontrivial because of the extreme complexity of all the involved artifacts (e.g., efficient compilers for C or C++ targeting a realistic hardware platform). Our simple secure compartmentalization proof constitutes a first step in this direction and a sanity check for our property.

We describe each of our three contributions in detail (§II–§IV), and close by discussing related work (§V) and future directions (§VI). The supplemental material submitted with this paper includes: (a) a Coq proof for Theorem III.4; (b) complete technical details for the secure compartmentalization instance from §IV; and (c) a trace mapping algorithm in OCaml supporting Assumption IV.9.

II. Secure Compartmentalization

In this section we give an intuitive explanation of low-level compartmentalization, its attacker model, and its security benefits, and we introduce our proposed secure compartmentalization property.

We consider low-level compartmentalization mechanisms provided by the compiler and runtime system for an unsafe programming language with some notion of components. (We use the term “runtime system” loosely to include operating system mechanisms [14], [28], [34], [55], [56] and/or hardware protections [13], [29], [54], [62] that may be used by the compiler.) In §IV we present a simple example in detail, but for the present discussion it suffices to think informally of C or C++ enriched with some compartmentalization mechanism (e.g., a mechanism for dividing an application into components during its initialization phase and a runtime enforcement mechanism for ensuring that these component boundaries are respected during the rest of its execution). Security-conscious developers can break up large applications written in such a language into mutually distrustful components running with least privilege and interacting only via well-defined interfaces. We assume that the interface of each component also gives a precise description of its privilege. Our notion of interface here is quite generic: interfaces might include any information that can be dynamically enforced on components, including module signatures, lists of allowed system calls, or more detailed access control specifications describing legal parameters to inter-compartment calls (e.g., ACLs for files). In §IV, we instantiate this picture with a rather simple and rigid notion of compartments and interfaces, where compartments don’t directly share any state and where the only thing one compartment can do to another one is to call the procedures allowed by the interfaces of both components. We also assume that the division of the application into components and the interfaces of those components are statically determined and fixed throughout execution.

We do not fix a specific compartmentalization mechanism; we just assume that whatever mechanism is chosen is able to guarantee that, even if a component is compromised by a control-hijacking attack, it will still be forced to adhere to its specified interface in its interactions with other components. What a compromised component can do in this model is use its access to other components, as allowed by its interface, to trick them into misusing their own privileges (confused deputy attacks) and/or attempt to mount further control-hijacking attacks to compromise more components.

We do not assume we know in advance which components will be compromised: the compartmentalization mechanism has to protect each component from all the others. This allows developers to reason informally about various compromise scenarios and their impact on the security of the whole application [28], relying on conditional reasoning of the form: “If these components get taken over and these do not, then this might happen (while that cannot), whereas if these other components get taken over, then this other thing might happen…”

If the practical consequences of some plausible compromise scenario are too serious, developers can reduce or separate privilege by narrowing interfaces or splitting components into smaller pieces. They can also make components more defensive by dynamically validating the inputs they receive from other potentially compromised components.

For instance, developers of a compartmentalized web browser [56] might reason about situations in which some subset of plugins and tabs gets compromised and how this might impact the browser kernel and the remaining plugins and tabs. One outcome of this exercise might be the observation that, if the browser kernel itself is compromised, then all bets are off for all the components and the application as a whole, so the developers should put extra energy in defending the kernel against attacks from compromised plugins or tabs. On the other hand, if interfaces between tabs and plugins are appropriately limited, then compromise of one might not disrupt another.

Our goal is to articulate a secure compartmentalization property that supports this sort of reasoning about multiple compromise scenarios and that clarifies the attacker model behind low-level compartmentalization. At the same time, secure compartmentalization may be useful as a target property for developers of compartmentalization mechanisms who want to argue formally that their mechanisms are secure. In the rest of this section we first explain the main idea behind secure
An application is a set $Cs$ of components, with corresponding interfaces $Cs$s. These components are separately compiled (written $Cs\downarrow$) and linked together (written $\bowtie(Cs\downarrow)$) to form an executable binary for the application. The interface information is carried along through compilation, linking, and loading, for use at runtime by the dynamic part of the compartmentalization mechanism.

Secure compartmentalization quantifies over all compromise scenarios—i.e., over all ways of partitioning the components into a set of compromised ones and a set of uncompromised ones. In order to ensure that the set of compromised components doesn’t expand during evaluation, we require that the uncompromised components be fully defined with respect to the interfaces of the compromised components. That is, the uncompromised components must not perform undefined behaviors even if we replace the compromised components with any code that obeys the original interfaces. (Intuitively, if an uncompromised component can be tricked into an undefined behavior by interface-respecting communication with other components, then we need to conservatively assume that the already compromised components will succeed in compromising this component dynamically, so it belongs in the set of compromised components from the start.) Figure 1 illustrates one way to partition five components $C_1, \ldots, C_5$ with interfaces $i_1, \ldots, i_5$, representing the scenario when $C_2$, $C_4$, and $C_5$ are compromised and $C_1$ and $C_3$ are not. In order for this compromise scenario to be accepted, $C_1$ and $C_3$ need to be fully defined with respect to interfaces $i_2$, $i_4$, and $i_5$, which means $C_1$ and $C_3$ cannot cause undefined behaviors when linked with any components $B_2, B_4, B_5$ satisfying interfaces $i_2, i_4, i_5$. Formally, full definedness is defined as follows:

**Definition II.1 (Full definedness).**

We call a set of components $Cs$ fully defined with respect to interfaces $Bs$s if for all components $Bs$ satisfying $Bs$s, $Cs$ cannot cause undefined behavior alongside $Bs$.

The relation defining when $Cs$ cannot cause undefined behavior alongside $Bs$ is a language-specific parameter to our definition of secure compartmentalization. For instance, in the simple imperative language in §IV, we say that $Cs$ cannot cause undefined behavior alongside $Bs$ if the program $\bowtie(Cs\downarrow \cup Bs \downarrow)$ does not reduce to a stuck non-final state (corresponding to undefined behavior) in which the currently executing component is one of the ones in $Cs$ (i.e., no component in $Cs$ can be “blamed” [60] for undefined behavior).

Secure compartmentalization states that, in all such compromise scenarios, the compiled compromised components must not cause more harm to the compiled uncompromised components via low-level attacks than some high-level components already could in the source language. Basically this ensures that any low-level attack can be mapped back to a high-level attack by compromised components satisfying the given interfaces. The property additionally ensures that the high-level components produced by this “mapping back” are fully defined with respect to the interfaces of the uncompromised components. So with secure compartmentalization, instead of having to reason about the low-level consequences of undefined behavior in the compromised components, we can reason in the source language and simply replace the compromised.....
components by equivalent ones that are guaranteed to cause no undefined behavior.

Formally, secure compartmentalization is stated by quantifying over multiple distinguishability games, one for each compromise scenario, where the individual games are reminiscent of full abstraction. The goal of the attacker in this game is to distinguish between two variants of the uncompromised components. Figure 2 illustrates these two variants as $C_1, C_3$ and $D_1, D_3$, where we use $\not\sim_H$ and $\not\sim_L$ to indicate that the behaviors of two (high- or low-level) complete programs are distinguishable, i.e., they produce different observable outcomes when executed. For this compromise scenario, secure compartmentalization specifies that, if compiled compromised components $C_2\downarrow, C_4\downarrow, C_5\downarrow$ can distinguish the two variants at the low level, then there must exist some (fully defined) components $A_2, A_4, A_5$ that distinguish the two variants at the high level. With all this in mind, the secure compartmentalization property is formally expressed as follows:

Definition II.2 (Secure Compartmentalization).

- For any complete compartmentalized program and for all ways of partitioning this program into a set of uncompromised components $C$s and their interfaces $CIs$, and a set of compromised components $B$s and their interfaces $BIs$, so that $C$s is fully defined with respect to $BIs$...
- for all ways of replacing the uncompromised components with components $D$s satisfying the same interfaces $CIs$ and fully defined with respect to $BIs$...
- if $\triangleright(C_s \cup B_s) \not\sim_L \triangleright(D_s \cup B_s)$,
- then there exist components $A$s satisfying interfaces $BIs$ and fully defined with respect to $CIs$ such that $\triangleright(C_s \cup A_s) \not\sim_H \triangleright(D_s \cup A_s)$.

III. FROM FULL ABSTRACTION TO SECURE COMPARTMENTALIZATION

§II presented secure compartmentalization by directly characterizing the attacker model against which it defends. In this section we step back and show how it can instead be obtained by starting from the well-established notion of full abstraction and successively removing each of the three limitations that make it unsuitable in our setting. This results in two properties, structured full abstraction and separate compilation, which we then combine and instantiate to obtain secure compartmentalization. Reducing secure compartmentalization to structured full abstraction is not only interesting from a theoretical point of view, but also practically useful, since structured full abstraction can be more easily shown by adapting existing proof techniques, as we will see in §IV.

Full abstraction: A fully abstract compiler protects compiled programs from their interaction with unsafe low-level code and thus allows sound reasoning about security (and other aspects of program behavior) in terms of the source language. Fully abstract compilation [1] intuitively states that no low-level attacker can do more harm to a compiled program than a program in the source language already could. This is a much stronger property than just compiler correctness [37], [40], as it requires enforcing high-level language abstractions against arbitrary low-level attackers, not just against code produced by the same compiler.

Formally, full abstraction is phrased as a distinguishability game where low-level and high-level attackers are specified as having exactly the same distinguishing power.

Definition III.1. We call a compilation function (written $\downarrow$) fully abstract if, for all $P$ and $Q$,

$$\forall A. \ A[P] \sim_H A[Q] \iff \forall a. a[P]_a \sim_L a[Q]_a.$$  

Here, $P$ and $Q$ are partial programs, $A$ is a high-level context whose job is to try to distinguish $P$ from $Q$, and $a$ is a low-level “attacker context” that tries to distinguish $P_a$ from $Q_a$. The relations $\sim_L$ and $\sim_H$ are parameters to the definition, representing behavioral equivalence at the two levels. To be useful, they should allow the context to produce an observable action every time it has control, thus letting it convert its knowledge into observable behaviors. For instance, a common choice for behavioral equivalence is based on termination: two (deterministic) programs are behaviorally equivalent if they both terminate or both diverge.

The right-to-left implication normally follows from compiler correctness by simply using the compiler to translate any high-level distinguishing context. The left-to-right direction is more interesting and difficult to obtain, since it requires some way to map each low-level distinguishing context to a high-level one. To see this, it may help to look at the contrapositive:

$$\forall a. a[P]_a \not\sim_L a[Q]_a \Rightarrow \exists A. A[P] \not\sim_H A[Q].$$

Problem 1: Undefined behavior: The first limitation of full abstraction is that it cannot meaningfully be applied to compiling an unsafe language with undefined behaviors to a deterministic machine. Undefined behaviors are highly nondeterministic, compilers to deterministic machines make choices that resolve this nondeterminism, and this breaks the full abstraction property, which requires that the source and target programs have exactly the same amount of nondeterminism.

To adapt full abstraction to a source language with undefined behaviors, we need to restrict attention only to well-defined complete programs in the source language, which we assume to be deterministic. (For a source language like C, obtaining determinism requires also expanding out any “implementation specific” behavior.) Defining full abstraction still requires a little care, though. For instance, the following variant is wrong:

$$\forall A. A[P] \text{ and } A[Q] \text{ defined } \Rightarrow A[P] \sim_H A[Q] \iff \forall a. a[P]_a \sim_L a[Q]_a.$$  

Any programs $P$ and $Q$ that trigger undefined behavior as soon as they get control would be considered equivalent in the high-level language because there is no context that can make these programs defined while observing some difference between them. All such programs would thus need to be equivalent at the low level, which is clearly not the case. The problem here is that if $P$ and $Q$ trigger undefined behavior then the context
often cannot make up for that and make the program defined in order to cause an observation that distinguishes \( P \) and \( Q \).

**Solution 1: Full abstraction for unsafe languages:**

The responsibility of keeping \( A[P] \) defined should be thus shared between \( A \) and \( P \). For this we need a compositional definition of **fully defined** behavior for programs and contexts. A direct definition fails though, since it would be circular: intuitively, a program is fully defined if in does not cause undefined behavior in any fully defined context, and dually a context is fully defined if it does not cause undefined behavior when we plug any fully defined program into it. As done implicitly in §II, we break this circularity using *blame* [60]: intuitively we call a partial program **fully defined** when it cannot be blamed for undefined behavior in any context. Similarly, we call a context fully defined when it cannot be blamed for undefined behavior for any program that we plug into it. This allows us to define a new variant of full abstraction that applies to unsafe source languages with undefined behavior:

**Definition III.2 (Full abstraction for unsafe languages).**

We call a compiler \( \downarrow \) for an unsafe language **fully abstract** if for all **fully defined** partial programs \( P \) and \( Q \)

\[
(\forall A. \text{ a fully defined } \Rightarrow A[P] \sim_H A[Q]) \iff \left( \forall a. a[P] \sim_L a[Q] \right)
\]

By requiring that \( P, Q, \) and \( A \) are fully defined we can safely apply \( \sim_H \) to \( A[P] \) and \( A[Q] \), because neither the programs nor the context can be blamed for undefined behavior. This property is incomparable with the original definition of full abstraction. If we look at the interesting (left-to-right) direction of the equivalence in contrapositive form

\[
\forall P, Q \text{ fully defined. } (\exists a. a[P] \not\sim_L a[Q]) \Rightarrow (\exists A. \text{ a fully defined } \Rightarrow A[P] \not\sim_H A[Q])
\]

the \( P, Q \text{ fully defined } \) pre-condition makes this weaker than full abstraction, while \( A \text{ fully defined } \) post-condition makes it stronger. The post-condition is valuable, as it allows us to replace reasoning about low-level contexts, with reasoning about high-level contexts that cannot cause undefined behavior.

One might wonder whether the \( P, Q \text{ fully defined } \) pre-condition is too restrictive, since full definedness is a rather strong property, requiring a partial program to be very defensive about validating the inputs it receives. Without additional information about the program’s context though, we can only be conservative and assume that if any context can cause undefined behavior in a program it can compromise it in a way that the compiler can provide no guarantees for this program. The structured full abstraction definition below (and separate compilation in §II) will in fact gather more information about contexts and thus use a weaker notion of full definedness. Moreover, separate compilation will allow us to cover all splits of a program into a fully defined partial program and a compromised context, which also makes the presence of the full definedness pre-condition more palatable.

**Problem 2: Open-world assumption about contexts:**

While full abstraction normally requires the contexts to be compatible with the partial program, for instance by respecting the partial program’s typed interface, these restrictions are minimal and do not restrict the shape, size, exported interface, or privilege of the contexts in any way. This open world assumption about contexts does not work in our compartmentalization setting though, in which the breakup of the application into components is fixed in advance and so are the interfaces (and thus privileges) of all the components. In our setting, the definition of full abstraction needs to be changed to keep track of and respect such structural constraints, otherwise a low-level context with 2 components could be mapped back to a high-level context with 3 components that have different exported interfaces (and thus privileges).

**Solution 2: Structured full abstraction:** For solving this problem we introduce a structured variant of full abstraction, in which partial programs (\( \bullet \)) and contexts (\( \circ \)) are assigned dual parts of pre-defined complete program shapes. A shape can be anything, from a breakup into components with their interface (Theorem III.4), to maybe the maximal size of a component’s code after compilation (exposing size leakage in a setting where it’s too costly to hide component sizes).

**Definition III.3 (Structured full abstraction).**

We say that a compiler \( \downarrow \) for an unsafe language satisfies structured full abstraction if, for all program shapes \( P \) and partially programs \( P \in^* s \) and \( Q \in^* s \) so that \( P \) and \( Q \) are **fully defined** with respect to contexts of shape \( \circ s \),

\[
\left( \forall A \in^0 s. \text{ A fully defined wrt programs of shape } \bullet s \Rightarrow A[P] \sim_H A[Q] \right) \iff \left( \forall a \in^0 s. a[P] \sim_L a[Q] \right)
\]

This property universally quantifies over any complete program shape \( s \) and requires that \( P \in^* s \) (read “program \( P \) has shape \( s' \)”), \( Q \in^* s \), and \( A \in^0 s \) (“context \( A \) matches programs of shape \( s' \)”). Moreover, the property only requires programs that are fully defined with respect to contexts of the right shape, and dually it only considers contexts that are fully defined with respect to programs of the right shape.

**Recovering secure compartmentalization:** Secure compartmentalization can be recovered in a natural way as an instance of structured full abstraction (Definition III.3). For both the source and the target language we take partial programs and contexts be sets of components and context application be set union. Compilation of sets of components works pointwise. To obtain an instance of structured full abstraction we additionally take shapes to be sets of component interfaces, where each interface is marked as either compromised or uncompromised.

**Theorem III.4.** Structured full abstraction instantiated to components as described above implies secure compartmentalization.

**Proof:** Straightforward (though tedious). A machine-checked Coq proof can be found in the auxiliary materials.
Problem 3: Statically known trusted/untrusted split: So we can recover secure compartmentalization as an instance of structured full abstraction. However, while secure compartmentalization can deal with multiple compromise scenarios, not all instances of structured full abstraction can. In general, if for a compiler satisfies (structured) full abstraction, how can we know whether it can deal with multiple compromise scenarios, and what does that even mean? While we can instantiate full abstraction to a particular compromise scenario by letting the partial program \( P \) contain the uncompromised components and the low-level context \( a \) contain the compromised ones, a fully abstract compiler (together with its linker, loader, runtime etc.) might exploit this static split and introduce only one single barrier protecting the uncompromised components from the compromised ones. When presented with a different compromise scenario for the same program, the compiler could adapt and produce a different output.

Formally, a fully abstract compiler does not need to compile contexts—only programs. In fact, even the types of contexts and of partial programs could well be completely different (for instance, the type of lambda calculus contexts and terms are different, and a compiler for lambda calculus terms cannot compile contexts). Even when the types do match and we can apply the same compiler to the context, the low-level context-application operation \( A \downarrow P \downarrow \) can freely exploit the fact that its first argument is a compiled untrusted context and its second argument is a compiled trusted program that should be protected from the context. This means that, if we start with a complete high-level program \( C \) and look at two different compromise scenarios \( C = A_1[P_1] \) and \( C = A_2[P_2] \), compiling each of the parts and combining the results using context application does not necessarily yield the same result (i.e., it could well be that \( A_1(P_1) \neq A_2(P_2) \)) or indeed even behaviorally equivalent results (i.e., it could well be that \( A_1(P_1) \not\sim L A_2(P_2) \)). This means that the user of a fully abstract compiler may need to commit in advance to a single compromise scenario.

This weakness significantly limits the applicability of full abstraction. Indeed, uncertainty about sources of vulnerability is precisely the motivation for compartmentalization: if we knew which components were safe and which were not, there would be no reason to distinguish more than two levels of privilege, and we could merge each group into a single mega-component. Even in rare cases where we are certain that some code cannot be compromised—for instance because we have verified its safety—protecting only the verified code from all the rest using a fully abstract compiler \([7]\) is likely still suboptimal in terms of protection, since this provides no guarantees for all the code that is not verified.

Moreover, this weakness is not hypothetical: several fully abstract compilers proposed in the literature are only capable of protecting a single trusted module from its untrusted context \([8], [39], [50], [51], [53]\) (although recently proposed extensions \([54]\) aim at lifting this restriction in some cases). While this setup is appropriate for protecting trusted (e.g., verified) code from its untrusted context \([7]\), it is not suitable for our compartmentalization setting, where we do not know in advance which components will be dynamically compromised and which ones not and want to simultaneously protect against all possible compromise scenarios.

Solution 3: Separate compilation: We can address this by requiring that the compiler (and low-level context application operation, e.g., linker and loader) satisfy one additional property that we call separate compilation:

**Definition III.5.**

We say that a compiler \( \downarrow \) satisfies separate compilation if:

1. The type of contexts and programs is the same, so the \( \downarrow \) compiler can also compile contexts; and
2. \( (A[P])\downarrow \sim L A[[P]]_1 \) for all \( A \) and \( P \).

Requiring that context application and compilation commute (condition 2) implies that, if some complete program \( C \) can be written as both \( C = A_1[P_1] \) and \( C = A_2[P_2] \), then separately compiling each of these splits yields behaviorally equivalent results: \( A_1[P_1]\downarrow \sim L A_2[P_2]\downarrow \). With separate compilation, the interesting implication direction of full abstraction (Definition III.2) can be instantiated as follows:

\[
\forall B. \forall P, Q \text{ fully defined. } ((B[P])\downarrow \not\triangleleft L (B[Q])\downarrow) \Rightarrow \exists A. A \text{ fully defined } \Rightarrow A[P] \not\triangleleft L A[Q]
\]

One compelling reading of this is that for all ways to break a complete program into a compromised context \( B \) and an uncompromised program \( P \) (i.e., for all compromise scenarios), and for all programs \( Q \) that we can replace for \( P \), if the context \( B \) can distinguish \( P \) from \( Q \) when compiled to low-level code, then there exists a fully defined context \( A \) that can distinguish them at the high-level.

Note that in a language without undefined behavior this property would be trivially true from just (whole program) correct compilation (e.g., see Assumption IV.2) by picking \( A = B \). However, it is nontrivial for a language with undefined behavior because context \( B \) can cause undefined behavior, so correct compilation does not apply for \( B[P] \) and \( B[Q] \).

In our setting, this property allows us to replace reasoning about the low-level implications of the undefined behavior in the original context, by reasoning about a fully defined high-level context. It is trivial to check that our instance of structured full abstraction from Theorem III.4 does satisfy separate compilation.

IV. A SIMPLE INSTANCE

We illustrate secure compartmentalization in a very simple setting that captures its main features. To this end, we build a compiler from an unsafe language with procedures to an abstract machine with compartments. We show how to adapt a standard full abstraction proof technique called trace semantics \([31], [32], [52]\) to prove secure compartmentalization.

**Source:** simple imperative language with components, procedures, and buffers: We work with a simple unsafe language whose programs consist of components with interfaces. Buffer overflows are considered undefined behavior and may
open the door to low-level attacks after compilation. However,
because of the low-level compartmentalization, the effects of
these attacks will be contained to the offending component.

The syntax for expressions, given below, is that of a standard
imperative language with mutually recursive procedures and
buffers. Each component \(C\) has local buffers \(b\) and proce-
dures \((P)\). Loops are encoded using recursive calls, sequencing
is encoded as a binary operation, and variables are encoded
using buffers. In particular, the variable that holds a procedure
call argument is always passed in the first cell of the first buffer
of the callee component. For simplicity, only integers are first
class values and can be passed over component boundaries,
using procedure calls and returns. Buffers are second class.

\[
e : i \mid e_1 \otimes e_2 \mid \text{if } e \text{ then } e_1 \text{ else } e_2 \mid \text{exit}
\]

where \(\otimes \in \{\;+, -, \times, =, \leq, \ldots\}\).

The small-step operational semantics of this language is
mostly standard and reduces configurations \((C, s, \sigma, K, e)\)
where \(C\) is the current component, \(s\) is a state recording
buffers values for each component, \(\sigma\) is a global call stack,
\(K\) is an evaluation context, and \(e\) the currently evaluated
expression. This semantics is deterministic for defined programs.
Given that undefined behaviors cannot be described by the
high-level semantics in any reasonable way, we model them
as stuckness [41]. Hence, reduction gets stuck when trying to
access or update a buffer out of bounds, which are the only
forms of undefined behavior in our language. We can state
partial type safety (partial progress and usual preservation),
which says that well-formed programs can only go wrong by
reducing to an out-of-bounds operation on a buffer.

**Conjecture IV.1 (Partial progress).** For any well-formed con-
furation \(cfg\), one of the following holds:

1. \(cfg\) is a final configuration (value or exit);
2. \(cfg\) takes a step;
3. \(cfg\) is stuck but of one of the following forms:
   a. \((C, s, \sigma, b[i] = K, i)\) where \(s[C, b, i]\) is undefined;
   b. \((C, s, \sigma, b[i] = K, i')\) where \(s[C, b, i]\) is undefined.

Procedures are either public or private. While the com-
ponent that defines a private procedure is the only component
that can call it, public procedures can be called by any other
component. However, the interface of the calling component
must mention that it imports the procedure.

 Interfaces specify which procedures a component exports
and imports. To satisfy an interface, a component must define
the procedures exported in the interface and can only perform
cross-component calls to procedures imported in the interface.
In our source language, this is checked statically. Interfaces are
the same in this language and our target abstract machine.

**Target: abstract machine with interacting compartments:**
Our target machine models a RISC instruction set extended
with a compartmentalization mechanism. The compiler will
map each source component to its own compartment. The
machine has a fixed set of registers and a program counter
\((pc)\) register. Because resource exhaustion and integer-size
related problems are not part of our concerns in this paper, we
assume unbounded words and infinite memory. This memory
is split into separate infinite address spaces, one for each
compartment. Memory addressing is relative to the current
compartment. The machine has the following instructions [13]:

\[
\text{instr} ::= \text{Nop} | \text{Const } i \rightarrow r_d | \text{Mov } r_s \rightarrow r_d
\]
\[
\text{Load } *r_p \rightarrow r_d | \text{Store } *r_p \leftarrow r_s
\]
\[
\text{Jump } r | \text{Jal } r | \text{Call } C.P | \text{Return}
\]
\[
\text{Binop } r_1 \otimes r_2 \rightarrow r_d | \text{Bnz } r_i | \text{Halt}
\]

Instructions Jal (jump-and-link) and Jump redirect control
flow to the address in the same address space stored in a
register, however Jal also saves the current program counter in
a register, so that the target code can later restore it and resume
execution at the point that follows the Jal instruction. We use
instructions Call \(C.P\) and Return to change compartments.
These instructions are subject to dynamic compartmentaliza-
tion constraints; in a realistic system such constraints could
be enforced e.g. using a shadow call stack [3], [23] or return
capabilities [33]. Since Call and Return would probably come
with monitoring overhead, Jump and Jal instructions are still
useful to perform compartment-local calls. The reduction rules
for Call and Return are provided below.

\[
\begin{align*}
\text{mem}[C,pc] = i \quad &\text{decode } i = \text{Call } C'.P' \quad pe' = E[C'][P'] \\
C' = C \lor C',P' \in \psi[C].\text{import} \quad &\sigma' = (C, pc+1) : : \sigma
\end{align*}
\]

\[
\begin{align*}
\psi; E \vdash (C, \sigma, \text{mem}, \text{reg}, pc) \rightarrow (C', \sigma', \text{mem}, \text{reg}, pc') \\
mem[C,pc] = i \quad &\text{decode } i = \Return C'.P' \\
\sigma = (C',pc') : : \sigma'
\end{align*}
\]

The Call instruction checks that the call is valid with
respect to the current compartment’s interface—i.e. the target
procedure is imported by the current compartment—which
ensures that even if a compiled component is compromised
it cannot exceed its static privilege level. Then, Call puts the
calling compartment’s name and program counter on a global,
protected call stack. Finally, it redirects control flow to the
entry point of the called procedure. The Return instruction
is issued to give back control to the caller compartment.
It retrieves the caller’s name and return address from the
protected call stack and resumes execution.

**Simple secure compiler:** We employ a very simple com-
piler which is nonetheless interesting in two ways. First, the
compiler plays a role in secure compartmentalization by being
defensive regarding registers values flowing from or to other
compartments. Second, it implements an optimization that
distinguishes cross-compartment calls from local calls and thus
removes security overhead on local calls.

The compiler can obviously not rely on other compartments
to preserve its register invariants, so it must store and restore
its register environment upon cross-compartment calls and
returns. Storing the environment means saving the current call
argument on the stack and writing the current stack pointer
to a fixed location in the current compartment’s memory. The
compiler must also clear the registers it uses before issuing a
Call or Return instruction. Practically, this prevents compromised compartments from reading intermediate computation results which could hold sensible information.

Our proof actually relies on the compiler cleaning all registers but the one that holds the call argument or return value. Formally, the compromised compartments must have no way to distinguish an uncompromised program from another based on low-level information that high-level fully defined attackers don’t get. Not cleaning unused registers would however provide a covert channel to two compromised compartments whose interfaces do not prevent them from having direct communication. These compartments could now exchange values through uncleared registers by interacting with an unsuspecting uncompromised component. We conjecture that this possible cooperation between compromised components doesn’t yield more attacker power in our very simple setting. In a slightly different setting where registers could be used to transmit e.g. capabilities, however, this would give more power to the attacker, so our compiler clears all but one registers, which also simplifies our proof.

The compiler partitions each compiled component’s memory into buffers, procedures’ code, and a local stack that can grow infinitely. This local stack is used both to store intermediate results and as a call stack. Because all procedures of a component will live in the same address space, local calls need not be monitored by the compartmentalization mechanism and can be implemented more efficiently using regular JAL and Jump instructions. We implement this optimization by using different procedure entry points for internal and external calls, and go a little further by skipping, for local calls, the steps that store and restore the register environment and perform register cleaning.

Because we do not check bounds when compiling buffer read and write operations, buffer overflows can corrupt a compartment’s memory in arbitrary ways. Many buffer overflow attacks can be reproduced in our simple setting, including, due of the optimization above, return-oriented programming attacks [15], [58]. In return-oriented programming, an attacker overwrites return addresses on the local stack to produce an unexpected sequence of instructions of his will by reusing parts of the code of component-local procedures. In our setting, buffer overflow attacks allow compiled components that are not defensive enough to shoot themselves in the foot. As we will prove (Theorem IV.7), buffer overflows can only do limited harm to other compiled components though.

We assume compiler correctness as stated below for our compiler. Note that in the presence of partial type safety (Conjecture IV.1), the determinism of our source language implies that a source program either terminates or diverges. Similarly, because we equate stuckness and termination in our deterministic abstract machine, target programs also either terminate or diverge. As a consequence, proving either (1) or (2) below is enough to get the other.

Assumption IV.2 (Compiler correctness).

∀P. P defined ⇒ (1) P terminates ⇔ P↓ terminates
(2) P diverges ⇔ P↓ diverges

Combining this assumption with separate compilation (Definition III.5), which trivially holds for our compiler, we immediately get the following corollary:

Corollary IV.3 (Separate compiler correctness).

∀s, A ⊢ s, P → s. A fully defined wrt s ⇒ P fully defined wrt os ⇒

Proof technique for structured full abstraction: Trace semantics were proposed by Jeffrey and Rathke [31], [32] to define fully abstract models for high-level languages via may testing. Patrignani et al. [51], [52] later showed how to use such a trace semantics [52] to prove full abstraction for a compiler targeting machine code [51].

The proof technique has two steps: First, one devises a trace semantics for low-level partial programs and contexts, which is very closely related to the target machine’s operational semantics (Assumption IV.4 and Assumption IV.5). Then, one uses the traces that characterize the interaction between an arbitrary low-level context and two compiled programs this context distinguishes, to build a high-level attacker that also distinguishes the programs (Assumption IV.9).

Adapting the technique to undefined behavior takes little effort: It amounts to proving standard full abstraction for the safe subset of the language. So, one simply needs to make sure that the context produced by the mapping is fully-defined, thus safe. However, adapting the technique to a closed world, as needed for structured full abstraction, takes more work. In our setting, even though the context owns many components, they all live within different compartments. A compartmentalization mechanism like ours does not only limit interaction between the program and the context, it also limits the internal actions within the program and within the context. The rule for the Call instruction would for instance prevent any direct communication among two context compartments that do not import each other’s procedures. Hence, when the context has control, the calls that it can make do not only depend on whether one of its components has imported the target procedure. It also depends on whether there is an internal chain of calls going from the component used as a context entry point to the component that could do this call back to the program.

Technically, adapting the technique to a closed world requires changes within the trace semantics itself rather than in the proof structure: Context external and internal actions must adhere to a fixed interface, instead of an arbitrary interface that the context could choose. We force this adherence by applying two changes: we track the context’s internal state and internal actions, and we use interfaces to constrain the actions (both internal and external) that can be taken from a given context internal state.
Trace semantics for the low-level language: We define a trace semantics where traces are words over an alphabet of external actions \( E_\alpha \) defined as follows:
\[
E_\alpha ::= \gamma? \mid \gamma!
\]
\[
\gamma ::= \text{Call}_{\text{reg}} \ C \mid \text{Return}_{\text{reg}} \mid \checkmark
\]

Traces alternate between program external actions “\( \gamma? \)” and context external actions “\( \gamma! \)”. Seeing actions as moves in a game, we say that the context and the partial program are each other’s opponent. Possible actions are program termination (\( \gamma = \checkmark \)), to which the opponent cannot respond (\( \checkmark \) ends the game), and cross-boundary communication actions (\( \gamma \neq \checkmark \)) which give control to the opponent. Programs and contexts can also diverge by going into an infinite chain of local actions, in this case they produce no action.

Intuitively, a program \( p \) has trace \( t \) if it answers with the program actions described in \( t \) when facing a context \( a \) that produces the context actions described in \( t \). We make no formal distinction between partial programs and contexts, so asking that \( a \) produces the context actions described in \( t \) is the same as asking that \( a \) (seen as a partial program) produces the program actions in \( t^{-1} \) when facing \( p \), where \( t^{-1} \) is obtained by swapping “\( ? \)” for “\( \gamma? \)” and “\( \gamma! \)” for “\( ! \)” in \( t \). The intuition is that, from \( a \)’s point of view, \( a \) is the program and \( p \) is the context. We define \( Tr(p) \) to be the set of traces of a partial program \( p \), and make use of this definition to talk about the traces of a context: context \( a \) produces the context actions in trace \( t \) when facing a program \( p \) that produces the program actions in \( t \) if and only if \( t^{-1} \in Tr(a) \).

The trace semantics is deterministic with respect to program actions and nondeterministic with respect to context actions. Nondeterminism disappears once we choose a particular context for a program, as the two useful lemmas below illustrate.

**Assumption IV.4 (Trace decomposition).**

\[
\forall s,p,a,t. \quad p \vdash s \cdot a \vdash s \cdot a[p] \text{ terminates } \Rightarrow \exists t. \quad t \text{ ends with } \checkmark \land t \in Tr(p) \land t^{-1} \in Tr(a)
\]

Trace decomposition is stated for terminating programs. It extracts the interaction between a program and a context with dual shapes by looking at how they reduce, synthesizing that interaction into dual traces \( t \) and \( t^{-1} \). Because execution terminates, these traces end with a termination marker.

**Assumption IV.5 (Trace composition).**

\[
\forall s,p,a,t \neq \epsilon. \quad p \vdash s \cdot a \vdash s \cdot a \in Tr(p) \land t \in Tr(p) \land t^{-1} \in Tr(a) \Rightarrow (\forall E_\alpha. \quad t.E_\alpha \notin Tr(p) \lor (t.E_\alpha)^{-1} \notin Tr(a)) \Rightarrow (a[p] \text{ terminates } \iff t \text{ ends with } \checkmark)
\]

Trace composition is the opposite of trace decomposition, and reconstructs a sequence of reductions based on synthesized interaction information: It considers a program and a context with dual shapes, such that one has trace \( t \) and the other has the dual trace \( t^{-1} \). The condition on the second line states that interaction is over: trace \( t \) cannot be extended by any action \( E_\alpha \) while having the dual \( (t.E_\alpha)^{-1} \) in \( Tr(a) \). Under these assumptions, trace composition tells us that one of the following holds: either (1) the trace ends with a termination marker \( \checkmark \) and in this case putting \( p \) in context \( a \) will produce a terminating program, or (2) putting \( p \) in context \( a \) will produce a diverging program. Intuitively, if interaction is over but there is no termination marker, it must be because the program or the context went into an infinite sequence of internal actions and will neither give control back nor terminate.

The statement of these assumptions and their equivalent in an open world setting are very close: for full abstraction, we wouldn’t care about the shapes of the program. However, we needed to adapt the trace semantics itself to deal with our closed world assumption. We did this by adding more options to the non-deterministic choice of the context action, by incorporating context internal actions within the trace semantics. To this end, we keep track of the current context component. When in control, the only external communication the context can perform is that allowed by the interface of the current component, but it can also perform internal communication actions that make it switch components (if allowed by the interface of the current component). So we use finer-grained traces that include internal communication and can later be directly mapped to high-level attackers (Assumption IV.9). The traces we use otherwise are obtained by erasing internal actions from the finer-grained traces.

Another property that will be useful for our proof is trace extensibility: it states formally that for each external action a context \( a \) would take as a partial program that has control, there is a non-deterministic choice leading to this action on the program’s side of the trace semantics. Note that because there is no formal distinction between partial program and context, the same property holds when swapping the two roles.

**Assumption IV.6 (Trace extensibility).**

\[
\forall s,p,a,t. \quad p \vdash s \cdot a \vdash s \cdot a \Rightarrow \exists t. \quad t \in Tr(p) \land (t.E_\alpha)^{-1} \in Tr(a) \Rightarrow t.E_\alpha \in Tr(p)
\]

**Secure compartmentalization proof:** We prove structured full abstraction, which implies secure compartmentalization by our general result from \( \S \text{III} \) (Theorem III.4).

**Theorem IV.7 (Structured full abstraction).** Our compiler satisfies structured full abstraction.

To prove structured full abstraction, we need to produce a fully defined high-level context that distinguishes two fully defined programs, knowing that some arbitrary low-level context can already distinguish them. To do this, we will have the high-level attacker reproduce part of the interaction between the low-level attacker and the programs, by using traces to characterize the interaction. We can then use the trace semantics to justify that the low-level interaction between the compiled fully defined context and the compiled fully defined programs lead to different behaviors, and apply compiler correctness to justify that the same happens in the low-level.

However, it turns out that compiled fully defined attackers cannot reproduce arbitrary traces: Before giving control to
the program, compiled fully defined attackers always clear all registers but the one used for communication. As a consequence, a compiled fully defined context can only produce traces in which registers are cleared in all context actions. Hopefully, a characteristic of fully defined programs comes to the rescue: Because they do not trust other components, compiled fully defined programs always overwrite registers before reading their values, except for the register used for communication. As a consequence, they cannot distinguish context actions based on the content of these unread registers, which are exactly the ones a compiled fully defined context cleans. This intuition yields the following assumption.

**Assumption IV.8 (Canonicalization).**

\[
\forall s, P, t. P \vdash \bullet s \land P \text{ fully defined wrt } \circ s \Rightarrow t \in \text{Tr}(P) \equiv \exists \gamma. A \vdash \circ s \land A \text{ fully defined wrt } \bullet s \land (1) t^{-1} \in \text{Tr}(A) \land (2) \gamma \neq \square \Rightarrow (t.\gamma!\gamma'?)^{-1} \in \text{Tr}(A) \land (3) \forall \gamma, \gamma'. (t.\gamma!\gamma'?)^{-1} \in \text{Tr}(A) \Rightarrow \zeta(\gamma) = \zeta(\gamma_0) \land \gamma' = \square
\]

Trace \( \zeta(t) \) is obtained from \( t \) by rewriting context external actions so that the context always clears registers. Because it clears registers, a compiled fully defined context can only hope to reproduce canonical traces, i.e. traces that satisfy \( t = \zeta(t) \). However, **Assumption IV.8** tells us that when confronted to a canonical trace \( \zeta(t) \), fully defined programs perform the exact same actions as the ones they take when facing the original trace \( t \). So in our proof, having the high-level attacker reproduce canonical traces instead of the original traces of the low-level context will be enough to lead programs into reproducing the actions they took when facing the low-level context. The definability assumption below gives the mapping from a trace to a compiled fully defined context that reproduces a canonical trace and, depending on the last action the program takes, ends interaction with either termination or divergence.

**Assumption IV.9 (Definability).**

\[
\forall s, t, \gamma_0. t = \zeta(t) \land (\exists p. P \vdash \bullet s \land t.\gamma_0 \in \text{Tr}(p)) \Rightarrow \exists A. A \vdash \circ s \land (1) t^{-1} \in \text{Tr}(A) \land (2) \gamma_0 \neq \square \Rightarrow (t.\gamma_0!\gamma'?)^{-1} \in \text{Tr}(A) \land (3) \forall \gamma, \gamma'. (t.\gamma!\gamma'?)^{-1} \in \text{Tr}(A) \Rightarrow \zeta(\gamma) = \zeta(\gamma_0) \land \gamma' = \square
\]

Consider a canonical trace \( t \), a program action \( \gamma_0 \), and a program \( p \), such that \( p \) performs action \( \gamma_0 \) after following trace \( t \). In this scenario, the definability assumption gives us a fully defined attacker that will follow the dual trace (condition (1)). At this point, the program will either diverge or trigger an action \( \gamma! \), which could terminate execution (\( \gamma = \square \)) or give control back to the context. If program action \( \gamma! \) gives control back to the context (\( \gamma \neq \square \)), the next move of the high-level attacker will be to make execution either terminate or diverge depending on action \( \gamma \): First, if \( \gamma = \gamma_0 \), the context will make execution terminate (condition (2)). Second, if the context can distinguish \( \gamma \) from \( \gamma_0 \) (\( \zeta(\gamma) \neq \zeta(\gamma_0) \)), it will make execution diverge. Equivalently (condition (3)), the only context action \( \gamma! \) that may follow program action \( \gamma \) is termination (\( \gamma' = \square \)), and it can only happen when the context cannot distinguish \( \gamma \) from \( \gamma_0 \) (\( \zeta(\gamma) = \zeta(\gamma_0) \)).

The intuition regarding distinguishability here is the same as for canonicalization (**Assumption IV.8**): As a compiled fully defined context, \( A \) only reads communicated values from register \( \text{r}_\text{com} \), the register that holds the call argument or return value. So \( A \) cannot possibly distinguish between \( \gamma \) and \( \gamma_0 \) when they only differ by holding different values in registers \( r \neq \text{r}_\text{com} \). Formally, this translates to the following: \( A \) cannot distinguish between \( \gamma \) and \( \gamma_0 \) when \( \zeta(\gamma) = \zeta(\gamma_0) \), where \( \zeta \) is the rewriting that clears these all registers but this one.

We built an algorithm (in OCaml) that constructs \( A \) out of \( t \). More precisely, the algorithm inputs a trace with internal cross-component actions (the finer-grained trace that erases to \( t \)) and builds an \( A \) that successfully reproduces context internal and external actions as prescribed by that trace. When, after following \( t \), the program finally gives back control to \( A \) by issuing action \( \gamma \), \( A \) chooses a different appropriate next action depending on the point in \( A \) to which action \( \gamma \) leads control flow. Because we input a trace that includes internal actions, and thus do not need to reconstruct these actions, our algorithm was not more difficult to come with than one for an open-world setting [51]. In the following, we assume that the algorithm is correct, i.e. that IV.9 holds. We can now prove our main theorem.

**Detailed proof of structured full abstraction:** Consider a low-level attacker \( a \vdash_\circ s \) that distinguishes between two fully defined partial programs with a dual shape \( P, Q \vdash_\bullet s \). Suppose w.l.o.g. that \( a[P] \) terminates and \( a[Q] \) diverges. We will build a fully defined high-level attacker \( A \vdash_\circ s \) that can distinguish between \( P \) and \( Q \).

We can first apply trace decomposition (**Assumption IV.4**) to \( a \) and \( P \vdash_\bullet s \) to get a terminating trace \( t_i \) of \( P \vdash_\bullet s \) such that \( t_i^{-1} \in \text{Tr}(a) \). Call \( t_p \) the longest prefix of \( t_i \) such that \( t_p \in \text{Tr}(Q) \). Because trace sets are prefix-closed by construction, we know that \( t_p \in \text{Tr}(P) \) and \( t_i^{-1} \in \text{Tr}(a) \).

Moreover, \( t_p \) is necessarily a strict prefix of \( t_i \): otherwise, we could apply trace composition (**Assumption IV.5**) to \( a \) and \( Q \vdash_\bullet s \) and get that \( a[Q] \) terminates, which is false. So there exists an external action \( E_{\alpha} \) such that also trace \( t_p. E_{\alpha} \) is a prefix of \( t_i \). \( E_{\alpha} \) cannot be a context action, or trace extensibility (**Assumption IV.6**) would imply that \( t_p. E_{\alpha} \) is a trace of \( \text{Tr}(Q) \), which is incompatible with \( t_p \) being the longest prefix of \( t_i \) in \( \text{Tr}(Q) \). Therefore, \( E_{\alpha} \) is a program action, i.e. there exists \( \gamma_0 \) such that \( E_{\alpha} = \gamma_0 \). Intuitively, \( P \vdash_\bullet s \) and \( Q \vdash_\bullet s \) take the same external actions until the end of \( t \), where \( P \vdash_\bullet s \) takes external action \( \gamma_0 \) and \( Q \vdash_\bullet s \) does not (it either takes a different action \( \gamma \neq \gamma_0 \), or no external action at all).

Now, let \( t_c \) be the canonicalization of trace \( t_p \), i.e. \( t_c = \zeta(t_p) \). By canonicalization (**Assumption IV.8**), \( t_c.\gamma_0 = \zeta(t_p.\gamma_0) \) is a trace of \( P \vdash_\bullet s \). We can thus use apply definability (**Assumption IV.9**) to trace \( t_c \) and action \( \gamma_0 \), using \( P \vdash_\bullet s \) as a witness having trace \( t_c.\gamma_0 \). This yields a fully defined context \( A \vdash_\circ s \) such that:

\[
\begin{align*}
(1) & t_c^{-1} \in \text{Tr}(A) \\
(2) & \gamma_0 \neq \square \Rightarrow (t_c.\gamma_0!\gamma'?)^{-1} \in \text{Tr}(A) \\
(3) & (\gamma' \gamma. t_c.\gamma'?) \in \text{Tr}(A) \Rightarrow \zeta(\gamma) = \zeta(\gamma_0)
\end{align*}
\]
We now show that conditions (1), (2) and (3) ensure that $A\downarrow$ distinguishes $P\downarrow$ from $Q\downarrow$: $A\downarrow [P]\downarrow$ terminates while $A\downarrow [Q]\downarrow$ diverges.

First, we look at $P\downarrow$. Consider the case where $\gamma_0 = \checkmark$. In this case, by applying trace extensibility to $A \downarrow$ in (1) (seeing $A$ as a partial program), we get that $t^{-1}.\checkmark!^{-1}$ is a trace of $A\downarrow$, so trace composition allows us to conclude that $A\downarrow [P]\downarrow$ terminates. Now if $\gamma_0 \neq \checkmark$ then this action gives back control to the context, which, given (2), will perform action $\checkmark'$. Applying trace extensibility to $P\downarrow$, $P\downarrow$ has trace $t_c.\gamma_0!\checkmark'$, so we can apply trace composition and deduce that $A\downarrow [P]\downarrow$ terminates in this case as well.

Now, regarding $Q\downarrow$, we first obtain the following by applying canonicalization to $t_p, t_p^{-1}, t_p.\checkmark!$ and $t_p.\gamma_0!$:

(a) $t_c \in \text{Tr}(Q\downarrow) \land \zeta(t_p^{-1}) \in \text{Tr}(A\downarrow)$
(b) $t_c.\checkmark! = \zeta(t_p.\checkmark!) \in \text{Tr}(Q\downarrow) \Rightarrow t_p.\checkmark! \in \text{Tr}(Q\downarrow)$
(c) $t_c.\gamma_0! = \zeta(t_p.\gamma_0!) \in \text{Tr}(Q\downarrow) \Rightarrow t_p.\gamma_0! \in \text{Tr}(Q\downarrow)$

Note that in (a), $t_p^{-1}$ is canonical, i.e. $\zeta(t_p^{-1}) = \zeta(t_p)^{-1}$. Indeed, compiled fully defined contexts only produce canonical traces and $P\downarrow$ (seen as a context for $A\downarrow$) produces this trace. So $\zeta(t_p^{-1}) = t_c^{-1}$.

After following trace $t_c$, $Q\downarrow$ cannot perform a terminating action: otherwise from (b) we could apply trace composition to $a$ and $Q\downarrow$ and get that $a(Q\downarrow)$ terminates, which is false. $Q\downarrow$ cannot perform action $\gamma_0$ either, since (c) would then violate the fact that $t_p$ is the longest prefix of $t_i$ in $\text{Tr}(Q\downarrow)$. So $Q\downarrow$ only has two options left.

The first option of $Q\downarrow$ is to perform no external action by going into an infinite sequence of internal transitions. In this case we can apply trace composition to get that $A\downarrow [Q\downarrow]$ diverges. The second option is to give control back to the context using an external action $\gamma$ so that $\checkmark \neq \gamma \neq \gamma_0$. Because fully defined compiled programs clean registers, they only yield canonical actions, i.e. $\gamma = \zeta(\gamma) \land \gamma_0 = \zeta(\gamma_0)$. Combined with (3), this entails that if $A\downarrow$ produced an action $\gamma'$, we would have $\gamma = \gamma_0$, which is false. Hence, $A\downarrow$ doesn’t produce any action: it goes into an infinite sequence of local transitions. We can again apply trace composition to get that $A\downarrow [Q\downarrow]$ diverges.

We finally separate compiler correctness (Corollary IV.3) to conclude the proof.

V. RELATED WORK

Fully abstract compilation: Fully abstract compilation was introduced in the seminal work of Martín Abadi [1] and later investigated by the academic community. (Much before this, the concept of full abstraction was coined by Milner [42].) For instance, Ahmed et al. [9]–[11] proved the full abstraction of type-preserving compiler passes for functional languages and devised proof techniques for typed target languages. Abadi and Plotkin [6] and Jagadeesan et al. [30] expressed the protection provided by a mitigation technique called address space layout randomization as a probabilistic variant of full abstraction. Fournet et al. [26] devised a fully abstract compiler from a subset of ML to JavaScript.

Patrignani et al. [39], [51] were recently the first to study secure compilation to machine code, starting from single modules written in simple, idealized object-oriented and functional languages and targeting hardware architectures with a new coarse-grained isolation mechanism. They also recently proposed proof techniques for full abstraction that work for untyped target languages [19], [52]. All this work studies fully abstract compilers that by design violate our separate compartmentation property, so it cannot be applied to our compartmentalization setting. More recent work by Patrignani et al. [54] proposes to study “multi-module fully abstract compilation”, and we expect that our secure compartmentalization property could be useful in that setting.

Verifying correct compartmentalization: Recent work focused on formally verifying correct compartmentalization mechanisms based on software fault isolation [36], [43], [65] or tagged hardware [13]. This work, however, only considers on the correctness of the low-level enforcement mechanism, not high-level security properties and reasoning principles. Communication between components is generally done by jumping to a specified set of entry points, while the model we consider in §IV is more structured and enforces correct calls and returns. Finally, seL4 is a verified operating system microkernel [35], that uses a capability system to separate user level threads and for which correct access control [57] and noninterference properties [44] were proved formally.

VI. CONCLUSION AND FUTURE WORK

We have introduced a new secure compartmentalization property, related it to the more established notion of full abstraction, and applied this property in a carefully simplified setting: a simple imperative language with procedures compiling to a compartmentalized abstract machine. This lays the formal foundations for studying the secure low-level interaction of mutually distrustful components.

In the future we plan to use this groundwork to study more realistic source and target languages, compilers, and enforcement mechanisms. In the long run we would like to apply this to the C language by devising a secure variant of CompCert and proving its security formally. We would like to implement the abstract compartmentalization machine from §IV in terms of various enforcement mechanisms, including: process-level sandboxing [14], [28], [34], [55], software-fault isolation (SFI) [64], micro-policies [13], capability machines [62], and multi-PMA systems [54]. As we target lower-level machines, new problems will appear: for instance trying to hide components’ size and resource consumption will likely be too costly, and we will need to weaken our property to allow controlled release of such information to the attacker. Finally, we would like to study more interesting compartmentalization models including dynamic component creation and nested components, and the way these extensions influence the security property.
Acknowledgments: We are grateful to Andrew Tolmach for helpful discussions and thoughtful feedback on earlier drafts. Yannis Juglaret is supported by a PhD grant from the French Department of Defense (DGA) and Inria. Arthur Azevedo de Amorim and Benjamin C. Pierce are supported by NSF award 1513854, Micro-Policies: A Framework for Tag-Based Security Monitors.

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