VPS: Excavating High-Level C++ Constructs from Low-Level Binaries to Protect Dynamic Dispatching

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ABSTRACT

Polymorphism and inheritance make C++ suitable for writing complex software, but significantly increase the attack surface because the implementation relies on *virtual function tables (vtables).* These vtables contain function pointers that attackers can potentially hijack and in practice, *vtable hijacking* is one of the most important attack vector for C++ binaries.

In this paper, we present *VTable Pointer Separation* (VPS), a practical binary-level defense against vtable hijacking in C++ applications. Unlike previous binary-level defenses, which rely on unsound static analyses to match classes to virtual callsites, VPS achieves a more accurate protection by restricting virtual callsites to validly created objects. More specifically, VPS ensures that virtual callsites can only use objects created at valid object construction sites, and only if those objects can reach the callsite. Moreover, VPS explicitly prevents false positives (falsely identified virtual callsites) from breaking the binary, an issue existing work does not handle correctly or at all. We evaluate the prototype implementation of VPS on a diverse set of complex, real-world applications (MongoDB, MySQL server, Node.js, SPEC CPU2017/CPU2006), showing that our approach protects on average 97.8% of all virtual callsites in SPEC CPU2006 and 97.4% in SPEC CPU2017 (all C++ benchmarks), with a moderate performance overhead of 11% and 9% geomean, respectively. Furthermore, our evaluation reveals 86 false negatives in *VTV*, a popular source-based defense which is part of GCC.

CCS CONCEPTS

• **Security and privacy** → *Software reverse engineering*.

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1 INTRODUCTION

Software implemented in the C++ language is vulnerable to increasingly sophisticated memory corruption attacks [8, 9, 19, 38, 44, 46]. C++ is often the language of choice for complex software because it allows developers to structure software by encapsulating data and functionality in *classes*, simplifying the development process. Unfortunately, the binary-level implementations of C++ features such as polymorphism and inheritance are vulnerable to control-flow hijacking attacks, most notably *vtable hijacking*. This attack technique abuses common binary-level implementations of C++ virtual methods where every object with virtual methods contains a pointer to a *virtual function table* (*vtable*) that stores the addresses of all the class's virtual functions. To call a virtual function, the compiler inserts an indirect call through the corresponding vtable entry (a *virtual callsite*). Using temporal or spatial memory corruption vulnerabilities such as arbitrary write primitives or use-after-free bugs, attackers can overwrite the vtable pointer so that subsequent virtual calls use addresses in an attacker-controlled alternative vtable, thereby hijacking the control flow. In practice, vtable hijacking is a common exploitation technique widely used in exploits that target complex applications written in C++ such as web browser and server applications [42].

Control-Flow Integrity (CFI) solutions [2, 5, 30, 34, 43, 45, 47] protect indirect calls by verifying that control flow is consistent with a Control-Flow Graph (CFG) derived through static analysis. However, most generic CFI solutions do not take C++ semantics into account and leave the attacker with enough wiggle room to build an exploit [19, 38]. Consequently, approaches that specifically protect virtual callsites in C++ programs have become popular. If source code is available, compiler-level defenses can benefit from the rich class hierarchy information available at the source level [6, 7, 43, 49]. However, various legacy applications are still in use [31] or proprietary binaries have to be protected which do not offer access to the source code (e.g., Adobe Flash [3]). Here, binary-level defenses [14, 17, 33, 35, 48] must rely on (automated) binary analysis techniques to reconstruct the information needed to guarantee security and correctness.

In this paper, we present *VTable Pointer Separation* (vps), a binary-level defense against vtable hijacking attacks. Unlike previous binary-only approaches that restrict the set of vtables permitted for each virtual callsite, we check that the vtable pointer remains unmodified after object creation. Intuitively, vps checks the vtable pointer's integrity at every callsite. Because the vtable pointer in a legitimate live object never changes and the virtual callsite uses it to determine its target function, vps effectively prevents vtable hijacking attacks. In essence, we want to bring a defense as powerful as *CFIXX* [7] (which operates at the source level) to binary-only applications, even though none of the information needed for the defense is available. Our approach is suitable for binaries because, unlike other binary-level solutions, we avoid the inherent inaccuracy in binary-level CFG and class hierarchy reconstruction. Because vps allows only the initial virtual pointer(s) of the object to ever exist, we reduce the attack surface even compared to hypothetical implementations of prior approaches that statically find the set of possible vcall targets with perfect accuracy.

Given that binary-level static analysis is challenging and unsound in practice, and may lead to false positives in identifying virtual callsites, we carefully deal with such cases by over-approximating the set of callsites and implementing an (efficient) slow path to handle possible false positives at runtime. Meanwhile, vps handles all previously verified callsite with highly-optimized fast checks. This approach allows us to prevent false positives from breaking the application as they do in existing work [14, 17, 35, 48]. Additionally, while existing work [24–26, 33] only considers *directly* referenced vtables, compilers also generate code that references vtables *indirectly*, e.g., through the Global Offset Table (GOT). vps can find all code locations that instantiate objects by writing the vtable, including objects with indirect vtable references.

Our prototype of vps is precise enough to handle complex, real-world C++ applications such as MongoDB, MySQL server, Node.js, and all C++ applications contained in the SPEC CPU2006 and CPU2017 benchmarks. Compared to the source code based approach *VTV*, which is part of GCC [43], we can on average correctly identify 97.8% and 97.4% of the virtual callsites in SPEC CPU2006 and SPEC CPU2017, with a precision of 95.6% and 91.1%, respectively. Interestingly, our evaluation also revealed 86 virtual callsites that are *not* protected by VTV, even though it has access to the source code. A further investigation with the help of the VTV maintainer showed that this is due to a conceptual problem in VTV which requires non-trivial engineering to fix. Compared to the source code based approach CFIXX, vps shows an accuracy of 99.6% and 99.5% on average for SPEC CPU2006 and CPU2017 with a precision of 97.0% and 96.9%. These comparisons show that vps's binary-level protection of virtual callsites closely approaches that of source-level solutions. While this still leaves a small attack window, it further closes the gap between binary-only and source-level approaches making vtable hijacking attempts mostly impractical.

Compared to state-of-the-art binary-level analysis frameworks like *Marx* [33], our analysis identifies 26.5% more virtual callsites in SPEC CPU2017 and thus offers improved protection. vps induces geomean performance overhead of 9% for all C++ applications in SPEC CPU2017 and 11% for SPEC CPU2006, which is slightly more than *Marx* induces but with significantly better protection.

Contributions. We provide the following contributions:

- We present vps, a binary-only defense against vtable hijacking attacks that sidesteps the imprecision problems of prior work on this topic. The key insight is that vtable pointers only change during initialization and destruction of an object (never in between), a property that vps can efficiently enforce.
- We develop an instrumentation approach that is capable of handling false positives in the identification of C++ virtual callsites which would otherwise break the application and which most existing work ignores. Unlike prior work, we also handle indirect vtable references.
- Our evaluation shows that our binary-level instrumentation protects nearly the same number of virtual callsites as the source-level defenses *VTV* and *CFIXX*. In addition, our evaluation uncovered a conceptual problem causing false negatives in VTV (part of GCC).

The prototype implementation of vps and the data we used for the evaluation are available under an open-source license at https://github.com/RUB-SysSec/VPS.

2 C++ AT THE BINARY LEVEL

This section provides background on C++ internals needed to understand how vps handles C++ binaries. We focus on how high-level C++ constructs translate to the binary level. For a more detailed overview of high-level C++ concepts, we refer to the corresponding literature [41].

2.1 Virtual Function Tables

C++ supports the paradigm of object-oriented programming (OOP) with polymorphism and (multiple) inheritance. A class can inherit functions and fields from another class. The class that inherits is called the *derived* class and the class from which it inherits is the *base* class. In addition to *single inheritance* (one class inherits from one other class), C++ also allows *multiple inheritance*, where a derived class has multiple base classes. A base class can declare a function as *virtual*, which allows derived classes to override it with their own implementations. Programmers may choose not to implement some functions in a base class, so called *pure virtual* functions. Classes containing such functions are *abstract* classes and cannot be instantiated. Classes deriving from an abstract base can only be instantiated if they override all pure virtual functions.

Polymorphism is implemented at the binary level using *virtual function tables* (*vtables*) that consist of the addresses of all virtual functions of a particular class. Each class containing at least one virtual function has a vtable. Instantiated classes (called *objects*) hold a pointer to their corresponding vtable, which is typically stored in read-only memory. Since each class has its own corresponding vtable, it can also be considered as the type of the object. Throughout this paper, we refer to the pointer to a vtable as a *vtblptr*, while the pointer to the object is called *thisptr*.

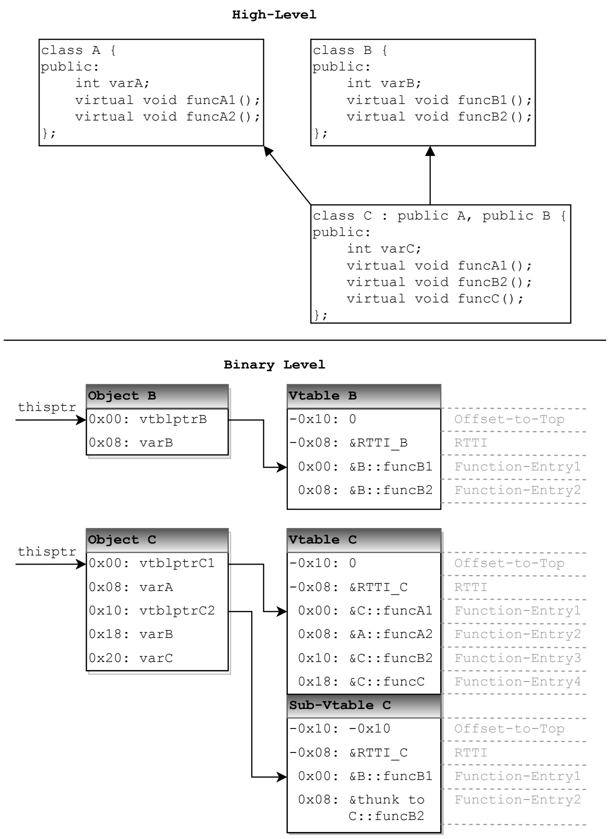


Figure 1: Example C++ class structure. The code at the top shows base classes *A* and *B*; derived class *C* which overrides virtual functions *funcA1* and *funcB2*. The bottom shows the binary-level structure of objects of classes *B* and *C*.

The Itanium C++ ABI [16] defines the vtable layout for Linux systems.[1] The *vtblptr* points to the first function entry in the vtable, and the vtable contains an entry for each virtual function (either inherited or newly declared) in the class. For example, in Figure 1, class *B*'s vtable contains two function entries because the class implements virtual functions *funcB1* and *funcB2*. Class *C* inherits from two classes, *A* and *B*, and therefore has two vtables (a base vtable and one sub-vtable). The base vtable contains all virtual functions inherited from class *A* and implemented by class *C*. The sub-vtable is a copy of class *B*'s vtable with a special entry that refers to the overwritten virtual function (called a *thunk*). Preceding the function entries, a vtable has two metadata fields: *Runtime Type Identification* (RTTI) and *Offset-to-Top*. RTTI holds a pointer to type information about the class. Among other things, this type information contains the name of the class and its base classes. However, RTTI is optional and often omitted by the compiler. It is only needed when the programmer uses, e.g., *dynamic_cast* or *type_info*. Hence, a reliable static analysis cannot rely on this information. Classes that do not contain RTTI have the RTTI field set to zero. Offset-to-Top is needed when a class uses multiple inheritance (hence has a base vtable and one or more sub-vtables) as class *C* does. Offset-to-Top specifies the distance between a sub-vtable's own *vtblptr* and the base *vtblptr* at the beginning of the object. In our example, the *vtblptr* to class *C*'s sub-vtable resides at offset `0x10` in the object, while the *vtblptr* to the base vtable resides at offset `0x0`. Hence, the distance between the two, as stored in the Offset-to-Top field in sub-vtable *C*, is `-0x10`. Offset-to-Top is `0` if the vtable is the base vtable of the class or no multiple inheritance is used.

Vtables can contain one additional field, called *Virtual-Base-Offset*, but it is only used in case of virtual inheritance, an advanced C++ feature for classes that inherit from the same base multiple times (diamond-shaped inheritance). An explanation is out of scope here because VPS needs no adaptations to support virtual inheritance, so we defer to the ABI [16].

[1] Linux uses the Itanium C++ ABI for x86-64 (amd64), our target architecture.

2.2 C++ Object Initialization and Destruction

Because VPS secures virtual callsites by protecting the *vtblptr* set at initialization time, we explain object initialization of classes with vtables. For the remainder of this paper, we only consider classes and objects that have at least one virtual function and therefore a vtable.

During object instantiation, the *vtblptr* is written into the object by the *constructor*. The lower part of Figure 1 depicts an object's memory layout at the binary level. The *vtblptr* is at offset `0x0`, the start of the object. For classes with multiple inheritance, the constructor also initializes *vtblptr*s to the sub-vtable(s). In addition, the programmer may initialize class-specific fields in the constructor. These fields are located after the *vtblptr* and, in case of multiple inheritance, after any sub-*vtblptr*s.

For classes that have one or more base classes, the constructors of the base classes are called before the derived class's own initialization code. As a result, the base class places its *vtblptr* into the object, which is subsequently overwritten by the derived class's *vtblptr*. Depending on the optimization level, constructors are often inlined, which may complicate binary analysis that aims to detect constructors.

An analogous principle is applied for object destruction through *destructor* functions. However, the destructors are executed in reversed order (destructor of the base class is executed last).

Abstract classes form a special case: although programmers cannot instantiate abstract classes, and despite the fact that their vtables contain *pure_virtual* function entries, the compiler can still emit code that writes the *vtblptr* to an abstract class into an object. However, this happens only when creating or releasing an object of a derived class, and the abstract *vtblptr* is immediately overwritten.

2.3 C++ Virtual Function Dispatch

Because classes can override virtual functions, the compiler cannot determine the target of a call to such a function at compile time. Therefore, the emitted binary code uses an indirect function call through the vtable of the object. This is called a *virtual function call*, or *vcall* for short. In the Itanium C++ ABI [16], the compiler

Table 1: C++ binary-only mitigation mechanisms

Defense	Binary-only	Protects vcalls	Protects type	Protects dangl. ptrs	Tolerates FP vcalls	Security Strategy
Marx (VTable) [33]	✓	✓	✗	✓	✓	*vtblptr* in reconstructed class hierarchy (fallback *PathArmor* [45]).
Marx (Type-safe) [33]	✓	✗	✗	✓	n.a.	Memory allocator uses class hierarchy as type.
vfGuard [35]	✓	✓	✗	✓	✗	Call target resides in at least one vtable at correct offset.
T-VIP [17]	✓	✓	✗	✓	✗	*vtblptr* and random vtable entry must point to read-only memory.
VTint [48]	✓	✓	✗	✓	✗	Verifies vtable ID, vtable must be in read-only memory.
VCI [14]	✓	✓	✗	✓	✗	*vtblptr* must be statically found, in class hierarchy, or *vfGuard*-allowed.
VTPin [37]	needs RTTI	✗	✗	✓	n.a.	Overwrites *vtblptr* when object freed.
VPS	✓	✓	✓	✓	✓	Check at vcall if object was created at a legitimate object creation site.

emits the following structure:

```
mov RDI, thisptr
call [vtblptr + offset]
```

The *thisptr* is an implicit call argument, so it is moved into the first argument register, which is `RDI` on Linux x86-64 systems. Next, the `call` instruction uses the *vtblptr* to fetch the target function address from the object's vtable. The `offset` added to the *vtblptr* selects the correct vtable entry. Note that the offset is a constant, so that corresponding virtual function entries must be at the same offset in all vtables of classes that inherit from the same base class.

The same code structure holds for cases that use multiple inheritance. Depending on which (sub-)vtable the virtual function entry resides in, the *vtblptr* either points to the base vtable or one of the sub-vtables. However, if the *vtblptr* points to a sub-vtable, *thisptr* does not point to the beginning of the object, but rather to the offset in the object where the used *vtblptr* lies. Consider the example from Figure 1: when a function in the sub-vtable of class *C* is called, the call uses the *vtblptr* to its sub-vtable, and the *thisptr* points to offset `0x10` of the object. Because the code structure is the same, the program treats calls through sub-vtables and base vtables as analogous.

2.4 Threat Model: VTable Hijacking Attacks

As we explained in Section 2.3, virtual callsites use the *vtblptr* to extract the pointer to the called virtual function. Since the object that stores the *vtblptr* is dynamically created during runtime and resides in writable memory, an attacker can overwrite it and hijack the control flow at a virtual callsite.

The attacker has two options to hijack an object, depending on the available vulnerabilities: leveraging a vulnerability to overwrite the object directly in memory, or using a dangling pointer to an already-deleted object by allocating attacker-controlled memory at the same position (e.g., via a use-after-free vulnerability). In the first case, the attacker can directly overwrite the object's *vtblptr* and use it to hijack the control flow at a vcall. In the second case, the attacker does not need to overwrite any memory; instead, the vulnerability causes a virtual callsite to use a still existing pointer to a deleted memory object. The attacker can control the *vtblptr* by allocating new memory at the same address previously occupied by the deleted object.

We assume the attacker has an arbitrary memory read/write primitive, and that the $W \oplus X$ defense is in place as well as the vtables reside in read-only memory. These are standard assumptions in related work [2, 14, 43, 48]. The attacker's goal is to hijack the control flow at a virtual callsite (forward control-flow transfer). Attacks targeting the backward control-flow transfer (e.g., return address overwrites) can be secured, for example, by shadow stacks which are orthogonal to VPS and thus out of scope. Furthermore, data-only attacks are also out of scope.

2.5 Related Work on Binary-only Defenses

Here, we briefly compare our design against binary-only related work as shown in Table 1. A detailed discussion including source-level approaches is provided in Section 9.

Most existing vtable hijacking defenses assign a set of allowed target functions to each virtual callsite (e.g., *Marx VTable Protection* [33], *vfGuard* [35], T-VIP [17], *VTint* [48] and *VCI* [14]). The inaccuracy of binary analysis forces them to overestimate the target set, leaving room for attacks [38]. In contrast, VPS enforces that vtable pointers remain unmodified after object construction, ensuring that only validly created objects can be used at virtual callsites and reducing the attack surface even compared to a hypothetical defense with a perfect set of allowed targets. *Marx Type-safe Object Reuse* and *VTPin* [37] protect against the reuse of dangling pointers by modifying the memory allocator. VPS protects against dangling pointers *without* any further modification.

As the comparison in Table 1 shows, VPS combines the protection targets given by related work and additionally protects the type integrity of the object itself.

3 VTABLE POINTER SEPARATION

Our approach is based on the observation that the *vtblptr* is only written during object initialization and destruction and cannot legitimately change in between. Therefore, only the *vtblptr* that is written by the constructor (or destructor) is a valid value. If a *vtblptr* changes between the object was created and destroyed, a vtable hijacking attack is in progress. Since these attacks target virtual callsites, it is sufficient to check at each virtual callsite if the *vtblptr* written originally into the object still resides there.

Figure 2 depicts the differences between a traditional application and a VPS-protected application. The traditional application initializes an object and uses a vcall and the created object to call a virtual function. As explained in Section 2.3, the application uses the vtable to decide which virtual function to execute. If an attacker is able to corrupt the object between the initialization and vcall, she can place her own vtable in memory and hijack the control flow. In contrast, the VPS-protected application adds two additional functionalities to

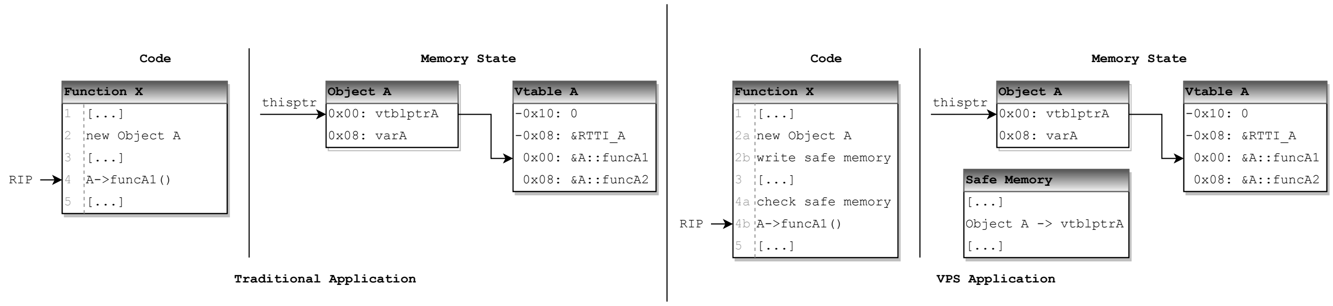


Figure 2: High-level overview of the object instantiation and virtual callsite of a traditional application (left side) and a VPS protected application (right side). For both applications the memory state is given while the instruction pointer executes the function call.

the executed code. While the object is initialized, it stores the *vtblptr* in a safe memory region. Before a vcall, it checks if the *vtblptr* in the object is still the same as the one stored for the object in the safe memory region. The vcall is only executed when the check succeeds. As a result, the same attacker that is able to corrupt the object in between can no longer hijack the control flow. The same concept holds for *vtblptr*s written in the destructors. The *vtblptr* is written into the object and used for vcalls during its destruction (if it is used at all). Since a VPS-protected application stores the written *vtblptr* into the safe memory region and checks the integrity of the one in the object if it is used at a vcall, the approach does not need to differentiate between object initialization and destruction.

In contrast to other binary-only defenses for virtual callsites [14, 17, 33, 35, 48] that allow a specific overestimated set of classes at a virtual function dispatch, VPS has a direct mapping between an object initialization site and the reachable vcalls.

Even though VPS looks conceptually similar to CFIXX, adding this protection at the binary level encounters multiple hurdles. Performing accurate analysis at the binary level is a challenging problem, especially with regards to object creation sites, where false negatives would break the protected application. Our analysis has to take direct and indirect vtable accesses into account, which do not exist on the source level. The virtual callsite identification has to be as precise as possible in order to provide a high level of security and it has to be performed without type information. Any false positive in this result breaks the application, which makes an instrumentation capable of handling these necessary (a problem that other binary-only approaches do not consider).

4 ANALYSIS APPROACH

VPS protects binary C++ applications against control-flow hijacking attacks at virtual callsites. To this end, we first analyze the binary to identify C++-specific properties and then apply instrumentation to harden it. We divide the analysis into three phases: *Vtable Identification*, *Vtable Pointer Write Operations*, and *Virtual Callsite Identification*. At a high-level, our analysis first identifies all vtables in the target binary in the Vtable Identification phase. Subsequently, the identified vtables are used to find all locations in the binary that write *vtblptr*s. Eventually, the identified vtables are also used to identify and verify vcalls in the Virtual Callsite Identification phase. While the Vtable Identification static analysis is an improved and more exact version of Pawlowski et al. [33] (finding vtables in `.bss` and GOT, considering indirect referencing of vtables), the other analyses are novel to VPS. In the remainder of this section, we explain the details of our analysis approach. Note that we focus on Linux x86-64 binaries that use the Itanium C++ ABI [16]. However, our analysis approach is conceptually mostly generic and with additional engineering effort can be applied to other architectures and ABIs as well. For architecture-specific steps in our analysis, we describe what to modify to port the step to other architectures.

4.1 Vtable Identification

To protect *vtblptr*s in objects, we need to know the location of all vtables in the binary. To find these, our static analysis searches through the binary and uses a set of rules to identify vtables. Whenever all rules are satisfied, the algorithm identifies a vtable. As explained earlier, Figure 1 shows a typical vtable structure. The smallest possible vtable in the Itanium C++ ABI [16] consists of three consecutive words (*Offset-to-Top*, *RTTI*, and *Function-Entry*). We use the following five rules to determine the beginning of a vtable:

R-1. In principle, our algorithm searches for vtables in read-only sections such as `.rodata` and `.data.rel.ro`. However, there are exceptions to this. If a class has a base class that resides in another module and the compiler uses copy relocation, the loader will copy the vtable into the `.bss` section [18]. Additionally, vtables from other modules can be referenced through the Global Offset Table (GOT), e.g., in position-independent code [1]. To handle these cases where the vtable data lies outside the main binary, we parse the binary's dynamic symbol table and search for vtables that are either copied to the `.bss` section or referenced through the GOT. Note that we do not rely on debugging symbols, only on symbols that the loader uses, which cannot be stripped.

R-2. Recall that the *vtblptr* points to the first function entry in a class's vtable, and is written into the object at initialization time. Therefore, our algorithm looks for code patterns that reference this first function entry. Again, there are special cases to handle. The compiler sometimes emits code that does not reference the first function entry of the vtable, but rather the first metadata field at offset `-0x10` (or `-0x18` if virtual inheritance is used). This happens for example in position-independent code. To handle these cases, we additionally look for code patterns that add `0x10` (or `0x18`) to the reference before writing the *vtblptr* into the object, which is necessary to comply with the Itanium C++ ABI [16]. Our algorithm also checks for the special case where vtables are referenced through the GOT instead of directly.

R-3. As depicted in Figure 1, the *Offset-to-Top* is stored in the first metadata field of the vtable at offset `-0x10`. In most cases this field is `0`, but when multiple inheritance is used, this field gives the distance between the base *vtblptr* and the sub-*vtblptr* in the object (see Section 2.1). Our algorithm checks the sanity of this value by allowing a range between `-0xFFFFFF` and `0xFFFFFF`, as proposed by Prakash et al. [35].

R-4. The RTTI field at offset `-0x8` in the vtable, which can hold a pointer to RTTI metadata, is optional and usually omitted by the compiler. If omitted, this field holds `0`; otherwise, it holds a pointer into the **`data`** section or a relocation entry if the class inherits from another class in a shared object.

R-5. Most of the vtable consists of function entries that hold pointers to virtual functions. Our algorithm deems them valid if they point into any of the **`.text`**, **`.plt`**, or **`.extern`** sections of the binary, or are relocation entries.

Abstract classes are an edge case. For each virtual function without implementation, the vtable points to a special function called *pure_virtual*. Because abstract classes are not meant to be instantiated, calling *pure_virtual* throws an exception. Additionally, the first function entries in a vtable can be `0` if the compiler did not emit the code of the corresponding functions (e.g., for destructor functions). To cope with this, Pawlowski et al. [33] allow `0` entries in the beginning of a vtable. We omit this rule because our approach can safely ignore the instantiation of abstract classes, given that *vtblptr*s for abstract classes are overwritten shortly after object initialization.

In case of multiple inheritance, we do not distinguish between vtables and sub-vtables. That is, in the example in Figure 1, our approach identifies *Vtable C* and *Sub-Vtable C* as separate vtables. As discussed later, this does not pose any limitations for our approach given our focus on *vtblptr* write operations (as opposed to methods that couple class hierarchies to virtual call sites).

The combination of multiple inheritance and copy relocation poses another edge case. In copy relocation, the loader copies data residing at the position given by a relocation symbol into the **`.bss`** section without regards to the type of the data. For classes that use multiple inheritance, the copied data contains a base vtable and sub-vtable(s), but the corresponding relocation symbol holds only information on the beginning and length of the data, not the vtable locations. To ensure that we do not miss any, we identify every `8`-byte aligned address of the copied data as a vtable. For example, if the loader copies a data chunk of `0x40` bytes to the address `0x100`, we identify the addresses `0x100`, `0x108`, `0x110`, … up to `0x138` as vtables. While this overestimates the set of vtables, only the correct vtables and sub-vtables are referenced during object initialization.

Note that on other architectures, the assumed size of 8-byte per vtable entry as used by our rules may have to be adjusted. For example, Linux on x86 (32-bit) and ARM would use 4-byte entries, with no conceptual changes.

4.2 Vtable Pointer Write Operations

The next phase of our static analysis is based on the observation that to create a new object, its *vtblptr* has to be written into the corresponding memory object during the initialization. This is done in the constructor of the class which can be either an explicit function or inlined code. The same holds for object destruction by the corresponding destructor function. Hence, the goal of this analysis step is to identify the exact instruction that writes the *vtblptr* into the memory object. This step is Linux-specific but architecture-agnostic.

First, we search for all references from code to the vtables identified in the previous step. Because vtables are not always referenced directly, the analysis searches for the following different reference methods:

(1) A direct reference to the start of the function entries in the vtable. This is the most common case.
(2) A reference to the beginning of the metadata fields in the vtable. This is mostly used by applications compiled with position-independent code (e.g., MySQL server which additionally uses virtual inheritance).
(3) An indirect reference through the GOT. Here, the address to the vtable is loaded from the GOT.

Starting from the identified references, we track the data flow through the code (using Static Single Assignment (SSA) form [12]) to the instructions that write the *vtblptr*s during object initialization or destruction. We later instrument these instructions, adding code that stores the *vtblptr* in a safe memory region. Our approach is agnostic to the location the C++ object resides in (i.e., heap, stack, or global memory). Furthermore, since we focus on references from code to the vtables, our approach can handle explicit constructor functions as well as inlined constructors and destructors.

During our research, we encountered functions with inlined constructors where the compiler emits code that stores the *vtblptr* temporarily in a stack variable to use it at multiple places in the same function. Therefore, to ensure that we do not miss any *vtblptr* write instructions, our algorithm continues to track the data flow even after a *vtblptr* is written into a stack variable. Because we cannot easily distinguish between a temporary stack variable and an object residing on the stack, our algorithm also assumes that the temporary stack variable is a C++ object. While this overestimates the set of C++ objects, it ensures that we instrument all *vtblptr* write instructions, making this overapproximation comprehensive.

4.3 Virtual Callsite Identification

Because VPS specifically protects vcalls against control-flow hijacking, we first have to locate them in the target binary. Hence, we have to differentiate between vcalls and normal C-style indirect call instructions. We follow a two-stage approach to make this distinction: we first locate all possible vcall candidates and subsequently verify them. The verification step consists of a static analysis component and a dynamic one. In the following, we explain this analysis in detail.

4.3.1 Virtual Callsite Candidates. To find virtual callsite candidates, we use a similar technique as previous work [14, 17, 35, 48]. We search for the vcall pattern described in Section 2.3, where the *thisptr* is the first argument (stored in the `RDI` register on Linux x86-64) to the called function and the vcall uses the *vtblptr* to retrieve the call target from the vtable. Note that the *thisptr* is also used to extract the *vtblptr* for the call instruction. A typical vcall looks as follows:

```
mov RDI, thisptr
mov vtblptr, [thisptr]
call [vtblptr + offset]
```

Note that these instructions do not have to be consecutive in the application, but can be interspersed with other instructions. Two patterns can be derived from this sequence: the first argument register always holds the *thisptr*, and the call instruction target can be denoted as `[[`*thisptr*`] + offset]`, where `offset` can be `0` and therefore omitted. This specific dependency between call target and first argument register is rare for non-C++ indirect calls. With the help of the SSA form, our algorithm traces the data flow of the function. If the previously described dependency is satisfied, we consider the indirect call instruction a *vcall candidate*.

Note that the same pattern holds for classes with multiple inheritance. As described in Section 2.3, when a virtual function of a sub-vtable is called, the *thisptr* is moved to the position in the object where the sub-vtable resides. Therefore, the first argument holds *thisptr* `+ distance`, and the call target `[[`*thisptr* `+ distance] + offset]`. This still satisfies the aforementioned dependency between first argument and call target. Furthermore, the pattern also applies to Linux ARM, Linux x86, and Windows x86-64 binaries, requiring only a minor modification to account for the specific register or memory location used for the first argument on the platform (`R0` for ARM, the first stack argument for Linux x86, and `RCX` for Windows x86-64).

To effectively protect vcalls, it is crucial to prevent false positive vcall identifications, as these may break the application during instrumentation. This is also required for related work [14, 17, 35, 48]. While the authors of prior approaches report no false positives with the above vcall identification approach, our research shows that most larger binary programs do indeed contain patterns that result in indirect calls being wrongly classified as virtual callsites.

A possible explanation for the lack of false positives in previous work is that most prior work focuses on Windows x86 [17, 35, 48], where the calling conventions for vcalls and other call instructions differ. That is, on Windows x86, the *thisptr* is passed to the virtual function via the `ECX` register (*thiscall* calling convention), while other call instructions pass the first argument via the stack (*stdcall* calling convention) [15]. This is not the case for Windows x86-64 and Linux (x86 and x86-64). On these architectures, the *thisptr* is passed as the first argument in the platform's standard calling convention (*Microsoft x64*, *cdecl* and *System V AMD64 ABI*, respectively). While Elsabagh et al. [14], who work on Linux x86, did not report false positives, our evaluation does show false positives in the same application set. We contacted the authors, but they could not help us find an explanation for these differing outcomes and could not give us access to the source code to allow us to reproduce the results.

4.3.2 *Virtual Callsite Verification.*

Because a single false positive can break our approach, the next phase in our static analysis verifies the virtual callsite candidates. Basically, we perform a data-flow analysis in which we track whether a *vtblptr* is used at a virtual callsite candidate. If the candidate uses the *vtblptr* to determine the call target, we consider it as verified. However, a data-flow graph alone is not sufficient to verify this connection. The control flow and actual usage of the *vtblptr* have also to be considered. Figure 3 depicts an overview of the analysis process. The following describes our analysis in detail.

Data-Flow Graphs. First, our analysis tracks the data flow backwards with the help of SSA form starting from all vtable references in the code (which create the *vtblptr*). The data flow is tracked over function boundaries when argument registers or the return value register `RAX` are involved. This means the tracking is done interprocedurally. The same data-flow tracking is done for the call target of each virtual callsite candidate. As Figure 3a shows, we obtain data-flow graphs showing the source of the data used by the vtable-referencing instructions and the virtual callsite candidates. Whenever a data-flow graph for a virtual callsite candidate has the same data source as a vtable-referencing instruction, we group them together as depicted in Figure 3b.

Control-Flow Path. Virtual callsite candidates and vtable-referencing instructions that share the same data source represent a possible connection between a created *vtblptr* and a corresponding vcall. However, this connection alone does not give any information on whether the *vtblptr* is actually used at the virtual callsite candidate. To verify this, we have to check if a control-flow path exists that starts at the data-source instruction, visits the vtable-referencing instruction, and ends at the vcall instruction. For this, our analysis searches all possible data-flow paths through the graph that start at a data-source instruction and end in a vtable-referencing instruction. Additionally, all data-flow paths through the graph are identified that start at a data-source instruction and end at a virtual callsite candidate. Then, they are split into common and unique parts as Figure 3c depicts.

Next, our analysis tries to transform these data-flow paths into a control-flow path by translating each data-flow node into the basic block that contains the corresponding instruction (see Figure 3d). With the help of the Control-Flow Graph (CFG), our analysis then searches for a path from basic block to basic block until it reaches the final block as Figure 3e shows. Eventually, if a path exists, the algorithm finds a possible control-flow path that starts from the data-source instruction, visits the vtable-referencing instruction, and ends at the vcall instruction.

Symbolic Execution. As a last step, we symbolically execute the obtained control-flow paths to track the flow of the *vtblptr* through the binary. When an instruction writes a vtable into the memory state, we replace that *vtblptr* with a symbolic value. To keep the analysis scalable to large real-world applications, our symbolic execution simply executes basic blocks without checking whether branches can actually be taken in a concrete execution. If a basic block contains a call instruction that is not part of our original data-flow path, we simply execute a return instruction immediately after the call instead of symbolically executing the called function. When the symbolic execution reaches the vcall instruction, we check the obtained memory state to verify that the *vtblptr* is used for the call

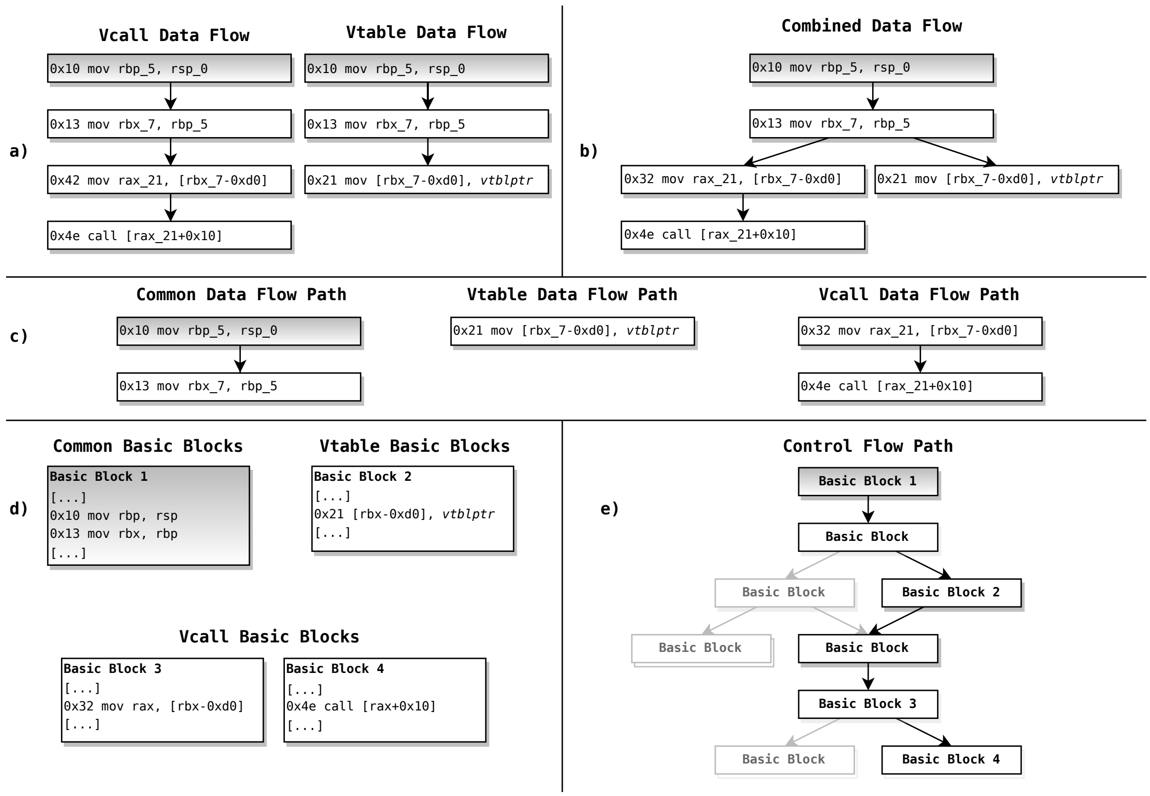


Figure 3: Data-flow and control-flow analysis of our vcall verification phase. Step *a)* shows the data-flow graph in SSA form, with the starting node in gray (data source). Step *b)* combines the data-flow graphs of *a)*. Step *c)* divides the paths through the data-flow graph into three components. Step *d)* shows the basic blocks corresponding to the data-flow paths. Step *e)* shows a path through the CFG containing all previously identified basic blocks.

target. If so, we conclude that the vcall candidate is in fact a vcall and consider it a *verified vcall*.

In addition to explicit vtable-referencing instructions, this analysis phase checks implicit vtable references as well. In case the earlier backward data-flow analysis shows that a vcall target stems from the first argument register, we check whether the calling function is a known virtual function (by checking whether the function resides in any previously identified vtable). If it is, we add a special virtual function node to the data-flow graph. We then search for a path from this virtual function node to the vcall instruction. If a path is found, we apply the steps described previously for transforming the data-flow path to a control-flow path. For such paths, before starting the symbolic execution, we add an artificial memory object containing the *vtblptr* and place the *thisptr* in the first argument register. This way, we simulate an implicit use of the vtable through the initialized object.

We perform the whole vcall verification analysis in an iterative manner. When the data-flow tracking step stops at an indirect call instruction, we repeat it as soon as our analysis has verified the indirect call as a vcall and has therefore found corresponding vtables for resolving the target. The same applies to data-flow tracking that stops at the beginning of a virtual function (because no caller is known). As soon as we can determine a corresponding vcall instruction, we repeat the analysis. The analysis continues until we reach a fixed point where the analysis fails to find any new results.

4.3.3 Dynamic Profiling. Our approach includes a dynamic profiling phase that further refines the vcall verification. During this phase, we execute the application with instrumentation code added to all virtual callsite candidates (only the vcall candidates, not the already verified vcalls). Whenever the execution reaches a vcall, the instrumentation code verifies that the first argument contains a valid *thisptr*. To verify this, we check if the first element of the object the *thisptr* points to contains a valid pointer to a known vtable (*vtblptr*). If it does, we consider the vcall verified. Otherwise, we regard the vcall as a false positive of the static analysis and discard it.

Because this phase only instruments vcall candidates identified by the static analysis described in Section 4.3.1, it is safe to assume the dependency between first argument and call instruction target. Hence, the above dynamic profiling check is sufficient to remove false positives seen during the profiling run, given that the odds of finding a C-style indirect callsite with such a distinctive pattern that uses C++ objects is extremely unlikely. We did not encounter any such case during our comprehensive evaluation. Also note, that only this dynamic analysis step discards vcall candidates as false positives. Vcalls that could not be verified by the static analysis (or not reached during this dynamic profiling) are still considered vcall candidates since the reason for the failed verification can be missing information (e.g., analysis gaps through indirect control-flow transfers).

5 INSTRUMENTATION APPROACH

VPS protects virtual callsites against control-flow hijacking attacks by instrumenting the application using the results from the analysis phase. We instrument two parts of the program: *Object Initialization and Destruction* and *Virtual Callsites*. The following describes how both kinds of instrumentation work.

5.1 Object Initialization and Destruction

We use the data collected in Section 4.2 to instrument object initialization, specifically the instruction that writes the *vtblptr* into the object. When an object is created, the instrumentation code stores a key-value pair that uses the memory address of the object as the *key* and maps it to the *vtblptr*, which is the associated *value*. To prevent tampering with this mapping, we store it in a safe memory region.

Recall that during the creation of a C++ object whose class inherits from another class, the initialization code first writes the *vtblptr* of the base class into the object, which is then overwritten by the *vtblptr* of the derived class. Our approach is agnostic to inheritance and simply overwrites the *vtblptr* in the same order (because each *vtblptr* write instruction is instrumented).

Similarly, our approach is agnostic to multiple inheritance, because object initialization sites use the address where the *vtblptr* is written as the object address. As explained in Section 2.3, at a virtual callsite the *thisptr* points to the address of the object the used *vtblptr* resides in. For a sub-vtable, this is not the beginning of the object, but an offset somewhere in the object (in our running example in Figure 1 offset `0x10`). Because this is exactly the address that our approach uses as the key for the safe memory region, our approach works for multiple inheritance without any special handling.

Since this instrumentation only focuses on *vtblptr* write instructions, it is also agnostic to object initialization and destruction. Hence, we do not have to differentiate between constructor and destructor and can use it for both.

Moreover, despite the fact that we ignore object deletion, our approach does not suffer from consistency problems. This is because, when an object is deleted and its released memory is reused for a new C++ object, the instrumentation code for the initialization of this new object automatically overwrites the old value in the safe memory region with the current *vtblptr*.

5.2 Virtual Callsites

Because a single false positive virtual callsite can break the application, we designed the vcall instrumentation code such that it can detect false positives and filter them out. In doing so, the vcall instrumentation continuously refines the previous analysis results. The vcall instrumentation consists of two components, described next: *Analysis Instrumentation* and *Security Instrumentation*.

5.2.1 Analysis Instrumentation. We add analysis instrumentation code to all vcall candidates that we were unable to verify during our static vcall verification and dynamic profiling analysis. For verified vcall sites, we only add security instrumentation and omit the analysis code.

Before executing a vcall candidate, the analysis instrumentation performs the same check as the dynamic profiling phase described in Section 4.3.3. If the check fails, meaning that this is not a vcall but a regular C-style indirect call, we remove all instrumentation from the call site. If the check succeeds, we replace the analysis instrumentation with the more lightweight security instrumentation for verified virtual callsites described in Section 5.2.2, and immediately run the security instrumentation code.

Through our use of adaptive instrumentation, our approach is able to cope with false positives and further refine the analysis results during runtime. By caching the refined results on disk, we can reuse these in later runs of the same application, improving VPS's performance over time. Furthermore, caching also improves the security of our adaptive instrumentation as we discuss in Section 8.2.

Because the analysis instrumentation verifies all remaining vcall candidates for false positives at runtime, the static vcall verification from Section 4.3.2 and the dynamic profiling from Section 4.3.3 can be omitted. Omitting these steps does not affect the correctness of our approach, although we recommend using them for optimal performance.

5.2.2 Security Instrumentation. We protect verified vcall sites against control-flow hijacking by adding security instrumentation code that runs before allowing the vcall. The instrumentation uses the *thisptr* in the first argument register to retrieve the *vtblptr* stored for this object in the safe memory region. To decide whether to allow the vcall, the instrumentation code compares the *vtblptr* from the safe memory region with the one stored in the actual object used in the vcall. If they are the same, the instrumentation allows the vcall. If not, we terminate with an alert.

6 IMPLEMENTATION

Based on the approach from Section 4, we integrated our static analysis into the open source *Marx* framework [33]. This framework provides a basic symbolic execution based on the VEX-IR from the Valgrind project [13] and data structures needed for C++ binary analysis. It is written in C++ and targets Linux x86-64 (amd64) binaries. To support integration of our approach into the *Marx* framework, we added support for SSA and a generic data-flow tracking algorithm.

Because the VEX-IR supports multiple architectures, the framework is easily extendable to these. The same is true for our approach, which is mostly independent from the underlying architecture (Section 4). To balance precision and scalability, the symbolic execution emulates only a subset of the 64-bit VEX instructions that suits our focus on vtable-centered data-flow tracking in real-world applications.

We use IDAPython [23] for vtable identification and CFG extraction. Additionally, we use instruction data provided by IDA Pro to support the SSA transformation, and use Protocol Buffers [21] to export the results in a programming language–agnostic format. We implement dynamic profiling with Pin [29]. We build the runtime component of VPS on top of Dyninst v9.3.2 [4]. Dyninst is responsible for installing *vtblptr* write and (candidate) virtual callsite hooks. We inject these wrappers into the target program's address space by preloading a shared library.

To set up the safe memory region, our preloaded library maps the lower half of the address space as a safe region at load time; this is straightforward for position-independent executables as their segments are mapped exclusively in the upper half of the address space by default. To compute safe addresses, we subtract 64 TB[2] from the addresses used by *vtblptr* writes or virtual calls. To thwart value probing attacks in the safe region, we (i) mark all safe region pages as inaccessible by default and make them accessible on demand, and (ii) use a fixed offset chosen randomly at load time for writes to the safe region. To achieve the latter, we write a random value to the `gs` register and use it as the offset for all accesses to the safe region. To mark pages as readable/writable on demand, we use a custom segfault handler that uses `mprotect` to allow accesses from our library. This means that when a *vtblptr* is written into the safe memory region and the page is not yet accessible, our segfault handler checks if the write access is done by our library and makes the page accessible if it is. Otherwise, a probing attack is detected and execution is stopped. The page remains accessible which speeds up further *vtblptr* writes to it.

We omit an evaluation of potential optimizations already explored in prior work [7, 28], such as avoiding Dyninst's penalties for (re)storing unclobbered live registers or removing trampoline code left over after nopping out analysis instrumentation code. Similarly, we do not implement hash-based safe region compression that would reduce virtual and physical memory usage and allow increased entropy in the safe region, nor do we use Intel MPK [10] to further secure the safe region. Since we focus on the exact analysis of binary applications and the subsequent instrumentation, we consider these optimizations orthogonal to our work.

7 EVALUATION

In this section, we evaluate VPS in terms of performance and accuracy. We focus our evaluation on MySQL, Node.js, MongoDB, and the fifteen C++ benchmarks found in SPEC CPU2006 and CPU2017 [39, 40]. Even though our approach is able to handle proprietary software, we evaluate it on open source software since otherwise we are not able to generate a ground truth to compare against.

7.1 Virtual Callsite Identification Accuracy

In order to measure the accuracy of the protection of VPS, we evaluate the accuracy of the vcall identification analysis. The results show that VPS, although a binary-only approach, can almost reach the same degree of protection as a source based approach. Compared to the state-of-the-art binary-only approach *Marx*, it identifies more vcalls with fewer false-positives. As applications for our evaluation, we use the C++ programs of SPEC CPU2006 and SPEC CPU2017 that contain virtual callsites, as well as the MySQL server binary (5.7.21), the Node.js binary (8.10.0), and the MongoDB binary (3.2.4). We used the default optimization levels (`O2` for CPU 2006, `O3` for all

[2]Linux x86-64 provides 47 bits for user space mappings, and 2^{47} = 128 TB.

```
    Vector<double> us[dim];
    for (unsigned int i=0; i<dim; ++i)
        us[i].reinit (dof_handler.n_dofs());
```

Figure 4: Source code snippets from `grid_generator.cc` where VTV fails to identify a virtual callsite.

others). The analysis was performed on Ubuntu 16.04 LTS running on an Intel Core i7-2600 CPU with 32 GB of RAM.

VTV. In order to gain a ground truth of virtual callsites, we use *VTV* [43] and compare against our analysis results. Since VTV leverages source code information, its results are usually used as ground truth for binary-only approaches focusing on C++ virtual callsites. All programs except MongoDB are compiled with GCC 8.1.0. MongoDB crashed during compilation and had to be compiled with the older version GCC 4.9.3. Unfortunately, compiling *450.soplex* results in a crash and it is therefore omitted. Table 2 shows the results of our vcall accuracy evaluation.

Overall, we observe that the analysis of VPS is capable of identifying the vast majority of virtual callsites in the binary. This ranges from 91.7% (*510.parest_r*) to all vcalls detected (several benchmarks). Our average recall is 97.8% on SPEC CPU2006 and 97.4% on SPEC CPU2017. With the exception of one outlier (*526.blender_r* with precision 68.3%) we have a low number of false positives, with precision ranging from 87.0% (*447.dealII*) to no false positives at all (several benchmarks). The results are similar for large real-world applications with a recall ranging from 91.8% (*MongoDB*) to 97.6% (*MySQL*) and a precision ranging from 97.2% (*Node.js*) to 99.7% (*MongoDB*). The high recall rate shows that our binary-only approach is able to protect almost as many virtual callsites as VTV does and hence provides comparable security as this source-based approach. However, it still misses some vcalls which may leave an attacker with a small room to perform an attack under the right circumstances. The precision rates show that although we have a low false positive identification rate, we still have some.

In order to cope with the problem of false positive identifications, we verify vcalls before we actually instrument them with our security check. The static analysis verification is able to verify 37.9% in the best case (*526.blender_r*) and in the worst case none. On average we verified 20.4% on SPEC CPU2006 and 18.3% on SPEC CPU2017. For large applications, the best verification rate is 12.2% (*Node.js*) and the worst 3.1% (*MongoDB*). Dynamic verification (see Section 4.3.3) considerably improves verification performance, verifying 35.1% and 25.9% for SPEC CPU2006 and 2017. Unfortunately, we were not able to execute *510.parest_r*, *MySQL* and *MongoDB* with VTV. The applications crashed with an error message stating that VTV was unable to verify a vtable pointer (i.e., a false positive). Hence, the only large real-world application with dynamic verification *Node.js* verified 20.2% of the vcalls.

A manual analysis of the missed virtual callsites (false negatives) reveals two possibilities for a miss: the data flow was too complex to be handled correctly by our implementation, or the described pattern in Section 4.3.1 was not used. The former can be fixed by improving the implemented algorithm that is used for finding the described pattern. In the latter, the *vtblptr* is extracted from the object, however, a newly-created stack object is used as *thisptr* for

Table 2: Results of our vcall accuracy evaluation. For each application this table shows (i) the code size, time needed for the static analysis (hh:mm:ss) and the ground truth generated by VTV; (ii) static vcall identification, depicting the number of indirect call instructions identified as vcall that are true positives and false positives as well as recall and precision; (iii) static vcall verification results, listing the number of verified vcall instructions, verified vcalls in percentage and verified false positives; (iv) static and dynamic verification results, showing the number of verified vcall instructions, verified vcalls in percentage, verified false positives, and the number of identified false positives removed. Cases where dynamic verification failed due to VTV false positives are in parentheses.

				Static Identification				Static Verification			Static and Dynamic Verification			
Program	*Code Size*	*Time*	*# GT*	*# TP*	*# FP*	*Recall (%)*	*Precision (%)*	*#*	*%*	*# FP*	*#*	*%*	*# FP*	*# removed*
447.dealII	4.18 MB	0:02:15	1,558	1,450	215	93.0	87.1	379	24.3	7	423	27.2	18	0
450.soplex	–	–	–	–	–	–	–	–	–	–	–	–	–	–
453.povray	1.09 MB	0:00:04	102	102	10	100.0	91.1	32	31.4	0	55	53.9	0	6
471.omnetpp	1.17 MB	0:04:00	802	800	0	99.8	100.0	245	30.6	0	530	66.1	0	0
473.astar	0.04 MB	0:00:00	1	1	0	100.0	100.0	0	0.0	0	0	0.0	0	0
483.xalancbmk	7.17 MB	5:54:25	13,440	12,915	17	96.1	99.9	2,122	15.8	0	3,792	28.2	1	0
Average [SPEC CPU2006]						*97.8*	*95.6*		*20.4*			*35.1*		
510.parest_r	12.69 MB	1:00:00	4,678	4,288	528	91.7	89.0	660	14.1	13	(660)	(14.1)	(13)	–
511.povray_r	1.20 MB	0:00:05	122	122	14	100.0	89.7	33	27.1	0	62	50.8	0	6
520.omnetpp_r	3.60 MB	0:06:57	6,430	6,190	23	96.3	99.6	1,585	24.7	0	2,286	35.6	6	0
523.xalancbmk_r	10.34 MB	15:20:40	33,880	33,069	12	97.6	100.0	1,948	5.8	0	4,961	14.6	0	0
526.blender_r	11.47 MB	0:03:29	174	172	80	98.9	68.3	66	37.9	0	70	40.2	0	49
541.leela_r	0.33 MB	0:00:01	1	1	0	100.0	100.0	0	0.0	0	0	0.0	0	0
Average [SPEC CPU2017]						*97.4*	*91.1*		*18.3*			*25.9*		
MongoDB	48.22 MB	1:57:39	17,836	16,366	44	91.8	99.7	552	3.1	0	(552)	(3.1)	(0)	–
MySQL	35.95 MB	65:57:27	11,876	11,592	179	97.6	98.5	1,330	11.2	3	(1,330)	(11.2)	(3)	–
Node.js	38.13 MB	5:16:09	12,643	12,330	353	97.5	97.2	1,538	12.2	10	2,559	20.2	45	118

Table 3: Results of our comparison against CFIXX. For each application this table shows (i) the ground truth generated by CFIXX; (ii) static vcall identification, depicting the number of indirect call instructions identified as vcall that are true positives and false positives as well as recall and precision.

		Static Identification			
Program	*# GT*	*# TP*	*# FP*	*Recall (%)*	*Precision (%)*
447.dealII	–	–	–	–	–
450.soplex	553	553	10	100.0	98.2
453.povray	110	110	11	100.0	90.9
471.omnetpp	943	942	0	99.9	100.0
473.astar	1	1	0	100.0	100.0
483.xalancbmk	12,670	12,427	527	98.0	95.9
Average [SPEC CPU2006]				99.6	97.0
510.parest_r	7,288	7,194	265	98.7	96.5
511.povray_r	119	119	11	100.0	91.5
520.omnetpp_r	6,037	6,032	71	99.9	98.8
523.xalancbmk_r	23,661	26,407	528	98.9	97.8
526.blender_r	–	–	–	–	–
541.leela_r	2	2	0	100.0	100.0
Average [SPEC CPU2017]				99.5	96.9
MongoDB	20,873	20,716	448	99.3	97.9
MySQL	13,035	12,921	380	99.1	97.1
Node.js	13,013	12,982	491	99.8	96.4

the virtual callsite which does not follow a typical C++ callsite pattern. This could be addressed by considering additional vcall patterns, at the risk of adding false positives. Given our already high recall rates, we believe this would not be a favorable trade-off.

We also verified 86 cases which VTV did not recognize as virtual callsite instructions. A manual verification of all cases show that these are indeed vcall instructions and hence missed virtual callsites by VTV. An example is given in Figure 4 for *510.parest_r*. Here, a vector is created and the function `reinit()` is invoked on line 2547. However, since the class `dealii::Vector<double>` is provided by the application and `reinit()` is a virtual function of this class, this function call is translated into a virtual callsite. We contacted the VTV authors about this issue and they confirmed that this happens because the compiler accesses the memory of the objects directly when calling the virtual function in the internal intermediate representation. Usually, the compiler accesses them while going through an internal *vtblptr* field. Unfortunately, to fix this issue in VTV would require a lot of non-trivial work since the analysis has to be enhanced.

CFIXX. Since *CFIXX* performs the enforcement in a similar way, we also evaluated our binary-only approach against this source code based method. Hence, we compiled the applications with CFIXX which is based on LLVM and extracted the protected virtual callsites as ground truth for our comparison. Table 3 shows the results of this evaluation. Unfortunately, we were not able to compile *447.dealII* and *526.blender_r* with CFIXX. As the table shows, VPS can identify on average 99.6% of all SPEC CPU2006 and 99.5% of SPEC CPU2017 virtual callsites that are also protected by CFIXX. Furthermore, VPS also yields a high precision with 97.0% for SPEC CPU2006

Table 4: Results of *Marx*'s vcall accuracy evaluation. For each application this table shows (i) the ground truth generated by VTV; (ii) static vcall identification, depicting the number of indirect call instructions identified as vcall that are true positives and false positives as well as recall and precision.

		Static Identification			
Program	*# GT*	*# TP*	*# FP*	*Recall (%)*	*Precision (%)*
447.dealII	1,558	1,307	122	83.9	91.5
450.soplex	–	–	–	–	–
453.povray	102	98	10	96.1	90.7
471.omnetpp	802	701	3	87.4	99.6
473.astar	1	1	0	100.0	100.0
483.xalancbmk	–	–	–	–	–
Average [SPEC CPU2006]				91.8	95.4
510.parest_r	4,678	3,673	295	78.5	92.6
511.povray_r	122	115	11	94.3	91.3
520.omnetpp_r	6,430	5,465	22	85.0	99.6
523.xalancbmk_r	33,880	23,541	33	69.4	99.9
526.blender_r	174	171	1,347	98.3	11.3
541.leela_r	1	0	0	0.0	0.0
Average [SPEC CPU2017]				70.9	65.8
MongoDB	17,836	12,437	1,249	69.7	90.9
MySQL	11,876	10,867	1,214	81.3	88.8
Node.js	12,643	10,648	1,095	84.2	90.7

and 96.9% for SPEC CPU2017 on average. For large real-world applications, the recall and precision rates are similar with a recall of 99.1% for *MySQL* and 99.8% for *Node.js* and a precision of 97.1% and 96.4% respectively. A manual analysis of the missed virtual callsites (false negatives) showed the same two reasons for a miss that also occurred for VTV.

Marx. A direct comparison of the accuracy with other binary-only approaches is difficult since different test sets are used to evaluate it. For example, vfGuard evaluates the accuracy of their approach against only two applications, while T-VIP is only evaluated against one. VTint states absolute numbers without any comparison with a ground truth. VCI evaluates their approach against SPEC CPU2006, but the numbers given for the ground truth created with VTV differ completely from ours (e.g., 9,201 vs. 13,440 vcalls for *483.xalancbmk*) which makes a comparison difficult. Additionally, the paper reports no false positives during their analysis which we encounter in the same application set with a similar identification technique. Unfortunately, as discussed in Section 4.3.1, we were not able to determine the reason for this. Furthermore, most approaches target different platforms than vps (Windows x86 and Linux x86) and are not open source. Since Marx is the only open source approach that targets the same platform, we analyzed our evaluation set with it. In order to create as few false positives as possible, we used its conservative mode. Unfortunately, Marx crashed during the analysis of *483.xalancbmk*. The results of the analysis can be seen in Table 4. Compared to Marx, we have considerably higher recall with better precision. Averaged over the CPU2006 benchmarks supported by Marx, vps achieves 98.2% recall (91.8% for Marx) and on CPU2017 97.4% versus 70.9%, respectively. This does not come at the cost of more false positives, as our precision is similar on CPU2006 (94.5% vs. 95.4%) and much better on CPU2017 (91.1% vs. 65.8%). For large real-world applications like *MySQL* and *MongoDB*, vps identifies 16.3% and 28.1% more virtual callsites with better precision (98.5% vs. 88.8% for *MySQL* and 99.7% vs. 90.9% for *MongoDB*).

Table 5: Object creation and destruction accuracy results, showing the number of vtable references in the code as found in the ground truth and as identified or missed by our analysis.

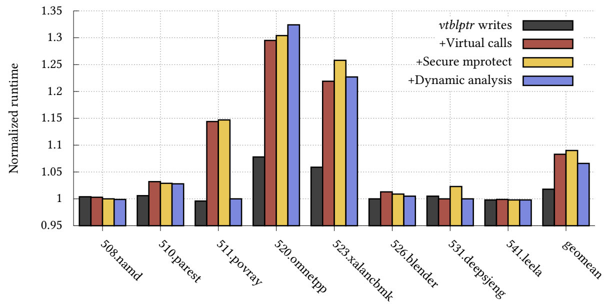
Program	*# GT*	*# identified*	*# missed*
447.dealII	–	–	–
450.soplex	102	228	0
453.povray	103	226	0
471.omnetpp	372	871	0
473.astar	0	8	0
483.xalancbmk	2,918	6,530	0
510.parest_r	12,482	25,804	0
511.povray_r	103	224	0
520.omnetpp_r	1,381	3,280	0
523.xalancbmk_r	2,790	6,323	0
526.blender_r	–	–	–
541.leela_r	87	180	0
MongoDB	8,054	11,401	0
MySQL	8,532	11,524	0
Node.js	7,816	19,204	0

Overall, our analysis shows that vps is precise enough to provide an application with protection against control-flow hijacking attacks at virtual callsites. The evaluation showed that on average only 2.5% when comparing against VTV and 0.5% comparing against CFIXX of the vcalls were missed. Since binary analysis is a hard problem, the results are very promising in showing that a sophisticated analysis can almost reach the same degree of protection as a source based approach. In addition, it shows that even source code approaches such as VTV do not find all virtual callsite instructions and can benefit from binary-only approaches such as vps. Furthermore, the number of false positives show the sensibility of our approach to handle them during instrumentation rather than assume their absence.

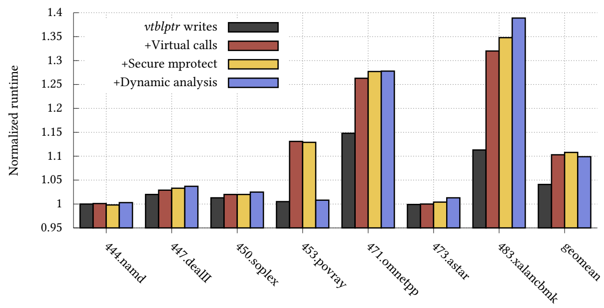
7.2 Object Initialization/Destruction Accuracy

To avoid breaking applications, vps must instrument all valid object initialization and destruction sites. To ensure that this is the case, we compare the number of vtable-referencing instructions found by vps to a ground truth. We generate the ground truth with an LLVM 4.0.0 pass that instruments Clang's internal function `CodeGenFunction::InitializeVTablePointer()`, which Clang uses for all vtable pointer initialization.

Table 5 shows the results for the same set of applications we used in Section 7.1. We omit results for *447.dealII* from SPEC CPU 2006 and *526.blender_r* from SPEC CPU 2017 because these benchmarks fail to compile with LLVM 4.0.0. The results for the remaining applications show that our analysis finds all vtable-referencing instructions. It conservatively overestimates the set of vtable-referencing instructions, ensuring the security and correctness of vps at the



(a) Microbenchmarks for SPEC CPU2017



(b) Microbenchmarks for SPEC CPU2006

Figure 5: Normalized runtime for C++ programs in SPEC CPU2006 and CPU2017, with cumulative configurations: (i) only instrument *vtblptr writes*; (ii) also instrument *virtual call instructions*; (iii) *secure the safe region* by marking all pages unwritable, and only selectively `mprotect`-ing them if they are accessed from our own instrumentation code; and (iv) include offline *dynamic analysis* results, reducing the need for hot-patching.

cost of a slight performance degradation due to the overestimated instruction set.

7.3 Performance

This section evaluates the runtime performance of VPS by measuring the time it takes to run each C++ benchmark in SPEC CPU2006 and CPU2017. We compare VPS-protected runtimes against the baseline of original benchmarks without any instrumentation. We compile all test cases as position-indepedent executables with GCC 6.3.0. For each benchmark, we report the median runtime over 11 runs on a Xeon E5-2630 with 64 GB RAM, running CentOS Linux 7.4 64-bit. We use a single additional run with more logging enabled to obtain statistics such as the number of executed virtual calls. Table 6 details our results.

Our results show the variety in properties of C++ applications. Some programs make little to no use of virtual dispatching, e.g., *444.namd*, *508.namd_r*, *531.deepsjeng_r*, and *473.astar*. Others contain thousands of *vtblptr* writes and virtual callsites, e.g., *510.parest_r* with over 12,000 *vtblptr* writes, or *483.xalancbmk* in CPU2006 with more than 1,300 verified virtual callsites. Further details are shown in the first group in Table 6.

The comparison of verified virtual calls (true positive) and regular indirect calls (false positive) shows the accuracy of our analysis. Almost all vcall candidates turn out to be real vcalls. Furthermore, with absolute numbers of executed virtual calls and *vtblptr* writes in the billions, it is clear that our instrumentation must be lightweight. The second group in Table 6 depicts the exact numbers.

The runtime overhead of our instrumentation varies from 0% for programs with little to no virtual dispatch code to 35% for the worst-case scenario (*483.xalancbmk*). In almost all cases, we see a correlation between increased overhead and number of instrumentation points (*vtblptr* writes and virtual calls). An exception is *511.povray_r*, which shows a 15% performance decrease despite a relatively low number of vcalls and *vtblptr* writes. Further inspection shows that this is caused by the 6 false positives candidate vcalls; if we disable hot-patching, our vcall instrumentation code is called over 18 billion times. While we remove instrumentation hooks for the majority of these cases, which are not real vcalls, our current implementation does not remove the Dyninst trampolines. These trampolines are the source of the unexpected overhead. The numbers depicting the comparison of the uninstrumented baseline runs to VPS-protected runs are shown in the third group in Table 6.

To better understand the overhead of VPS, we gathered detailed statistics for both SPEC CPU2006 and SPEC CPU2017 in varying configurations. We first run SPEC with only instrumentation for *vtblptr* writes enabled. In this run, the entire safe region is read-/writable and the instrumentation only (i) computes the address in the safe region to store the vtable pointer at, and (ii) copies the vtable pointer there. In the second configuration, we additionally instrument virtual calls. We check whether candidates are actual vcalls by testing the call's first argument and, if it can be dereferenced, looking this value up in the list of known vtables. We then either patch verified vcalls to enable the fast path, or remove instrumentation for false positives. The fast path fetches the vtable pointer by dereferencing the first argument, and then compares it against the value stored in the safe region. The third configuration additionally makes the safe region read-only and uses a segfault handler to mark pages writable on demand. Finally, the fourth configuration includes dynamic analysis results, removing the need to hot-patch previously verified vcalls at runtime. The results show that the majority of VPS's overhead stems from (i) *vtblptr* writes, and (ii) virtual callsite instrumentation. Figure 5 details the numbers of this evaluation.

Overall, with a geometric mean performance overhead of 11% for SPEC CPU2006 and 9% for SPEC CPU2017, VPS shows a moderate performance impact. As expected, it does not perform as well as a source-based approach such as VTV with reported 4% geometric mean for SPEC CPU2006 [43]. However, it outperforms comparable previous work (VCI with 14% [14] and T-VIP with 25% [17]) and performs slightly worse than Marx's *VTable Protection* with a reported 8% geometric mean for SPEC CPU2006, however, with better accuracy and additional type integrity.

8 DISCUSSION

This section first discusses the susceptibility of VPS to COOP attacks [38]. Next, we discuss the limitations of VPS.

Table 6: VPS performance results and runtime statistics. For each binary, this table shows (i) binary instrumentation details, depicting the number of instrumented *vtblptr* writes (*#vtblptr*), positive virtual calls (*#positive*), and candidate vcalls (*#candidates*); (ii) runtime statistics, listing the number of true positive (*#TP*) and false positive (*#FP*) virtual calls, and the total number of virtual calls (*#vcalls*) and *vtblptr* writes (*#vtblptr*); and (iii) runtime overhead, listing runtime overhead (*VPS*) compared to the baseline (*base*) in seconds.

	Binary instrumentation			Runtime statistics				Runtime overhead	
	# vtblptr	*# positive*	*# candidates*	*# TP*	*# FP*	*# vcalls*	*# vtblptr*	*base*	*VPS*
444.namd	6	0	2	0	0	0	2,018	343.5	342.9 (+ 0%)
447.dealII	4,283	161	1,459	47	0	97m	21m	289.7	299.2 (+ 3%)
450.soplex	120	195	364	48	0	1,665,968	40	215.8	220.2 (+ 2%)
453.povray	98	21	91	21	6	101,743	162	135.8	153.3 (+13%)
471.omnetpp	507	117	677	327	0	1,585m	2,156m	290.0	370.2 (+28%)
473.astar	0	0	1	0	0	0	0	350.3	351.6 (+ 0%)
483.xalancbmk	4,554	1,348	11,623	1,639	0	3,822m	2,316m	185.0	249.4 (+35%)
Geometric mean [SPEC CPU2006]									+ 11%
508.namd_r	48	0	0	0	0	0	21	271.8	271.8 (+ 0%)
510.parest_r	12,206	243	4,539	350	4	2,625m	119m	586.3	603.1 (+ 3%)
511.povray_r	113	19	121	21	6	4,577	183	498.7	572.0 (+15%)
520.omnetpp_r	2,591	447	5,310	751	0	7,958m	2,070m	507.4	661.7 (+30%)
523.xalancbmk_r	4,512	801	30,771	2,844	0	4,873m	2,314m	366.8	461.5 (+26%)
526.blender_r	43	37	174	4	46	11	3	325.8	328.6 (+ 1%)
531.deepsjeng_r	0	0	0	0	0	0	0	345.1	353.1 (+ 2%)
541.leela_r	177	0	2	0	0	0	404,208	535.5	534.6 (+ 0%)
Geometric mean [SPEC CPU2017]									+ 9%

8.1 Counterfeit Object-oriented Programming

CFI approaches targeting C++ must cope with advanced attackers using Counterfeit Object-oriented Programming (*COOP*) attacks [11, 38]. This attack class thwarts defenses that do not accurately model C++ semantics. As we argue below, VPS reduces the attack surface sufficiently that practical COOP attacks are infeasible.

For a successful *COOP* attack, an attacker must control a container filled with objects, with a loop invoking a virtual function on each object. The loop may be an actual loop, called a *main loop gadget*, or can be achieved through recursion, called a *recursion gadget*. We refer to both types as *loop gadget*. The attacker places counterfeit objects in the container, allowing them to hijack control flow when the loop executes each object's virtual function. To pass data between objects, the attacker can overlap the objects' fields.

The first restriction VPS imposes on an attacker is to prevent filling the container with counterfeit objects; because the objects were not created at legitimate object creation sites, the safe memory does not contain stored *vtblptr*s for them. An attacker has only two options to craft a container of counterfeit objects under VPS: either the program allows attackers to arbitrarily invoke constructors and create objects, or the attacker can coax the program into creating all objects needed on their behalf. The former occurs (in restricted form) only in programs with scripting capabilities. The latter scenario, besides requiring a cooperative victim program, hinges on the attacker's ability to scan data memory to find all needed objects without crashing the program (hence losing the created objects) and filling the container with pointers to these.

The second restriction is prohibiting overlapping objects (used for data transfer in COOP), since objects can only be created by legitimate constructors. As a result, a COOP attack would have to pass data via argument registers or scratch memory instead. Data passing via argument registers works only if the loop gadget does not modify the argument registers between invocations. Moreover, the virtual functions called must leave their results in the correct argument registers when they return. Passing data via scratch memory limits the attack to the use of virtual functions that work on memory areas. The pointer to the scratch memory area must then be passed to the virtual function gadgets either via an argument register (subject to the earlier limitations), or via a field in the object. To use a field in the object as a pointer to scratch memory, the attacker must overwrite that field prior to the attack, which could lead to a crash if the application tries to use the modified object.

As a third restriction, VPS's checks of the *vtblptr* at each vcall instruction limit the virtual functions attackers can use at a *loop gadget*. Only the virtual function at the specific vtable offset used by the vcall is allowed; attackers cannot "shift" vtables to invoke alternative entries. This security policy is comparable to *vfGuard* [35].

To summarize, VPS restricts three crucial COOP components: object creation, data transfer, and *loop gadget* selection. Because all proof-of-concept exploits by Schuster et al. [38] rely on object overlapping as a means of transferring data, VPS successfully prevents them. Moreover, Schuster et al. recognize *vfGuard* as a significant constraint for an attacker performing a *COOP* attack. Given that VPS raises the bar even more than *vfGuard*, we argue that VPS makes currently existing COOP attacks infeasible.

We found that multiple of the virtual callsites missed by *VTV* (as shown in Section 7.1) reside in a loop in a destructor function (similar to the *main loop gadget* example used by Schuster et al. [38]). Because the loop iterates over a container of objects and uses a virtual call on each object, COOP attacks can leverage these missed callsites as a *main loop gadget* even with *VTV* enabled. This demonstrates the need for defense-in-depth, with multiple hurdles for an attacker to cross in case of inaccuracies in the analysis.

8.2 Limitations

At the moment, our proof-of-concept implementation of the instrumentation ignores object deletion because it does not affect the consistency of the safe memory. As a result, when an object is deleted, its old *vtblptr* is still stored in safe memory. If an attacker manages to control the memory of the deleted object, they can craft a new object that uses the same vtable as the original object. Because the *vtblptr* remains unchanged, this attack is analogous to corrupting an object's fields and does not allow the attacker to hijack control. Thus, while our approach does not completely prevent use-after-free, it forces an attacker to re-use the type of the object previously stored in the attacked memory.

Another limitation of our approach lies in the runtime verification of candidate vcall sites. If an attacker uses an unverified vcall instruction, they can force the analysis instrumentation to detect a "false positive" vcall and remove the security instrumentation for this instruction, leaving the vcall unprotected. Because we cache analysis results, this attack only works for vcall sites that are unverified in the static analysis and have never been executed before in any run of the program (since otherwise only the security check is performed), leading to a race condition between the analysis instrumentation and the attacker. The only way to mitigate this issue is by improving coverage during the dynamic profiling analysis and therefore reducing the number of unverified vcalls. This is possible by running test cases for the protected program or through techniques such as fuzzing [22, 36]. Note also that this attack requires specific knowledge of an unverified vcall; if the attacker guesses wrong and attacks a known vcall, we detect and log the attack.

vps inherits some limitations from Dyninst, such as Dyninst's inability to instrument functions that catch or throw C++ exceptions and Dyninst's inability to instrument functions for which it fails to reconstruct a CFG. These limitations are not fundamental to vps and can be resolved with additional engineering effort.

Finally, we note that our safe memory region implementation—an orthogonal research topic [27] and merely a building block for vps—can be enhanced to provide stronger protection against probing attacks [20, 32]. For example, this can be done by using hardware features such as Memory Protection Keys (MPK) [10]. In the current implementation, an adversary might still be able to overwrite values in the safe memory region under the right circumstances.

9 RELATED WORK

Marx [33] reconstructs class hierarchies from binaries for *VTable Protection* and *Type-safe Object Reuse*. VTable Protection verifies at each vcall whether the *vtblptr* resides in the reconstructed class hierarchy. However, the analysis is incomplete and the instrumentation falls back to *PathArmor* [45] for missing results. Marx's Type-safe Object Reuse prevents memory reuse between different class hierarchies, reducing the damage that can be done with use-after-free. However, this approach leaves considerable wiggle room for attackers for large class hierarchies. In contrast, vps does not rely on class hierarchy information and provides stronger security by only allowing exactly correct types. Moreover, Marx only protects the heap whereas vps protects all objects.

VTint [48] instruments vtables with IDs to check their validity, but unlike vps allows exchanging the original *vtblptr* with a new pointer to an existing vtable. Moreover, VTint breaks the binary in case of false positives.

VTPin [37] overwrites the *vtblptr* whenever an object is freed to protect against use-after-free, but it requires RTTI and does not prevent *vtblptr* overwrites in general.

vfGuard [35] identifies vtables and builds a mapping of valid target functions at each vtable offset. At vcalls, it checks the target and calling convention. Unlike vps, vfGuard allows fake vtables as long as each entry appears in a valid vtable at the same offset. Further, vfGuard may break the binary in case of false positives.

T-VIP [17] protects vcalls against fake vtables, but breaks the binary when vtables reside in writable memory (e.g., in `.bss`). Moreover, unlike vps, T-VIP uses potentially bypassable heuristics.

VCI [14] only allows a specific set of vtables at each vcall, mimicking *VTV* [43]. When the analysis cannot rebuild the sets precisely, VCI falls back to vfGuard. Moreover, false positive virtual callsites in VCI break the application, as may incomplete class hierarchies (e.g., due to abstract classes [33]). In contrast, vps allows calls through any legitimately created object. Moreover, even in the hypothetical case of a perfect VCI analysis, VCI allows changing the *vtblptr* to another one in the set, unlike vps.

VTV [43] is a GCC compiler pass that only allows a statically determined set of vtables at each vcall, like most binary-only approaches [14, 17, 33, 35].

CFIXX [7] is the state-of-the-art source-based C++ defense. Like vps, it stores *vtblptr*s in safe memory and fetches them at each callsite. Given the lack of comparison against the *vtblptr* as stored in the object, CFIXX prevents but does not detect vtable hijacking. As an LLVM compiler extension, CFIXX cannot protect applications for which no source code (and LLVM compilation) is available. Therefore, proprietary legacy applications cannot be protected afterwards. While CFIXX and vps offer similar security, our binary-level analysis is completely novel. Unlike source-level analysis, our analysis must consider both direct and indirect vtable accesses. Moreover, identifying the virtual callsites for subsequent security instrumentation is challenging given the lack of type information.

10 CONCLUSION

In this paper, we presented vps, a practical binary-level defense against C++ vtable hijacking. While prior work restricts the targets of virtual calls, we protect objects *at creation time* and only allow virtual calls reachable by the object, sidestepping accuracy problems. vps improves correctness by handling false positives at vcall verification. During our evaluation, we also uncovered several inaccuracies in *VTV*, a source-based approach that is considered the state-of-the-art among C++ defenses. We release vps as open source software to foster research on this topic.

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