VTint: Protecting Virtual Function Tables' Integrity

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***Abstract*—In the recent past, a number of approaches have been proposed to protect certain types of control data in a program, such as return addresses saved on the stack, rendering most traditional control flow hijacking attacks ineffective. Attackers, however, can bypass these defenses by launching advanced attacks that corrupt other data, e.g., pointers indirectly used to access code. One of the most popular targets is virtual table pointers (*vfptr*), which point to virtual function tables (*vtable*) consisting of virtual function pointers. Attackers can exploit vulnerabilities, such as use-after-free and heap overflow, to overwrite the *vtable* or *vfptr*, causing further virtual function calls to be hijacked (*vtable hijacking*). In this paper we propose a lightweight defense solution VTint to protect binary executables against *vtable hijacking* attacks. It uses binary rewriting to instrument security checks before virtual function dispatches to validate *vtables*' integrity. Experiments show that it only introduces a low performance overhead (less than 2%), and it can effectively protect real-world *vtable hijacking* attacks.**

I. Introduction

Memory corruption bugs are still one of the most critical problems in computer security. Attackers can use these bugs to gain unauthorized access to the program state (e.g., the code and data in memory), or even corrupt the state, to indirectly control the program counter and hijack the control-flow to execute malicious code.

In the last decades, pointers to code (i.e., *control data*), such as return addresses, exception handlers and function pointers, are the most common targets to corrupt. Many defense solutions are proposed and deployed to defeat these control data corruption attacks, e.g., StackGuard [8] for return address, SafeSEH [27] for exception handlers. More generally, DEP (Data Execution Prevention [2]) or W⊕X prevents writable memory from being executed, and prevents executable memory from being written, making code corruption and code injection attacks ineffective. ASLR (Address Space Layout Randomization [37]) randomizes the locations of code and data, raising the bar of code reuse attacks such as return-to-libc [46] and ROP (Return Oriented Programming [45]).

The *vtable hijacking* attacks In addition to corrupting traditional control data, modern advanced attacks turn to corrupting other data, e.g., pointers to control data, including the virtual table pointers (*vfptr*). The *vfptr* is a hidden field of C++ objects that have virtual functions, generated by all major compilers (e.g., GCC, Visual C++ and LLVM). The *vfptr* points to a table called virtual function table (*vtable*), consisting of virtual function pointers associated with the object's class.

By overwriting a *vtable*'s contents (i.e., *vtable corruption*), or overwriting a *vfptr* to point to an attacker-crafted *vtable* (i.e., *vtable injection*), or overwriting a *vfptr* to point to some existing data (i.e., *vtable reuse*), attackers can hijack the control flow later when a virtual function is invoked. These three types of attacks form the well-known *vtable hijacking* attacks.

The *vtable hijacking* attacks have been increasing in popularity. Attackers can exploit vulnerabilities that can lead to arbitrary or specific memory write, e.g., traditional buffer overflow [42], type confusion [9, 30] and use-after-free [15], to corrupt the *vfptr* or *vtable*, and launch the *vtable hijacking* attacks. Among these vulnerabilities, the use-after-free vulnerability has become a major threat to applications. It accounts for over 80% attacks against the Chrome browser [49], and more than 50% known attacks targeted Windows 7 [25]. As far as we know, most public known real-world exploits against use-after-free vulnerabilities are *vtable injection* attacks, as this type of attack is more reliable than the other two types.

Researchers have proposed several defenses against the *vtable hijacking* attacks, such as SafeDispatch [20], a candidate GCC extension [48, 49], and VTGuard [28] (see Section III-C for more detail). These solutions, however, all need accurate type and class inheritance information of target applications, which are not available in binary executables. And thus, they cannot protect the massive legacy applications and closed-source applications. Moreover, these solutions either do not offer sufficient protection against *vtable hijacking* attacks or incur a high performance overhead.

The approach In this paper, we propose a novel solution VTint to protect binary executables against *vtable hijacking* attacks with a low performance overhead. In particular, we notice that most legitimate *vtables* are in read-only sections because no *vtables* are ever modified during runtime in legitimate programs. On the other hand, the *vtable corruption* attacks require *vtables* to be writable, and *vtable injection* attacks will inject writable *vtables* into applications at runtime. Based on this observation, we check whether *vtables* are read-only to validate their integrity at runtime. In this way, our solution VTint prevents writable *vtables* from being used in virtual calls, and can defeat all *vtable corruption* and *vtable injection* attacks. This solution is simple but effective, similar to the DEP solution that prevents writable memory from being executed.

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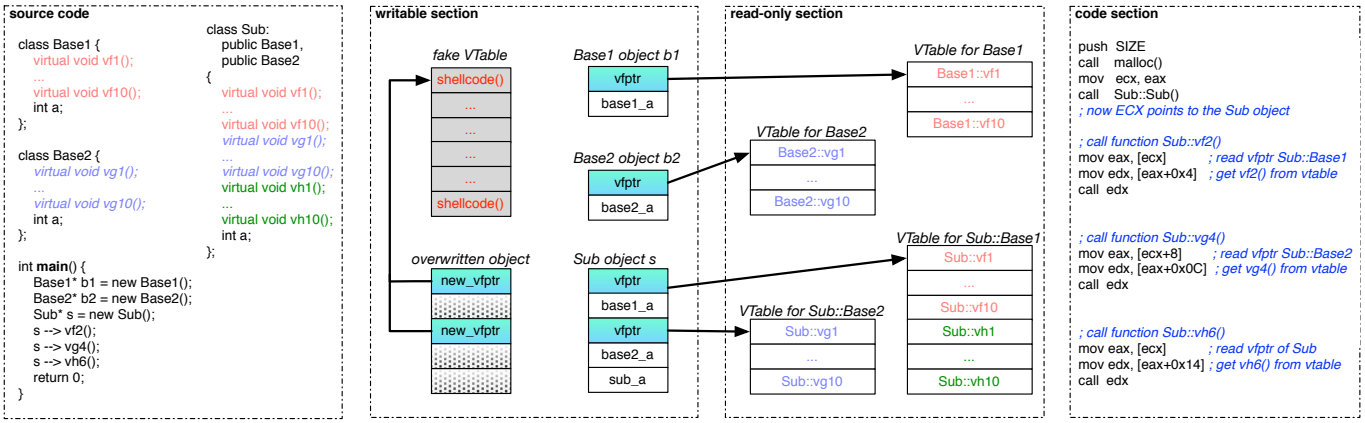


Fig. 1: The layout of C++ objects with *vtables*, and the executable code of virtual function dispatches. (a) is a sample code in which the sub-class `Sub` inherits from two base-classes `Base1` and `Base2`. (b) shows the basic layout of the object `b1` of type `Base1`, the object `b2` of type `Base2` and the object `s` of type `Sub`. There is also an `overwritten object` and a `fake VTable` controlled by attackers, illustrating a sample *vtable hijacking* attack. (c) shows the associated *vtables* of class `Base1`, `Base2` and `Sub`. (d) shows the assemble code (in Intel syntax) of the function `main` described in (a), demonstrating the process of classic object construction and virtual function dispatches.

Similar to DEP that can be bypassed by code reuse attacks, the basic version of `VTint` is also vulnerable to *vtable reuse* attacks. We thus introduce an extended version of `VTint` that will statically insert an ID for legitimate *vtables*, and dynamically check the existence of this ID before virtual function dispatches. In this way, `VTint` can defeat most *vtable reuse* attacks, e.g., attacks that overwrite *vfptr* to point to data rather than *vtables*. More importantly, unlike other secret-based solutions, `VTint` is resilient to information leakage attacks. Even if attackers know this ID (e.g., by exploiting information leakage vulnerabilities), they cannot forge a *vtable* in read-only sections to bypass `VTint`, instead only reuse existing *vtables*.

We have implemented `VTint` as a pure binary transformation. It does not depend on source code or debug information, instead only relies on relocation tables which are available due to the widely deployed ASLR on Windows platforms. `VTint` first disassembles target binary executables using our custom disassembler. Then it uses binary rewriting to deploy the mentioned security policy on the target binary, i.e., instruments *vtables* with an ID and instrument virtual function dispatches with security checks that validate whether the target *vtables* are read-only and are associated a with correct ID.

It is difficult to identify virtual function dispatches and *vtables* in binaries due to the lack of type information. We use a novel static data analysis approach to identify these structures. Furthermore, there is no efficient solution available to test whether a memory is read-only. We propose a new solution that incurs a low performance overhead, based on the Windows' Structured Exception Handling (SEH) mechanism. The performance overhead can be greatly reduced if the hardware or OS provides an easy-to-use API.

This solution can be deployed individually on each binary module (i.e., *modularity* support). The hardened module can interact with unhardened module seamlessly (i.e., *compatibility support*). Experiments show that the runtime performance overhead introduced by `VTint` is negligible (less than 2% on average), and the hardened binaries can defeat real-world *vtable hijacking* exploits.

In summary, our `VTint` protection approach has the following key advantages:

- Easy to enforce. The security policy of `VTint` is simple and easy to enforce. Similarly to DEP, it blocks all writable memory from being *vtables*, whereas DEP forbids writable memory from being executed.
- Binary compatible. By using binary analysis and rewriting, `VTint` is able to protect programs without any source code or debug information. It also offers modularity and backward compatibility support.
- Efficient. `VTint`'s performance overhead is negligible: less than 0.4% for SPEC benchmark, and less than 2% for real-world browsers.
- Effective. `VTint` can defeat all *vtable corruption* and *vtable injection* attacks by enforcing the *vtables*' integrity. It also defeats most *vtable reuse* attacks by checking the validity of *vtables*. Moreover, it is resilient to information leakage attacks.

The remainder of this paper is organized as follows: We describe the background and problem definition in Section II and III. We then give an overview of our approach and the implementation details in Section IV and V. Section VI evaluates the performance and security of `VTint`. Section VII discusses the future work and Section VIII compares our approach with the related work. Finally we conclude in Section IX.

II. Background

vtable is a general mechanism used in programming languages to support dynamic dispatch (or runtime method binding) [22], e.g., C++ virtual function calls. In this section, we briefly describe the background on how compilers use *vtable* to implement C++ virtual function calls, the related compiler calling conventions, and an important characteristic of Windows Portable Executable (PE) file format.

A. Virtual Function Calls

Virtual function is a key mechanism to support polymorphism in C++. Figure 1 shows a sample class inheritance, the associated object layout, and the final executable code regarding to the virtual function dispatches.

For each class with virtual functions, depending on the class inheritance hierarchy, the compiler will create one or more associated virtual function tables (*vtable*), as shown in Figure 1(c). In each instance of the class (i.e., object), there will be a *vfptr* pointing to the corresponding *vtable* (Figure 1(b)). If the sub-class has multiple base-classes, its object may have multiple *vfptr*. Moreover, if virtual inheritance is involved, the layout of objects is much more complex [22]. Unlike solutions that depend on class inheritance information, the `VTint` solution is independent to the layout of objects.

Figure 1(d) shows the sample executable code of virtual function dispatches. In general, *vfptr* is read out from the object first; then the target function pointer is read out from the *vtable* pointed by *vfptr*; and finally the target virtual function is called.

B. Calling Conventions

When compiling a function, a compiler needs to determine the interface to this function's caller, such as how caller functions pass arguments to this callee, who is responsible to clean the stack. There are many choices available for compilers. Each choice is called a calling convention.

For virtual function calls, a hidden argument is always passed to the callee, i.e., the pointer to the object instance (a.k.a. `this` pointer). On x86-32 processors, there are two calling conventions related to the `this` pointer: *thiscall* and *stdcall*, which are adopted by major compilers including Microsoft Visual C++ and GCC. In the *thiscall* convention, the `this` pointer is passed in the ECX register. In the *stdcall* convention, the `this` pointer is the first argument of the callee, passed by pushing on the stack. On x86-64 or ARM processors, the `this` pointer is always passed in registers.

C. Relocation Tables

To support dynamic linking and ASLR, the loader may load (i.e., relocate) an executable module's code and data to a different address other than the address assumed by the compiler. As a result, the loader should update all absolute addresses in the module at load time. Here we call the memory blocks whose contents need to be updated at load time as *relocation slots*, and call the contents as *relocation entries*.

The loader needs extra metadata to identify these relocation slots. In PE format files used by Microsoft Windows, the *relocation table* provides such metadata. The relocation table records both the offsets (e.g., $0x1000$ and $0x6000$ in Figure 2) of all memory pages that contain relocation slots, and the offsets (e.g., $0x102$ and $0x20A$) of all relocation slots within each page. The address of the relocation slot can be computed by the base address of this executable image, the page offset, and the slot offset. For example, the first relocation slot in this figure is $0x401102$ ($= 0x400000 + 0x1000 + 0x102$).

Relocation entries are usually absolute addresses of code (e.g., functions) or data (e.g., global variables). In this example,

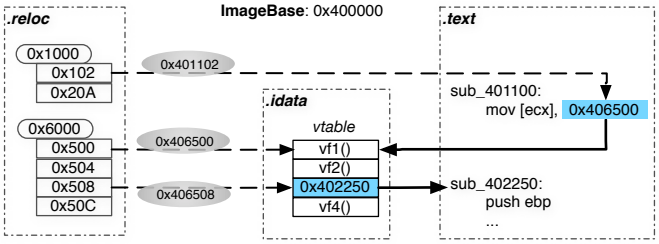


Fig. 2: A sample relocation table of Windows PE executables. The relocation entry 0x406500 at slot 0x401102 points to a *vtable*; the entry 0x402250 at slot 0x405508 points to a virtual function.

the relocation entry $0x406500$ is actually a pointer to a virtual function table. Entries in this table are all function pointers. And thus they are all relocation entries (e.g., $0x402250$) that are indexed by the relocation table too.

With relocation table, we can identify all absolute addresses in binary executables, including addresses of virtual function tables and functions whose addresses have been taken.

Note that we focus on the PE format in this paper, however our approach can be adapted to the Executable and Linkable Format (ELF) files used by Unix and Unix-like systems, since they also have a similar mechanism for relocation.

III. Problem Definition

In this section, we first describe our threat model, and then describe the attacks that `VTint` aims to defeat, and finally discuss some related defense solutions.

A. Threat Model

We assume attackers have the following capabilities: (1) attackers can read arbitrary readable memory, thus they can launch information leakage attacks and bypass secret-based defenses such as ASLR and VTGuard; (2) attackers can write to any writable memory, thus they can corrupt control data such as return addresses, and pointers to control data such as *vfptr*; (3) attackers cannot directly read from or write to registers, they can only achieve this indirectly by using existing instructions that propagate data between registers and memory. We further assume (4) attackers can leverage multiple concurrent threads to launch time-of-check-to-time-of-use attacks. Therefore, any writable memory content is untrusted even if it has been checked before it is read out again.

This assumption is strong enough because it covers most capabilities an attacker can retrieve in practice. This assumption is also realistic, because real-world attackers can exploit certain vulnerabilities to obtain these capabilities. Especially for browsers or other applications with script engines, the input data (e.g., JavaScript scripts) will be interpreted as code, or it will be Just-In-Time (JIT) compiled to native code and executed. By feeding these applications with scripts, it is easier for attackers to trigger vulnerabilities, read and write the program state, or create multiple threads to launch attacks.

On the defense side, we assume that (1) popular defenses such as ASLR and DEP are deployed; (2) no legitimate control flows in target applications will allow attackers to change the memory protection of a specified memory region, or to map an

attacker controlled file as read-only memory. If such control flows exist, the target applications should be re-designed to eliminate them, otherwise none of the current defenses can protect them from being exploited. Fortunately, we have not found any of the applications we have tested (including the Firefox and Chrome browsers) contain such behaviors. We also assume that (3) other control flow transfers except for the virtual function calls are well-protected, so they cannot be exploited to hijack the control flow before virtual function calls; (4) non-control-data attacks [6] that may lead to control-flow hijacking are out of the scope of this paper. Given these assumptions, the attackers cannot change any memory's protection attribute before any virtual function invocation.

On the defense side, we also assume that (1) the binary programs are not obfuscated; (2) the programs are compiled by major compilers that follow the traditional calling conventions, such as GCC and Visual C++. Given these assumptions, it is feasible to perform the needed binary analysis to recover the necessary information for binary instrumentation, by leveraging certain patterns and compiler conventions. Otherwise, it is still an open challenge to enable the needed binary analysis.

B. VTable Hijacking Attacks

The traditional control-flow hijacking attacks can be classified into three categories: code corruption attacks, code injection attacks and code reuse attacks. Similarly, *vtable* related attacks (denoted as *vtable hijacking* attacks) could also be classified into three categories: *vtable corruption* attacks, *vtable injection* attacks and *vtable reuse* attacks.

1) VTable Corruption Attacks: This type of attacks directly corrupts existing *vtables*' contents (i.e., virtual function pointers). Once a *vtable* is corrupted, any further virtual function call related to this *vtable* will be hijacked. This attack can succeed if and only if the *vtables* are writable.

2) VTable Injection Attacks: This type of attacks injects fake *vtables* into the applications, and overwrite *vfptr* to point to the fake *vtables*. These injected *vtables* are in writable sections, and attackers can control their contents. Among all three types of *vtable hijacking* attacks, the *vtable injection* attack is the most popular due to its high reliability.

Figure 1(b) shows a typical *vtable injection* attack. Attackers first build a fake *vtable* filled with crafted function pointers; then overwrite the *vfptr* in an object to point to this fake *vtable* by exploiting certain vulnerabilities, e.g., use-after-free. As a result, any further virtual function invocation that accesses this *vfptr* will be hijacked and the function specified in the fake *vtable* will get executed. Advanced attack techniques such as ROP could be used here to launch reliable attacks.

For example, attackers may first fill the fake *vtable* with pointers to some ROP gadgets, such as a gadget (`xchg eax, esp; ret;`). This gadget changes the stack to a memory controlled by attackers (i.e., pointed by `eax`), and guides further ROP gadgets to get executed. Then, if there is a use-after-free vulnerability in the program, attackers may try to allocate a memory block taking the exact address as the original freed object (e.g., object `s` in Figure 1(b)), in order to overwrite the *vfptr*. Then, attackers can hijack the control flow when this freed object's virtual function is called.

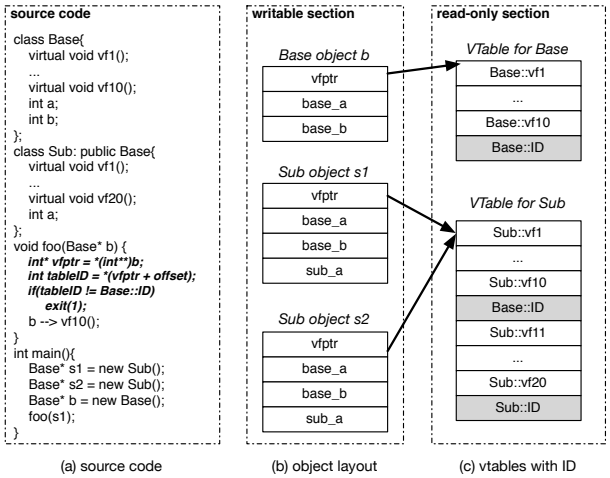


Fig. 3: The VTGuard defense solution. (a) is a sample code snippet including a sub-class derived from the base-class. The code in bold font is instrumented by VTGuard to match a secret cookie (i.e., ID) in target *vtable*. (b) shows three objects of type `Base` and `Sub`. (c) shows the associated *vtables* instrumented with the secret cookies.

3) VTable Reuse Attacks: This type of attacks also overwrites *vfptr* like the *vtable injection* attacks. The corrupted *vfptr*, however, do not point to attacker-crafted *vtables*, but point to existing *vtables*, or other existing code or data in memory which are not controlled by attackers.

Unlike the *vtable injection* attack, it is hard to reliably launch the *vtable reuse* attack in practice, because there are few existing *vtables* or data that can be used for attack purposes. As far as we know, there are no publicly known *vtable reuse* attacks.

C. Existing Solutions

In this section, we introduce three state-of-the-art defenses against *vtable hijacking* attacks and compare them with our solution. The comparison is summarized in Table I.

1) VTGuard: VTGuard [28] is a defense solution deployed in Internet Explorer (IE) browsers. As shown in Figure 3, the basic idea is to insert a secret cookie into each *vtable*. Before each virtual function dispatch, VTGuard inserts a security check to examine the corresponding secret cookie. If the cookie in the *vtable* does not match the expected secret, a security violation is detected and the program is terminated. This solution can effectively mitigate certain *vtable reuse* attacks, but not *vtable corruption* and *vtable injection* attacks.

TABLE I: Comparison between defense solutions against *vtable hijacking* attacks, including whether they can defeat *vtable hijacking* attacks, whether they are resilient to information leakage attacks, whether they support binary programs, and the performance overhead comparison. The abbreviation `SD` here stands for SafeDispatch.

defense solution	*vtable hijacking*			info leakage	binary support	perf. overhead
	corrupt	inject	reuse			
VTGuard	N	N	Y	N	N	0.5%
SD-vtable	N	Y	Y	N/A	N	30%
SD-method	Y	Y	Y	N/A	N	7%
DieHard	partial	partial	partial	N/A	N	8%
VTint	Y	Y	partial	Y	Y	2%

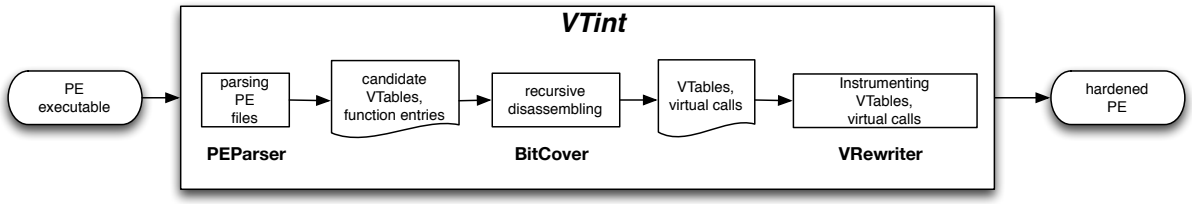


Fig. 4: Architecture of VTint. It first parses the target binary with PEParser, then disassembles the binary with BitCover, and then finally rewrites the binary to instrument *vtables* and instrument virtual function dispatches with security checks using VRewriter.

For example, by launching an information leakage attack, attackers can read the expected secret cookie from the security check instruction before the virtual function dispatch. Then, it can forge a *vtable* and put the secret cookie at the proper location inside the fake *vtable*. In this way, the fake *vtable* can pass the security check and the control-flow will be hijacked.

2) SafeDispatch: SafeDispatch [20] is a compiler-based defense solution against *vtable hijacking* attacks. It first makes a whole program Class Hierarchy Analysis, to infer the valid *vtable* set and valid virtual function set that may be used for each virtual function dispatch. Then it inserts security checks before each virtual function dispatch.

There are two types of security checks that may be inserted by SafeDispatch: *method checking* and *vtable checking*. The first approach will check whether the target virtual function is in a valid method set, and the second approach will check whether the target *vtable* is in a valid table set.

Without runtime profile information, the average overhead of the method checking solution is about 7%, and the *vtable* checking solution's overhead is about 30%. In addition, the later solution cannot defeat *vtable corruption* attacks.

3) DieHard: DieHard [4, 34] provides a custom memory allocator to randomize and separate the memory allocation. In this way, it can provide a probabilistic guarantee against vulnerabilities such as heap overflow and use-after-free, and thus can protect *vtable* and *vfptr* from being overwritten in some cases. Its average overhead is about 8%.

In comparison, our solution VTint directly protects binaries with a low performance overhead. It can defeat all *vtable corruption* and *vtable injection* attacks, and most *vtable reuse* attacks. It is also resilient to information leakage attacks.

IV. The VTint Design

In this section, we describe the design of VTint: starting with the security policy we deploy to defeat *vtable hijacking* attacks, followed by an overview of the VTint workflow, and finally the design of each component.

A. Security Policy

The basic security policy enforced by VTint is simple and effective. Similar to the widely deployed DEP, it enforces that all *vtables* are read-only, prohibiting *vtable injection* and *vtable corruption* attacks. It, however, also faces a similar challenge as DEP. Attackers can launch *vtable reuse* attacks, e.g., by overwriting the *vfptr* to point to existing read-only *vtables* or other data or code in read-only sections, although there are no publicly known exploits that work in this way.

Therefore VTint also enforces another security policy. It separates *vtables* from other data or code in read-only sections, and enforces that all virtual function calls' targets are legitimate *vtables*, not other data. In particular, VTint moves *vtables* to a read-only section, and instruments them with a special ID. At runtime, the virtual function dispatches' target *vtables* are validated against this ID. As a result, attackers can only reuse existing *vtables*, not any other data in read-only sections. Most *vtable reuse* attacks are thus defeated.

B. Overview of VTint

To deploy this security policy on binary executables, VTint uses the binary rewriting technique. In general, it parses and disassembles the target binary first, and then identifies all *vtables* and virtual function dispatches in this binary, and finally instruments these *vtables* and virtual function dispatches with security checks to enforce aforementioned security policy. Figure 4 describes the overall workflow.

PEParser first parses the target executable. All basic information, including sections, the export table and relocation table, are retrieved from the binary. The executable binary's EntryPoint, relocation entries and entries in export table, together form the set of candidate function entries.

Given the candidate function entries, our custom disassembler BitCover conducts a recursive disassembly and identifies all functions, instructions and control tables such as jump tables. This disassembler is based on our previous work [52]. We have introduced some new forward and backward data-flow analysis to BitCover, to identify *vtables* and virtual function dispatches, in order to support VTint.

Given the *vtables* and virtual function dispatches, the last component VRewriter rewrites the original binary program. More specifically, it moves all identified *vtables* to a read-only section, and instruments them with a special ID to differentiate them from other code or data in read-only sections. Then it inserts security checks before each virtual function call to enforce the aforementioned security policies, i.e., to check whether it is read-only and whether its ID is correct.

Finally, a hardened PE format binary executable is generated. This hardened executable is immune to *vtable corruption*, *vtable injection* attacks, and most *vtable reuse* attacks. Furthermore, it is resilient to information leakage attacks.

C. Binary Parsing

As described earlier, PEParser will parse the target PE binary, retrieve all basic information, and build a set of candidate function entries to guide further recursive disassembling.

Moreover, it will also identify all candidate *vtables*. Addresses of *vtables* will be used in classes' constructor functions, and thus they must be relocation entries. Furthermore, entries in *vtables* are all virtual function pointers, and thus they are also relocation entries. So, by examining all relocation entries via the relocation table, we can identify a candidate *vtable*, if and only if we find a relocation entry which points to an array, and this array is composed of (maybe only one) other relocation entries which point into code sections. In this way, `PEParser` is able to find out all candidate *vtables*.

D. Disassembling and Identification

`BitCover` disassembles the given binary and identifies all *vtables* and virtual function call sites in it.

1) Disassembling: `BitCover` takes the original PE executable file as input, as well as the candidate function entries produced by the previous component `PEParser`, and generates the target binary's disassembly information.

In general, the workflow of `BitCover` consists of two phases. First, it greedily explores all the code and data in the program, looking for function entries. For each candidate function entry, `BitCover` starts a recursive disassembling from it until a stop condition is met. For example, it will stop when an invalid instruction is met, or the new instruction overlaps with some previous resolved instructions. Once `BitCover` meets an invalid byte sequence, it will mark the code entry from where it starts disassembling as invalid, and continues disassembling the next candidate function entry. In the second phase, `BitCover` refines the disassembling result. In particular, it will mark some code entries as reliable code entries based on some rules. Then, it propagates the reliability across code entries based on the call graph. Finally it removes unreachable code entries. More details of the disassembler are discussed in our previous work [52], including other stop conditions, and other control data tables' identification.

2) Identifying VTables: After disassembling, we are able to filter out all *vtables* from the candidate *vtables* provided by `PEParser`. For each *vtable*, it will be used in the associated class's constructor function, and its pointer will be assigned to the generated object's *vfptr*. Based on this observation, we propose a new forward data-flow analysis to identify these *vtable* assignment operations, and filter real *vtables*. More implementation details are discussed in Section V-B.

3) Identifying Virtual Function Call Sites: As Section II-A describes, to dispatch a virtual function call, *vtables* will be indirectly accessed to get the address of a target virtual function. The workflow is that, a *vfptr* is read out from an object first (i.e., a memory read operation), and then a function pointer is read out from the table pointed by the *vfptr* (i.e., another memory read operation), finally the function pointer is indirectly jumped to or called.

Based on this general pattern, we can identify all candidate virtual function call sites. Combining with other rules, such as the propagation of the object pointer (i.e., `this` pointer) from caller to callee functions, we can further filter out all valid virtual function call sites. In this process, we use several backward data-flow analyses to help identify the propagation of pointers to objects and pointers to *vtables*. More implementation details are discussed in Section V-C.

E. Rewriting and Instrumentation

Based on the output of `BitCover`, the next component `VRewriter` will rewrite the original binary to enforce our security policies. First, it creates a read-only section, and moves all identified *vtables* to this section, and instruments IDs for *vtables* in this section. Then, all references to *vtables* will be updated to the *vtables*' new addresses. Finally, `VRewriter` will instrument security checks for each virtual function dispatch to validate the target *vtable*'s integrity (i.e., whether it is read-only and contains a correct ID). The instrumented security enforcement code will be put into a new code section, leaving most of the original code section intact.

1) VTable Instrumentation: First, we will move all identified *vtables* to a new read-only section, in case that some legitimate *vtables* locate in writable sections. To move a *vtable*, we need to know its size. However, it is challenging to identify the exact size of a *vtable* in a binary. We conservatively move several more bytes around *vtables* to the new section. More implementation details are discussed in Section V-D1.

Then we instrument an identical ID at the beginning of each page in the new *vtable* section. This ID is randomly selected, and is different from any words at the beginning of any page in any exiting read-only sections. In this way, we only need to instrument an identical ID for a small number of pages, and do not need to modify *vtable*' layout. More implementation details are discussed in Section V-D2.

2) Virtual Call Instrumentation: As Figure 1 shows, when a virtual function is invoked, the corresponding *vtable* will be implicitly accessed to get the target virtual function. `VTint` will validate the property of the *vtable* before it is accessed, to enforce the aforementioned security policy.

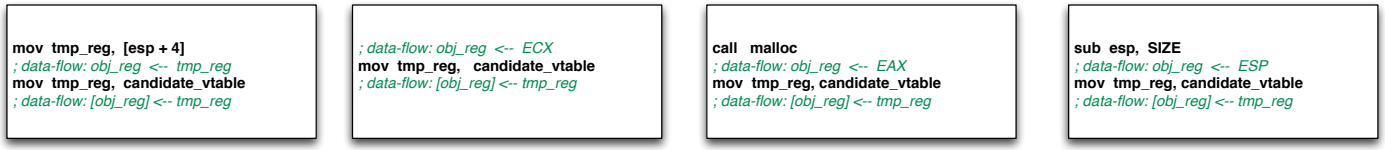
In particular, before the *vtable* is accessed, `VTint` will check whether it is read-only, and whether there is a legitimate ID associated with this table. If either condition is not satisfied, the target *vtable* is a vulnerable table that can be exploited, and thus will be rejected by the security check. Once a *vtable* is rejected, a warning can be raised and the program can be blocked. More implementation details are discussed in Section V-E.

F. Modularity and Compatibility Support

For any single module, `VTint` can identify all of its *vtables* and virtual function dispatches, only using information collected from the module itself. In addition, the security checks instrumented for this module also work independently. As a result, `VTint` can be applied to a single module without any problem, i.e., it has a perfect modularity support.

Further, if `VTint` hardens all modules, there will be no compatibility issues. For legitimate virtual function calls, the instrumented security checks should not fail, because all legitimate *vtables* have been put in read-only sections and instrumented with an identical ID.

In some cases, however, we cannot harden all modules at the same time. For example, some modules belonging to the operating system cannot be altered by user space applications. For these unhardened modules, there are no IDs instrumented for their *vtables*. As a result, when these *vtables* are accessed in



(a) the object instance is from its caller by argument (stdcall)

(b) the object instance is from its caller by register ECX (thiscall)

(c) the object instance is allocated by itself in the heap (inline allocating)

(d) the object instance is allocated by itself on the stack (inline allocating)

Fig. 5: Constructor functions. Generally, the constructor first accepts a memory block from its caller, e.g., (a) by the first argument or (b) by the register ECX; or allocates a memory block by itself, e.g., (c) by allocating memory in the heap or (d) by reserving space on the stack. Then the constructor function copies this class's *vtables* into proper locations inside this memory block (i.e., the object instance's memory region).

the hardened modules, the security checks instrumented before virtual function dispatches will fail and cause false positives.

To solve this problem, `VTint` also deploys a fail-safe check to eliminate these false positives. In this fail-safe check, only the basic security policy will be enforced, i.e., only read-only property of *vtables* will be checked. In this way, it is still able to defeat all *vtable corruption* and *vtable injection* attacks, and will not generate false positives even for unhardened modules that do not have IDs.

Note, in some very rare cases, the unhardened modules' *vtables* are writable, and cannot pass our fail-safe check. We believe this type of modules is extremely dangerous because attackers can directly corrupt these *vtables*. So we do not handle this type of incompatibility, and will advise the administrator/user to harden such modules immediately.

V. Implementation

In this section, we describe the detail of `VTint`'s prototype implementation, including the identification of some key structures (e.g., *vtables*, constructor functions, and virtual function dispatches) in binary executables, and the implementation of the security checks. It is worth noting that, in our current prototype, we focused on Windows PE binaries on x86 platforms compiled from C++ language; but the techniques we developed for `VTint` is general, and can be extended to handle other executables on other platform.

A. *Identifying Constructor Functions*

Every C++ object will be initialized explicitly or implicitly by a constructor function. The general workflow of a constructor function is shown in Figure 5. The object to be initialized can be passed in from the constructor's caller, or be directly allocated in the constructor itself.

In the first case, the constructor's caller has already prepared the object's memory, and passes in the `this` pointer. As described in Section II-B, for `stdcall`, the `this` pointer is pushed on stack (Figure 5(a)); and for `thiscall`, the `this` pointer is passed through the `ECX` register (Figure 5(b)).

In the second case, the constructor will allocate the object directly, by calling memory allocation function, such as `malloc` or `new` (Figure 5(c)), or by reserving the space on stack (Figure 5(d)).

During the initialization, if the object's class contains virtual function(s), the constructor function will assign associated *vtables* to proper locations inside the object's memory. In particular, the first word of this object's memory must be a *vfptr*. If this class has multiple base classes, then there may be other *vtable* assignments to this memory block.

Based on this observation, we are able to identify constructor functions as follows. For each candidate *vtable* (i.e., `candidate_vtable` in Figure 5) identified by the `PEParser` component of `VTint`, we locate all assignment operations that assign the candidate *vtable* to registers or memory (i.e., `tmp_reg` in the figure).

Then, for each of such assignments, `VTint` uses a forward data-flow analysis to find out the final target memory of this assignment. If a target memory of form `[register+offset]` is found, and no other accesses to the candidate *vtable* (e.g., reading an entry of this table) are made before assigning it to this memory, we mark the register (i.e., `obj_reg` in the figure) used in this memory as a candidate pointer to an object.

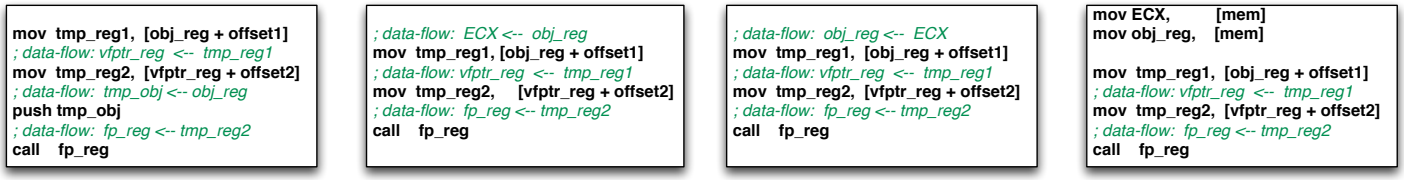
Finally, `VTint` makes a backward data-flow analysis to locate the source of this object register. If the source of this object register is (1) the first argument of the current function, e.g., Figure 5(a), (2) the register `ECX` that is passed from the caller function, e.g., Figure 5(b), (3) the return value of memory allocation function such as `malloc`, `new` or user-defined functions, e.g., Figure 5(c), or (4) an address on the stack, e.g., Figure 5(d), then this candidate object register very likely points to an object, and the current function is a constructor function.

B. *Identifying* vtables

After identifying all constructor functions, real *vtables* can be filtered out from the candidate *vtables* identified by `PEParser`. In particular, for each candidate *vtable*, it is a real *vtable* if and only if all references to it are assignment operations in constructor functions that assign them to *vfptr*.

For each candidate *vtable*, we first find out all its references by examining all relocation entries. Then, for each reference, we apply a static forward data-flow analysis to recognize the propagation of this pointer. If and only if the pointer is assigned to a memory (e.g., *vfptr*) before all other accesses to this pointer (e.g., read an entry from the candidate table) are made, this reference is marked as a `vtable assignment`. Finally, if and only if all of these references are *vtable* assignments, this candidate *vtable* is marked as an actual *vtable*.

For these real *vtables*, their sizes are still not clear. Since it is challenging to recover the exact size of a *vtable* from a binary, we conservatively compute an upper bound of the size. We scan the words (4-byte) from the beginning of the *vtable* and increase the size one by one, and stop counting the size if one of the following conditions is met. First, if a word



(a) explicit object assignment by argument (stdcall)

(b) explicit object assignment by ECX (thiscall)

(c) implicit object assignment by ECX (thiscall)

(d) implicit object assignment by ECX (thiscall)

Fig. 6: Virtual function dispatches. In general, the *vfptr* is first read from the object; and then the function pointer is read out from the *vtable*. The object instance is either passed to the callee (a) explicitly through the first argument, or implicitly by ECX register. When using ECX, there are three cases: (b) the ECX is from the object register, (c) the object register is from ECX, or (d) they are from a same memory.

scanned is not a relocation entry, then it does not belong to the *vtable*. Second, if a word scanned is directly referenced by some code or data, then it is not an entry of the *vtable* since virtual function pointers will only be indirectly accessed.

C. Identifying Virtual Function Dispatches

As described in Section II-A, a virtual function dispatch has at least three steps. First, the *vfptr* that points to a *vtable* is read out from the target object. Then, the target function pointer is read out from the *vtable*. Finally the control-flow transfers to the target function by an indirect call or jump instruction. In most cases, the last two steps can be done in one instruction because the call and jump instructions on x86 platforms support memory operands.

Before introducing the process of identifying virtual function dispatches, we first define a notion `definition point`. For a register, its definition point is a memory access operation that retrieves the value from the memory and finally propagates it to this register. A special case is that, if the target register is EAX, its definition point could also be any function invocation instruction, because for Intel x86 ABI, EAX always holds the return value of a function invocation.

By analyzing the definition points of registers used in instructions, especially in call/jump instructions, we are able to identify candidate virtual function calls. First, all indirect call/jump instructions (e.g., `call fp_reg` in Figure 6) are extracted from the disassembling result. Then, for each indirect call/jump instruction, `VTint` makes a backward data-flow analysis to locate the definition point of the target function (i.e., `fp_reg` in Figure 6) of this call/jump. If the definition point is found, and there is another register used in this memory access, this new register is thus a candidate *vfptr* (i.e., `vfptr_reg` in Figure 6). Next, `VTint` makes another backward data-flow analysis to locate the definition point of the `vfptr_reg` register. If a definition point is found, and there is a register used in this memory access, this register is thus a candidate `this` pointer register (i.e., `obj_reg` in Figure 6). Now, a candidate virtual function dispatch is identified.

Nevertheless, not all function calls that match the above pattern are real virtual function dispatches. For example, if a structure member is a pointer to a function pointer, the corresponding code to invoke the target function will also match this pattern. Fortunately, in addition to this basic pattern, the callers of virtual function calls need to pass a special argument, i.e., the `this` pointer, to the callees.

As described in Section II-B, there are two calling conventions adopted by major compilers for passing `this` pointers, i.e., `thiscall` and `stdcall`. For `thiscall`, the `this` pointer is passed to the callee implicitly through ECX register; and for `stdcall` it is passed explicitly through the first argument. So there must be a data-flow between the candidate `this` pointer and ECX register, or the first argument. Considering this data-flow, there are four general cases of a virtual function dispatch, as shown in Figure 6.

Taking this additional data-flow into consideration, we can further remove invalid candidate virtual function dispatches identified earlier. Starting from the indirect function call or jump instructions (e.g., `call fp_reg` in Figure 6) in these candidate virtual function dispatches, `VTint` performs a backward data-flow analysis to identify all possible `this` pointers. If one of the following conditions is met, the identified virtual function dispatch is marked as real.

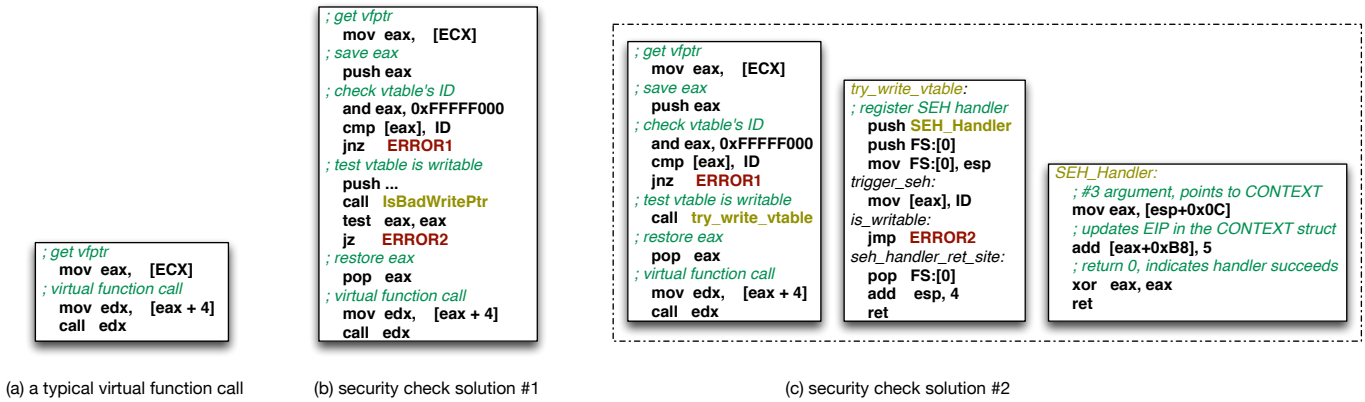
- The first argument pushed for the callee is from the candidate `this` register (`obj_reg` in Figure 6(a)). It is worth noting that, push operations are not the only way to pass arguments. For example, the instruction `mov [esp], arg` also prepares the argument for the callee. `VTint` also covers this case.
- At the call site of the indirect call, the value of ECX register is from the candidate `this` pointer (`obj_reg` in Figure 6(b)).
- The `this` pointer (`obj_reg` in Figure 6(c)) is from ECX register, and the value of ECX has not changed between the load of *vfptr* and the call site.
- The candidate `this` pointer (`obj_reg` in Figure 6(d)) and ECX both come from the same memory.

D. Instrumenting vtables

To enforce the security policy, we need to instrument *vtables* first, i.e., put them in read-only sections, and differentiate them with other code or data in read-only sections.

1) Moving vtables: We will first create a read-only section (denoted as `vtsec`) and copy all identified *vtables* to this section. So, even if the original *vtable* is writable, the new copy will be set to read-only. During this copy operation, we will conservatively copy more bytes, to make sure all *vtable* related accesses can work without any problem.

As far as we know, there are two other types of *vtable* related metadata that may be accessed by the program. First, there may be a RTTI (RunTime Type Information) pointer for some C++ objects (e.g., operands of `dynamic_cast<>`).



(a) a typical virtual function call (b) security check solution #1 (c) security check solution #2

Fig. 7: Illustration of VTint's security check solution. (a) describes the original virtual function dispatch which needs to be checked. (b) is a basic solution which first check the ID of the target *vtable*, and then test whether the *vtable* is writable by calling existing system APIs such as IsBadWritePtr. (c) is an optimized solution to check the ID and test the memory property, by utilizing Windows' Structured Exception Handling (SEH) mechanism.

This pointer usually locates right before the *vtable*, and is used for runtime introspection. Second, there may be some offset adjustors before the *vtables* used for adjusting this pointer before invoking a virtual function. For each virtual function in this *vtable*, there is at most one this adjustor. So, for a *vtable* of size N, there may be at most 4+4*N bytes metadata before the *vtable*. It is worth noting that, all this adjustors are not relocation entries, so we can reduce the size of the metadata if a relocation entry is found within this metadata region.

As mentioned in Section V-B, PEParser is able to get an upper bound of the count of virtual functions in a *vtable*. Suppose this upper bound is N, we will copy this *vtable* and its candidate metadata (at most 4+4*N bytes) before the *vtable* to the new section vtsec.

2) Instrument IDs*:* Then we will instrument *vtables* with a special ID, to differentiate them from other data or code in read-only sections. To instrument IDs, we need to answer two questions: where to instrument, and what IDs can we use?

As shown in Figure 3, the VTGuard solution instruments the secret cookies (i.e., IDs) at the end of *vtables*. It will thus change the layout of *vtables*. For binary programs, we can hardly get the class hierarchy information, and we cannot instrument IDs in this way. Similarly, we cannot instrument IDs before *vtables*, as there may be some metadata. Instead, we instrument an ID at the beginning of each page in the new *vtable* section (i.e., vtsec). In this way, we do not modify the layout of *vtables*, introducing no compatibility issues.

We propose a novel ID selection solution. More specifically, we randomly select an ID that is different from any words at the beginning of any page of any existing read-only sections. Because the count of such words is limited (i.e., smaller than page count), there are many candidate IDs for use. Unlike other secret-based solutions, this solution is resilient to information leakage attacks. Even if attackers know the ID we are using, they cannot forge a read-only page with this ID, and can only reuse pages in our new *vtable* section.

E. Instrumenting Security Checks

After identifying all *vtables* and virtual function dispatches, VTint inserts security checks before the virtual function calls to validate the *vtable*'s integrity. As shown in Figure 7, the security checks will first match the ID of the target *vtable*. Then, it will test whether the *vtable* is read-only. If any one check fails, the control flow will transfer to pre-defined error handlers, blocking the program from further execution.

It is worth noting that, there is no efficient solution available to check whether a memory location is writable (or readable, executable) on x86 platforms. In Windows, a program is able to get all its loaded modules' memory map information by traversing a special data structure in the Process Environment Block (PEB). In Linux, it is possible to get the memory map information from the /proc file system. In practice, however, both solutions are too slow.

Some APIs provided by the operating system can also be used to test whether a memory is writable. For instance, the IsBadWritePtr (Figure 7(b)) in Windows is a qualified API. But this function is inefficient, and will break the program state in some cases. The read() function in Linux, which reads file contents to the target memory, can also be used to test whether the target memory is writable. The performance overhead, however, is also high.

VTint deploys another solution to check whether the target memory is read-only, by utilizing the Windows' Structured Exception Handling (SEH) machanism. The function try_write_vtable in Figure 7(c) will check if the *vtable* is writable. In particular, this function first registers a special SEH exception handler on the stack. Then, it tries to write the target *vtable* with the ID (i.e., trigger_seh in the figure).

If the *vtable* is writable, this write operation will succeed, and the following instruction jmp ERROR2 will be executed. This error handler terminates the application and never returns to the original virtual function call that is vulnerable.

Otherwise, this write operation will trigger an exception, and the exception handler (i.e., SEH_handler in Figure 7) registered in this function will be automatically invoked by the operating system. In this exception handler, we only increase the EIP register in the program state record (i.e., a CONTEXT structure) by 5 and then return. The return value of this handler is set to 0, instructing the operating system to restore the program state and transfer the control flow to the user application.

TABLE II: Analysis results of the SPEC2000 and SPEC2006 benchmark applications. Only applications having virtual functions are listed here. The second column SPEC describes which SPEC subset the application is from. The third column source LOC counts the applications' line of source code. The following column shows the analysis time that VTint takes to disassemble and rewrite this application's binary. The file size group of columns shows the original file size, the hardened binary's size and the size overhead. The VTbale info group of columns shows the statistics of information related to virtual tables, including the count of instructions, the count of virtual tables and the count of virtual function dispatches. The last group of columns describes the runtime performance of the original SPEC applications and the hardened ones, and the performance overhead is then computed. The geometry mean value of these performance overheads is about 0.37%.

App	SPEC	source LOC	analysis time (sec)	file size (KB)			VTable info			Run Time (sec.)		
				orig	new	size overhead	#inst	#vtables	#vcalls	orig	new	perf. overhead
252.eon	spec2000int	41,188	5.9	572	605	5.81%	135,253	68	205	33.8	34.8	2.96%
444.namd	spec2006fp	3,886	5.5	532	544	2.20%	127,593	3	2	989	979	-1.0%
447.dealII	spec2006fp	94,384	12.3	1,639	1,684	2.70%	360,153	115	200	2,172	2,180	0.37%
450.soplex	spec2006fp	28,277	6.7	516	551	6.72%	210,347	51	495	556	560	0.72%
453.povray	spec2006fp	78,705	136.3	1,170	1,221	4.34%	226,923	48	112	440	429	-2.5%
471.omnetpp	spec2006int	19,991	9.2	811	910	12.20%	242,166	127	1,431	218	221	1.38%
473.astar	spec2006int	4,280	1.5	119	119	0.03%	41,710	2	0	292	292	0
483.xalancbmk	spec2006int	267,399	211.6	3,767	4,030	6.97%	872,069	29	4,248	179	181	1.12%

As a result, the instruction jmp ERROR2 (5 bytes) will be skipped, and the testing function try_write_vtable will return to the original virtual function call site.

F. Fail-Safe Check

As discussed earlier, in case that some modules are not hardened by VTint, the instrumented security check may fail and cause false positives, because there is no ID instrumented before the *vtables* in unhardened modules.

To handle such cases, we deploy a fail-safe solution to provide better compatibility. This fail-safe solution will only enforce the *vtable* to be read-only, but not enforce the existence of the ID if the current module is not hardened. If VTint is not deployed on all modules, the error handler for ID mismatching (i.e., ERROR1 in Figure 7) will first check whether current module is hardened by VTint. If yes, it blocks the control flow, and reports an attack. Otherwise, it will not block the control flow, but just return back to the original code to further check whether the *vtable* is read-only. In this way, VTint can still defeat all *vtable corruption* and *vtable injection* attacks, and also provides better compatibility.

To tell whether a module is hardened by VTint, we will traverse the Process Environment Block (PEB) of the current executable at runtime, to resolve the section information of the current module. If a read-only section named vtsec is found, and the first 4-bytes of this section is the matching ID we used for *vtables*, we can conclude that it is hardened by VTint.

VI. Evaluation

We have implemented a prototype of VTint for x86 PE executables on the Windows platform. In this C++ implementation prototype, our custom disassembler BitCover uses an open source disassembler library Udis86 [47] to decode x86 instructions. In addition to the 8K LOC of Udis86, BitCover takes about 5k LOC, while VRewriter takes another 3k LOC and PEParser takes 2k LOC.

We have tested VTint with the SPEC2000 [16] and SPEC2006 [17] benchmark binaries, and some real world browsers including Firefox, Chrome and Internet Explorer 6 and 8 (denoted as IE6 and IE8),[1] to evaluate our defense solution's performance and protection effectiveness.

A. Performance of Static Analysis

In this section, we first describe our experiment setup, and then show the analysis time of VTint, and the file size overhead brought by VTint.

1) Experiment Setup: The SPEC2000 and SPEC2006 benchmarks are composed of applications written in C/C++ and Fortran. To harden these benchmarks with VTint, we first compile them to PE binaries supporting relocation table. For those applications written in C/C++, we compile them using the Microsoft Visual Studio 2010 compiler. For those written in Fortran, we use the Intel Fortran Compiler to compile them. For each application, all modules are statically linked together. This experiment is performed on a Windows 7 32bit system, with an Intel Core2 Duo CPU at 3.00GHz.

We use VTint to automatically disassemble and rewrite all these benchmark applications' binaries. The functions identified by VTint are compared with the symbol information from the source code. The result shows that there is no false positives or false negatives when parsing these binaries. We then replace the original SPEC binaries with the hardened binaries generated by VTint, and then run the performance test by using the original SPEC harness scripts. The performance test scripts also check the behavior of the benchmark applications by matching the outputs of applications with expected outputs. Results show that the behavior of hardened applications is same as the original ones', indicating that the disassembling and rewriting of VTint is correct and does not break target applications.

For real world browsers including Firefox and Chrome, we harden each executable module (i.e., *.dll and *.exe) of the browser using VTint. We then replace the original modules with the hardened ones to test the performance and protection effectiveness. Results show that VTint is able to disassemble

[1] Newer browsers often require newer OS, causing them hard to be replaced by VTint due to the integrity protection provided by the OS. In addition, there are fewer public exploits available for defense evaluation. Hereby, we chose these two old version browsers for testing.

TABLE III: Analysis results of Firefox modules, including the analysis time each module takes, the file size information of the original binaries and the hardened binaries, and the statistics of the VTable related information.

App	analysis time (sec)	file size (KB)			VTable info		
		orig	new	size overhead	#inst	#vtables	#vcalls
crashreporter.exe	1.8	116	117	0.52%	18,461	3	15
updater.exe	3.7	271	276	1.77%	112,693	9	17
webapprt-stub.exe	1.6	96	97	0.61%	38,589	2	17
D3DCompiler_43.dll	74.3	2,106	2,202	4.53%	2,135,041	48	1338
d3dx9_43.dll	36.9	1,998	2,184	9.33%	627,400	124	4152
gkmedias.dll	84.9	4,221	4,493	6.45%	2,130,418	483	5542
libEGL.dll	0.99	59	64	7.99%	17,772	3	156
libGLESv2.dll	23.7	473	519	9.91%	913,890	87	983
mozjs.dll	123.6	2,397	2,444	1.95%	4,553,743	35	174
msvcp100.dll	5.0	421	450	6.79%	78,586	116	438
msvcr100.dll	13.2	770	778	0.92%	291,484	91	270
xul.dll	328.9	15,112	17,768	17.57%	5,801,649	6548	54743

TABLE IV: Analysis results of Chrome modules, including the analysis time each module takes, the file size information of the original binaries and the hardened binaries, and the statistics of the VTable related information.

App	analysis time (sec)	file size (KB)			VTable info		
		orig	new	size overhead	#inst	#vtable	#vcalls
delegate_execute.exe	21.9	2,105	2,162	2.68%	629,884	428	1,628
chrome_elf.dll	1.9	132	131	-1.3 %	49,138	11	34
d3dcompiler_43.dll	59.4	2,106	2,199	4.42%	2,135,041	48	1,338
d3dcompiler_46.dll	106.8	3,231	3,489	7.99%	3,979,873	622	5,017
libegl.dll	1.8	174	174	-0.41%	47,588	9	64
libglesv2.dll	31	1,097	1,157	5.43%	1,303,450	165	1,885
libpeerconnection.dll	23.6	2,461	2,668	8.37%	658,700	437	6,416
metro_driver.dll	6	504	534	5.88%	181,753	177	920
pdf.dll	70.3	8,577	8,868	3.38%	2,145,987	1,048	8,602
ppgooglenaclpluginchrome.dll	3.7	331	334	0.64%	106,962	51	142
widevinecdmadapter.dll	1.5	137	138	0.28%	47,423	29	110
xinput1_3.dll	1.3	81	79	-3.12%	23,645	2	4
chrome.exe	8.1	852	865	1.47%	231,493	82	433

and rewrite real world binaries. For example, the module `xul.dll` in Firefox (version 17.0.1) has a size of more than 15MB, containing about 6 millions instructions and more than 70,000 functions. `VTint` can automatically handle these binaries without causing problems. These experiments are also performed on a Windows 7 32bit system.

We also harden some modules of the IE browser (e.g., the core module `mshtml.dll`). These modules are extracted from a virtual machine running Windows XP SP3. After hardening, they are copied back to replace the original ones.[2]

2) Analysis Time: The disassembling and rewriting process of `VTint` is quite fast. In general, it only takes several seconds to harden a single module.

As shown in Table II, for most SPEC applications that have virtual functions, the analysis time is less than 10 seconds. The application `483.xalancbmk` from the SPEC2006int subset takes the longest time, about 212 seconds. Actually it is also the largest one among these applications, with 267K lines of source code and about 900K machine code instructions.

Table III and Table IV show the analysis time of `VTint` when handling the executable modules of Firefox and Chrome respectively. Most of these modules take less than one minute. The module `xul.dll` in Firefox takes the longest time (about 5 minutes), due to its size. This module is more than 15MB and has about 6 million instructions.

We also tested the IE 6 browser, the results are similar and we omit the detailed data. For example, the core module `mshtml.dll` is about 4MB and has about 1 million instructions. It only takes 51 seconds for `VTint` to disassemble and rewrite this module.

3) File Size Overhead: `VTint` needs to copy the original *vtables* to a new data section, instrument IDs in this section, and also instruments virtual function dispatches with security checks to validate the target *vtable*'s integrity. In general, `VTint` will allocate new data and code sections to put the new *vtables* and the security check instructions in.

Moreover, entries in the *vtables* are all function pointers, and need to be relocated when the binary module is loaded into memory. As a result, we need to update the executable binary's relocation table, to record all entries in the new *vtables*. In addition, the security checks may introduce new relocation entries too. These entries also need to be recorded in the relocation table. Therefore, `VTint` also creates a new relocation table. All these newly generated relocation table and sections contribute to a file size increment to target binaries.

As shown in Table II, Table III and Table IV, the overall file size overhead brought by `VTint` is about 5%. For example, the file size overhead of the SPEC2006 application `483.xalancbmk` is about 6.97%. The file size overhead of

[2]These modules are in the system directory, protected by the operating system. And thus, special efforts are needed to replace these modules.

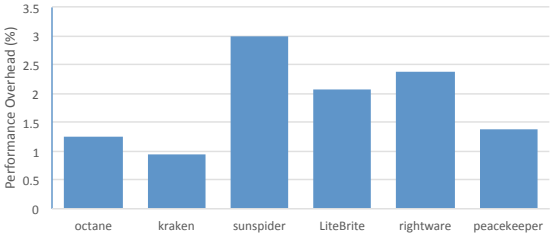


Fig. 8: Performance overhead of `VTint`-hardened Firefox, tested against 6 popular web benchmarks. The geometry mean performance overhead is 1.84%.

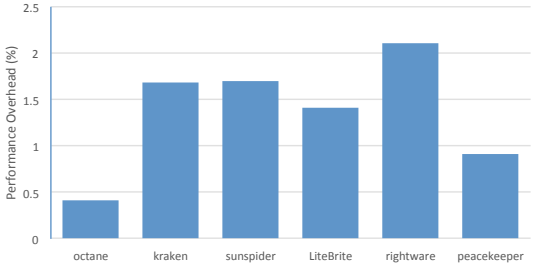


Fig. 9: Performance overhead of `VTint`-hardened Chrome, tested against 6 popular benchmarks. The geometry mean performance overhead is 1.37%.

the `xul.dll` module of Firefox is about 17.57%, largest in all these tested modules. Meanwhile, the file size overhead of the `chrome.exe` is about 1.47%.

It is worth noting that, for some binaries, their hardened version may have a smaller file size. For example, the `libegl.dll` module of Chrome becomes smaller after hardening. This is a normal phenomenon, because some binary files include extra padding bytes at the end of the files. `VTint` will remove these padding bytes, and output a smaller file.

For the IE browser which is not listed here, its file size overhead is also low. For example, the file size overhead of `mshtml.dll` is about 9%.

Furthermore, in the current implementation of `VTint`, the security checks are all instrumented in a new code section because there may not be enough space around the virtual function dispatch instructions. But in fact, compilers usually instrument several padding bytes around functions. With an accurate disassembling, it is possible to identify all these padding bytes in the original code section, and then reuse these bytes to instrument security checks. Our future work will apply this optimization to get a smaller file size overhead.

B. VTable Statistics

`VTint` first needs to identify some key structures in the binary, including *vtables* and virtual function dispatches, before deploying the security policy. In this section, we give the statistics of these structures.

In the Table II, columns in the `VTable info` group show the statistics related to *vtables*. For example, among the 872K instructions of `483.xalancbmk`, there are 4248 virtual function dispatches. These virtual function dispatches may use virtual function pointers coming from 29 *vtables*. There is a special case in this table, the application `473.astar` has two *vtables*, but has no virtual function dispatches. After manual analysis of the source code, we confirm that, there are two classes containing virtual functions. But any objects of these two classes or their sub-classes never invoke these virtual functions. In other words, the *vtable* information of `473.astar` identified by `VTint` is correct.

Table III and Table IV shows the statistics of *vtable* related information of the modules from Firefox and Chrome. For example, in the Firefox's module `xul.dll`, there are 6548 *vtables*, and 54743 virtual function dispatches. As there are about 5.8 millions of instructions in total in this module, we find that, in every 100 instructions, there is about one virtual function dispatch. This ratio is quite high, indicating that the attack surface of *vtable hijacking* attacks is large.

C. Runtime Performance

The runtime performance is a key factor affects the adoption of a defense solution. In this section, we test the overhead brought by `VTint` by comparing the performance of original binaries and the hardened binaries of several applications.

1) Runtime Overhead on SPEC *benchmarks:* First, we test the SPEC benchmark applications' performance. Table II shows the performance data of the original binaries and the hardened binaries. The average performance overhead is about 0.37% for these applications. The highest performance overhead is the one of `252.eon`, about 2.96%. Some applications even get faster after hardening, such as the `444.namd` and `453.povray`, maybe due to some experimental errors.

It is worth noting that, because there is no virtual function dispatches in `473.astar` (although it has 2 *vtables*), `VTint` does not instrument any security checks for this application. On the other hand, the performance overhead of `473.astar` is zero, also respects this fact.

2) Runtime Overhead on Real World Browsers: Then we test the performance of two real-world browsers: Firefox and Chrome. We use six popular browser benchmarks to evaluate these two browsers' performance on JavaScript, HTML rendering and HTML5 support, including Google's Octane [14], Mozilla's Kraken [29], Apple's Sunspider [3], Microsoft's LiteBrite [26], RightWare's BrowserMark [41] and FutureMark's PeaceKeeper [12].

As shown in Figure 8 and Figure 9, the overall performance overhead for all browsers is about 1.6%, whereas the average overhead of Chrome and Firefox is about 1.37% and 1.84%.

3) Performance Analysis: The runtime performance overhead brought by `VTint` is lower than most of previous solutions. The SafeDispatch solution [20] introduces a performance overhead of 7% (using method checking) and 30% (using *vtable* checking), when no runtime profile information is provided. The memory allocation solution DieHard [4] also introduces an overhead of 8%.

The major performance overhead of `VTint` comes from the instrumented security checks that test the target *vtable*'s read-only property. As described in Section V-E, the current

TABLE V: Real World Exploit Samples Prevented By VTint. These exploits are collected from public resources, including the famous database exploit-db. IE here stands for Internet Explorer, and FF stands for Firefox.

CVE-ID	App	Vul Type	POC Exploit	Protected
CVE-2010-0249	IE6	use-after-free	*vtable injection* [5]	YES
CVE-2012-1876	IE8	heap overflow	*vtable injection* [38]	YES
CVE-2013-3205	IE8	use-after-free	*vtable injection* [7]	YES
CVE-2011-0065	FF3	use-after-free	*vtable injection* [40]	YES
CVE-2012-0469	FF6	use-after-free	*vtable injection* [15]	YES
CVE-2013-0753	FF17	use-after-free	*vtable injection* [23]	YES

solution used by VTint is based on Windows' Structured Exception Handling (SEH) mechanism. When testing legitimate read-only *vtables*, the security checks will trigger memory write violation exceptions. The registered SEH handler will then catch these exceptions, and finally redirect the control flow to the application's original code. When testing illegal *vtables* that are writable, the security checks will not trigger exceptions, but flow to a predefined error handler.

This testing works fine, but its performance overhead is still high. Most (even all) target *vtables* at runtime are legitimate and thus are read-only, and cause many memory write exceptions, leading to a higher-than-desired performance overhead. As far as we know, there is no other alternate solution that is more efficient to check whether a memory is writable. We hope the hardware or the operating system can provide supports to do a quick memory property check in the future, similar to the hardware support for DEP. In that case, the performance overhead of VTint will become much lower.

D. Protection against Real World Exploits

To evaluate the effectiveness of VTint, we choose 6 publicly available *vtable hijacking* exploits from the Internet, including security research blogs such as Vupen's blog, the penetration testing framework Metasploit [24] and the exploit database exploit-db [35], as shown in Table V. These exploits all target real-world browser applications, e.g., Internet Explorer and FireFox, by exploiting vulnerabilities such as use-after-free and heap overflow. They all inject fake *vtables* into the memory and finally launch the *vtable injection* attacks. For the Chrome browser, there are very few public available exploits. So, it is not tested here.

Table V also lists other detailed information of the 6 exploits we used. For each exploit, we list the vulnerabilities' CVE-ID, application version, and type of the vulnerabilities and the reference of a detailed POC description.

We carry out these experiments in a virtual machine running Windows XP SP3. Core modules of target applications are extracted from the virtual machine first. Then VTint disassembles and rewrites these modules in our analysis platform. Finally, these hardened modules are copied back to the virtual machine to replace the original ones. After deploying the hardened modules, we drive target applications to access the malicious URLs that contain these exploits. Results show that, the security checks instrumented by VTint will block all these exploits. In other words, VTint is able to protect real world browsers from *vtable hijacking* attacks.

VII. Security Analysis and Discussion

By checking the memory property of *vtables*, VTint can defeat all *vtable corruption* and *vtable injection* attacks. The only way to bypass VTint is to reuse fake *vtables* in the read-only memory, like using code reuse attacks (e.g., return-to-libc or ROP attacks) to bypass DEP.

In addition, VTint will instrument a special ID for *vtables*, and match this ID before virtual function dispatches. This solution is resilient to information leakage attacks. As a result, attackers can only reuse existing *vtables* in the read-only memory to bypass VTint. As there are not many *vtables* in an application, the attack surface is small. It is also hard to find an existing *vtable* that is useful to launch further attacks.

Further, a fine-grained *vtable* checking policy is able to defeat this type of *vtable reuse* attack. A fine-grained *vtable* checking policy, however, requires the type information and class hierarchy information of target applications, which is hard to retrieve using binary analysis. Our future work will focus on identifying this information from binaries, and extend VTint to defeat all *vtable hijacking* attacks.

Currently, to deploy VTint in practice, we can combine it with other lightweight solution to provide better security. For example, the solution VTGuard is a perfect choice. VTGuard also instruments IDs for each *vtable*. It can defeat *vtable reuse* attacks, but is vulnerable to *vtable corruption* and *vtable injection* attacks. It also has a negligible performance overhead (less than 0.5%).

Combining with the VTGuard solution, VTint can defeat all *vtable hijacking* attacks. The overall performance overhead is still acceptable (it should be less than 2% together). It is worth noting that, after combining, the VTGuard solution is also resilient to information leakage attacks, because the target *vtables* must be read-only and cannot be forged or corrupted by attackers even if VTGuard's IDs have been leaked. Moreover, currently VTint implements a similar ID checking to defeat some *vtable reuse* attacks. When combining with VTGuard, this ID checking can be merged with VTGuard's. As a result, the overall performance overhead can be further reduced.

VTint is a pure binary instrumentation solution. It identifies *vtable*-related components (e.g., virtual function tables and virtual function call sites) from the binary, by leveraging certain patterns and calling conventions implemented by major compilers. VTint also cannot handle binaries that are obfuscated. For binaries which are obfuscated or do not follow these patterns, it is still an open challenge to enable the needed binary analysis.

VIII. Related Work

In this section, we discuss some representative works on the defense of *vtable hijacking* attacks and other control flow hijacking attacks, as well as the binary analysis work.

Memory Safety. The memory safety policy ensures that no out-of-bounds or dangling pointers can be exploited for unauthorized read or write of memory. As a result, attackers cannot tamper target applications' program states, or launch control-flow hijacking attacks, including the *vtable hijacking* attacks. There are many memory safety solutions proposed

by researchers, e.g., [10, 36, 44, 51]. These solutions usually instrument pointers with extra metadata when they are created, track the metadata during the program execution, and check the metadata when these pointers are used to access memory.

The most representative solutions are the *spatial* memory safety solution SoftBound [31] and the *temporal* memory safety solution CETS [32]. These solutions, however, all introduce a high performance overhead, prohibiting their adoptions. For example, the combination of SoftBound and CETS will enforce a *complete* memory safety at the cost of 2x or more performance overhead.

Code Pointer Integrity (CPI) [21] proposed a lightweight memory safety solution, protecting only sensitive pointers including code pointers. The performance overhead of CPI is about 8.4%. It also needs the source code to enable the code pointer analysis and instrumentation.

Control-Flow Integrity. The CFI solution provides a strong guarantee that all control flow transfers must comply with programmers' intentions, i.e., they must respect the program's compile-time Control-Flow Graph (CFG). It can stop many different types of control flow hijacking attacks including ROP [45], return-to-libc [46] and *vtable hijacking* attacks. The original CFI solution was proposed in 2005 [1]. It is not adopted by the industry, however, because it requires source code of target applications and introduces a high overhead.

Recently proposed coarse-grained CFI solutions [52, 53] deploy CFI directly on binary executables. As there is no type information in binaries, only a coarse-grained CFI policy is thus enforced. As a result, attackers can bypass the protection in some cases [13].

Some other CFI solutions [33, 50] enforce a fine-grained CFI policy on target applications. These solutions depend on the information collected by compilers at compile-time or by virtual machines. The dependency on source code and the high performance overhead, however, also restrict their adoption.

VTable Hijacking Defense. SafeDispatch [20] hardens the program at compile time to defeat *vtable hijacking* attacks. It utilizes the LLVM complier infrastructure to perform a whole program Class Hierarchy Analysis, and compute the valid method set and the valid *vtable* set for each virtual function dispatch, and then validate them at runtime. Researchers from Google also proposed a similar approach [48]. These solutions, however, need source code of all modules and require recompiling target applications to deploy the security policy. In addition, their runtime overhead is still high.

VTGuard [28] is another solution to protect applications from *vtable hijacking* attacks, deployed in the modern Internet Explorer browsers. This solution instruments secret cookies in *vtables* and match them at runtime. It has a negligible performance overhead and can defeat *vtable reuse* attacks. But it also needs target applications' source code. Furthermore, it is vulnerable to information leakage attacks, *vtable corruption* and *vtable injection* attacks.

DieHard [4] proposes a custom memory allocator to provide probabilistic memory safety at runtime. By randomizing and replicating heap objects, it can tolerate various memory errors (e.g., heap-based buffer overflows, use after free) with a high probability. As a result, some of the *vtable hijacking* exploits can be eliminated by this solution. The average overhead of this solution is about 8%.

Binary Analysis. Many studies have been made to analyze binaries. Schwarz et al. [43] discuss the disassembly problem in detail, including two standard algorithms and a new combination. Many other approaches [18, 19, 39] are proposed to identify code and data in binaries. Some systems such as Vulcan [11] provide a general framework for binary rewriting. The work [9] also discusses some efforts to identify virtual function dispatches. It also uses some similar heuristics as `VTint`, but targets different problems. In addition, it is built on top of the commercial disassembler IDA Pro [18], and relying on the LLVM compiler framework to analyze target binaries.

IX. Conclusion

In this paper, we propose a novel approach `VTint` to defeat *vtable hijacking* attacks. It validates the *vtables*' integrity by checking the memory's read-only property and the ID associated with *vtables*. It is able to defeat all *vtable corruption* and *vtable injection* attacks, and most *vtable reuse* attacks. It is also resilient to information leakage attacks.

`VTint` can be applied through binary rewriting on executables generated by modern compilers. It provides a good modularity support and backward compatibility support. Its runtime performance overhead is low (less than 2%). It can defeat real world *vtable hijacking* attacks.

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