A compiler level intermediate representation based binary analysis and rewriting system

Abstract

This paper presents component techniques essential for converting executables to a high-level intermediate representation (IR) of an existing compiler. The compiler IR is then employed for three distinct applications: binary rewriting using the compiler's binary back-end, vulnerability detection using source-level symbolic execution, and source-code recovery using the compiler's C backend. Our techniques enable complex high-level transformations not possible in existing binary systems, address a major challenge of input-derived memory addresses in symbolic execution and are the first to enable recovery of a fully functional source-code.

We present techniques to segment the flat address space in an executable containing undifferentiated blocks of memory. We demonstrate the inadequacy of existing variable identification methods for their promotion to symbols and present our methods for symbol promotion. We also present methods to convert the physically addressed stack in a binary (with a stack pointer) to an abstract stack (without a stack pointer). Our methods do not use symbolic, relocation, or debug information since these are usually absent in deployed binaries.

We have integrated our techniques with a prototype x86 binary framework called SecondWrite that uses LLVM as IR. The robustness of the framework is demonstrated by handling binaries totaling more than a million lines of source-code, produced by two different compilers (gcc and Microsoft compiler), three languages (C, C++, and Fortran), two operating systems (Windows and Linux) and a real world program (Apache).

1. Introduction

In recent years, there has been a tremendous amount of activity in executable-level research targeting varied applications like security vulnerability analysis [10, 33], bug testing [12], and binary optimizations [22, 30]. In spite of a significant overlap in the overall goals of various source-code methods and executables-level techniques, several analyses and sophisticated transformations that are well-understood and implemented in source-level infrastructures have yet to become available in binary frameworks. Many of the executable-level tools suggest new techniques for performing elementary source-level tasks. For example, PLTO [30] proposes a custom alias analysis technique to implement a simple transformation like constant propagation in executables. Similarly, several techniques for detecting security vulnerabilities in source code [7, 36] remain outside the realm of current executable level frameworks.

It is a well known fact that a standalone binary without any metadata is less amenable to analysis than the source code. *However, we believe that one of the prime reason why current binary frameworks resort to devising new techniques is that these frameworks define their own low-level intermediate representations (IR) which are significantly more constrained than an IR used in a source-code framework*. IRs used in existing binary frameworks lack high level features like an abstract stack, symbols and are machine dependent in some cases. This severely limits the application of source-level analyses to binaries and necessitates new research to make them applicable. We convert the binaries to the same high-level IR that compilers use, enabling the application of source-level research to executables.

We have integrated our techniques in a prototype x86 binary framework called SecondWrite that uses LLVM, a widely-used compiler infrastructure, as IR. Our framework has the following main applications:

- **Binary Rewriting** Existing binary backend of the compiler is employed to obtain a new rewritten binary. The presence of a compiler-IR provides various advantages to our framework (i) It enables every complex compiler transformation like automatic parallelization and security enforcements to run on binaries without any binary-specific customization. (ii) Sharing the IR with a mature compiler allows leveraging the full set of compiler passes built up over decades by hundreds of developers.
- **Symbolic Execution** KLEE [8], a source level symbolic execution engine, is employed directly in our framework without any modifications for detecting vulnerabilities in binaries. Various source level symbolic executors [8, 9] reason about symbolic memory[1] using logical solvers. Directly applying such source-level models to binaries results in high overhead; hence, previous symbolic execution engines for binaries [12, 33] make unsound assumptions by concretizing symbolic memory accesses to a fixed location. Our techniques of obtaining a compiler

[1] The symbolic memory access arises whenever the address referenced in a load or store operation is an expression derived from the user input

```
int main(){
    int z;
    z = foo(10,20);
    return z;
}
int foo(int a, int b)
{
    int c,temp3;
    int temp1,temp2;

    c = a*b;
    temp1 = a+b;
    temp2 = a/b;
    if(a>40){
        temp3 = temp1 + temp2;
    }
    else {
        temp3 = temp1 - temp2;
    }
    c = c + temp3;
    return c;
}
```

(a) Original C Code

```
//Global Stack Pointer
int* llvm_ESP;

char *main()
{
    llvm_ESP = llvm_ESP-2; //Local Allocation

    llvm_ESP2[1] = 20; //Outgoing argument
    llvm_ESP2[0] = 10;
    int llvm_tmp_3 = rewritten_foo();
    return llvm_tmp3;
}

int rewritten_foo()
{
    int* llvm_EBP = llvm_ESP;
                //Local Frame Pointer
    llvm_ESP = llvm_ESP-10;
                //Local Allocation

    int tmpIn1 = llvm_EBP[0]; //Incoming Arg
    int tmpIn2; = llvm_EBP[1];

    int  llvm_tmp1 = tmpIn1*tmpIn2;
    llvm_ESP[1] = llvm_tmp1;

    int  llvm_tmp2 = tmpIn1+tmpIn2;
    llvm_ESP[2] = llvm_tmp2;

    int llvm_tmp3 = tmpIn1-tmpIn2;
    llvm_ESP[3] = llvm_tmp3;

    int llvm_tmpIn3 = llvm_EBP[0];
    if (llvm_tmpIn3 > 40){
       int llvm_tmp5 = llvm_ESP[2];
       int llvm_tmp6 = llvm_ESP[3];
       llvm_ESP[5] = llvm_tmp5 + llvm_tmp6;
    }
    else {
      int llvm_tmp7 = llvm_ESP[2];
      int llvm_tmp8 = llvm_ESP[3];
      llvm_ESP[5] = llvm_tmp7 - llvm_tmp8;
   }
   int llvm_tmp9 = llvm_ESP[1];
   int llvm_tmp10 = llvm_ESP[5];
   int llvm_tmp11 = llvm_tmp9 +
               llvm_tmp10;
  return llvm_tmp11;
}
```

(b) Recovered C Code with physical stack

```
char *main()
{
   int llvm_ESP2[10];

   llvm_ESP2[1] = 20;
   llvm_ESP2[2] = 10;
   int llvm_tmp1 = llvm_ESP2[1];
   int llvm_tmp2 = llvm_ESP2[2];
   int llvm_tmp_3 =
         rewritten_foo(llvm_tmp2,
                      llvm_tmp1);
   return llvm_tmp3;
}

int rewritten_foo( int llvmArg1,
                   int llvm_Arg2)
{
    int llvm_ESP1[10];

    int llvm_tmp1 = llvm_Arg1*llvm_Arg2;
    llvm_ESP1[1] = llvm_tmp1;
    int llvm_tmp2 = llvm_Arg1+llvm_Arg2;
    llvm_ESP1[2] = llvm_tmp2;
    int llvm_tmp3 = llvm_Arg1 - llvm_Arg2;
    llvm_ESP1[3] = llvm_tmp3;

    if (llvm_Arg1 > 40) {
       int llvm_tmp5 = llvm_ESP1[2];
       int llvm_tmp6 = llvm_ESP1[3];
       llvm_ESP1[5] =
                llvm_tmp5 + llvm_tmp6;
    }
   else {
        int llvm_tmp7 = llvm_ESP1[2];
        int llvm_tmp8 = llvm_ESP1[3];
        llvm_ESP1[5] =
               llvm_tmp7 - llvm_tmp8;
    }

    int llvm_tmp9 = llvm_ESP1[1];
    int llvm_tmp10 = llvm_ESP1[5];
    int llvm_tmp11 = llvm_tmp9 +
                         llvm_tmp10;
    return llvm_tmp11;
}
```

(c) Recovered C Code with abstract stack

```
char *main()
{
   int llvm_tmp3;
   llvm_tmp_3 = rewritten_foo(10,20);
   return llvm_tmp3;
}

int rewritten_foo(int llvm_Arg1,
                  int llvm_Arg2)
{
   int llvm_tmp4;
   int llvm_tmp1 = llvm_Arg1 * llvm_Arg2;
   int llvm_tmp2 = llvm_Arg1 + llvm_Arg2;
   int llvm_tmp3 = llvm_Arg1 - llvm_Arg2;
   if (llvm_Arg1 <= 40){
     llvm_tmp4 =  llvm_tmp2 + llvm_tmp3;
   }
   else {
     llvm_tmp4 = llvm_tmp2 - llvm_tmp3;
   }
   llvm_tmp4 = llvm_tmp1 + llvm_tmp4
   return llvm_tmp4;
}
```

(d) Recovered C Code with abstract stack and symbol promotion

Figure 1. *Source Code Example*

```
foo(int a, int b) {
    int *p, *q;

    p = &a;
    ...
    *q = ...;
    ... = b;
    }
```

Source Code

```
Stack        q: edx             p: esp + 8
allocations  a: esp + 20        b: esp + 24

foo:
   subl $16, %esp          // Allocate 16-byte stack frame
   lea 20(%esp), 8(%esp)   // Put &a(esp+20) into p(esp+8)
   store ..., (%edx)       // Store to MEM[q]
   load 8(%esp),%ecx       // Temp ecx ← p (same as &a)
   load 4(%ecx)            // Load “b” by using the fact that
                              &b = &a + 4 = ecx + 4
```

Pseudo Assembly Code

Figure 2. *An example showing the limitation of existing methods for detecting arguments*

```
main() {
   int A[10], i, x;
   x = read-from-file();
   for (i = 0; i < x; i++) {
      A[i] = 10;
   }
}
```

Source Code

```
main:
         subl     $48, %esp
         %ebx = read_from_file
         mov %ebx, 44(%esp) //Initializing x
         movl     $0, 40(%esp) //Initializing i
         jmp      L2  // jump to condition check
     L3:
         movl     40(%esp), %eax  //load i
         movl     $10, (%esp,%eax,4) //Reference A[i]
         addl     $1, 40(%esp)      //Increment i
     L2:
         cmpl     40(%esp), 44(%esp) //compare x and i
         jl       L3
```

Pseudo Assembly Code

Figure 3. *An example showing that variable identification and symbol promotion are different*

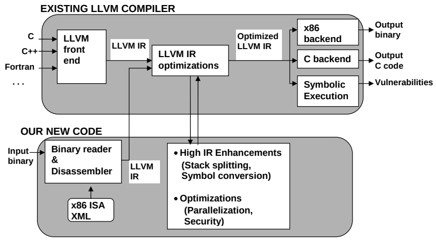


Figure 4. *System Flow*

IR enable simulating a fully symbolic memory using logical solvers for binaries without any excessive cost.

- **Source code recovery** The compiler's C backend is used to convert the IR obtained from a binary to C source-code. Unlike existing tools, which do not ensure the functional correctness of the output code [5, 19], our framework produces a functionally correct source code which can be updated and recompiled by any source code compiler.

  Various organizations like the US Department of Defense(DoD) [2] have critical applications that have been developed for older systems and need to be ported to future versions. In many cases, the application source code is no longer accessible requiring these applications to continue to run on outdated configurations. The ability of our framework to recover a functionally correct source code is highly useful in such scenarios.

It is conventional wisdom that static analysis of executables is a very difficult problem, resulting in a plethora of dynamic binary frameworks. However, a static binary framework based on a compiler IR enables applications not possible in any existing tool and our results establish the feasibility of this approach for most pragmatic scenarios. This work should be seen as what it is: a first successful attempt to statically recover a functional compiler IR from executables, rather than the last word. We do not claim that we have fully solved all the issues; statically handling every program in the world may still be an elusive goal. However, the resulting experience of expanding the static envelope as much as possible is a hugely valuable contribution to the community.

The rest of the paper is organized as follows. Sec 2 highlights our contributions and describes our framework, Sec 3 and Sec 4 present our analyses for obtaining compiler IR. Sec 5 presents the extensions of our analysis for symbolic execution. Sec 6 discusses some practical issues regarding our framework followed by the evaluations in Sec 7.

2. Contributions

We have identified the two tasks below as key for translating binaries to compiler IR. We demonstrate the advantages of these two methods through the source-code recovered from a binary corresponding to the example code in Fig 1(a).

- **Deconstruction of physical stack frames** A source program has an abstract stack representation where the local variables are assumed to be present on the stack but their precise stack layout is not specified. In contrast, a binary program has a fixed (but not explicitly specified) physical stack layout, which is used for allocating local variables as well as for passing the arguments between procedures.

To recreate a compiler IR, the physical stack must be deconstructed to individual abstract frames, one per procedure. Since the relative layout of these frames might change in the rewritten binary, the correct representation requires *all* the arguments (interprocedural accesses through stack pointer) to be recognized and translated to symbols in the IR.

Unfortunately, guaranteeing the static discovery of all the arguments is impossible. Some indirect memory references with run-time-computed addresses might make it impossible for an analysis to statically assign them to a fixed stack location, resulting in undiscovered interprocedural accesses. *Existing frameworks circumvent this problem by preserving the monolithic unmodified stack in the IR, resulting in a low-level IR where no local variables can be added or deleted.*

Some binary analysis tools analyze statically determinable stack accesses to recognize *most* arguments [3], aiding limited code understanding. However, the lack of guaranteed discovery of *all* the arguments renders such best-effort techniques insufficient for obtaining a functional IR. Fig 2 shows an example procedure where the first argument *a* can be recognized statically while the second argument *b* is not statically discoverable. In the assembly code, *&a (esp+20)* is stored to the memory location for *p (esp+8)* (Line 2), which is loaded later to temporary *ecx* (Line 4). The source compiler exploited the layout information (*&a+4=&b*) to load *b* by incrementing *p (&a)* by 4 (Line 5). This is safe since the compiler was able to determine that *p* does not alias *q*. However, the binary framework may not be able to establish this relation, since alias analysis in binaries is less precise. Hence, it has to conservatively assume that the **q* reference (Line 3) could modify *p* which contained the pointer to *a*. Consequently, the source address at Line 5 is no longer known and argument *b* does not get recognized.

Our analysis in Sec 3 defines a source-level stack model and checks if the executable conforms to this model. If the model is verified for a procedure, the analysis discovers the arguments statically when possible, but when not possible, embeds run-time checks in IR to maintain the correctness of interprocedural data-flow. Otherwise, stack abstraction is discontinued only in that procedure.

Fig 1(c) demonstrates the impact of abstract stack on the recovered source code. Fig 1(b) employs a global pointer *llvm_ESP*, corresponding to the physical stack frame in the input binary, for interprocedural communication as well as for representing local allocations in each procedure. However, in Fig 1(c), the stack pointer disappears; instead, local allocations appear as separate local arrays *llvm_ESP1* and *llvm_ESP2* and arguments are passed explicitly.

• **Symbol promotion** Another key challenge we solve is *symbol promotion*, which is the process of safely translating a memory location (or a range of locations) to a symbol in the recovered IR. Existing frameworks do not promote symbols; instead they retain memory locations in their IR [20, 26, 28, 30, 35]. Some post-link time optimizers like Ispike [22] promote memory locations to symbols employing the symbol-table information in the object files. However, deployed binaries do not contain symbol information, rendering such solutions unsuitable for our framework.

At first glance, it seems that the well-known methods for variable identification in binaries, such as IDAPro [21] and Divine [4], can be used for symbol promotion. However, this is not the case. The presence of potentially aliasing memory references is a key hindrance to the legal promotion of these identified variables to symbols.

IDAPro characterizes statically determinable stack offsets in the program as local variables while Divine divides stack memory region into abstract locations by analyzing indirect memory accesses instructions as well.

The example in Fig 3 illustrates the key limitations of both these methods. When the code is compiled, we obtain a stack frame for *main()* of size 48 bytes (10*4 bytes for array *A[]*, and 4×2 = 8 bytes for the scalars *i* and *x*). The accesses to variables *i* and *x* appear as direct memory references (Lines 3,4,6,9) while the array *A* is accessed using an indirect memory reference (Line 7). Both Divine and IDAPro identify memory locations *(esp+44)(x)* and *(esp+40)(i)* as variables based on the direct references. Since the upper bound for the indirect reference to *A[i]* is statically indeterminable, even Divine does not generate any useful information about this access. Hence, it creates three abstract locations - two scalars of 4 bytes each, and a leftover range of 40 bytes.

Despite dividing stack memory region into three abstract locations, none of them can be promoted to symbols. It is impossible to statically prove from the stripped binary that the indirect reference at Line 7 does not alias with references to *i* or *x*. Hence, the promotion of memory locations corresponding to *i* and *x* to symbols would be unsafe since it leads to potentially inconsistent data-flow for underlying memory locations. (Source-level alias analyses often assume that any *A[x]* will access *A[]* within its size. However, such size information is not present in a stripped binary.)

Since identification is inadequate for promotion, we have devised a new algorithm to safely promote a set of memory locations to symbols. It computes a set of non-overlapping promotion lifetimes for each memory location taking into consideration the impact of aliasing memory accesses. Our method is oblivious to the underlying method employed for identifying these locations. The locations can be identified by IDAPro, Divine or through a similar method we use.

Fig 1(d) shows the improvement in source code recovery from symbol promotion. Fig 1(d) demonstrates the replacement of all access to local array *llvm_ESP1* and *llvm_ESP2* in procedures *foo* and *main* respectively by local symbols. As evident, this greatly simplifies the IR and the source code.

Benefits of abstract stack and symbols The presence of abstract stack and symbols has the following advantages:

→ Improved dataflow analysis since standard dataflow analyses only track symbols and not memory locations.

→ Improved readability of the recovered source-code.

→ The ability to employ source-level transformations without any changes. Advanced transformations like compiler-level parallelization [34, 39] add new local variables as barriers and rely on the recognition of induction variables. Several compile-time security mechanisms like StackGuard [14] and ProPolice [16] modify stack layout by placing a *canary* (a memory location) on the stack or by allocating local buffers above other local variables. These methods can be implemented only if the framework supports stack modification and symbol promotion.

→ Efficient reasoning about symbolic memory in case of symbolic execution.

Fig 4 presents an overview of our framework. SecondWrite produces an initial LLVM IR using custom binary reader modules. The initial IR is enhanced using the above mentioned techniques to obtain an enhanced IR, which can be employed for multiple applications described in Sec 1.

3. Deconstruction of physical stack frames

In order to recover a source-level stack representation, we first recognize the local stack frame of a procedure and represent it as a local variable in the IR. As explained in Sec 2, this local variable is coupled with the rest of the stack due to interprocedural accesses. We achieve this decoupling by recognizing interprocedural accesses and replacing them with symbolic accesses to the procedure arguments. Below, both these techniques are presented in detail.

3.1 Representing the local stack frame

We begin by finding an expression for the maximum size of the local stack frame in a procedure X. We analyze all the instructions which can modify the stack pointer, and find the maximum size, P, to which the stack can grow in a single invocation of procedure X among all its control-flow paths. P need not be a compile-time constant; a run-time expression for P suffices when variable-sized stack objects are allowed. An array ORIG_FRAME of size P is then allocated as a local variable at the entry point of procedure X in the IR.

The local variables for the frame pointer and stack pointer are initialized to the beginning of ORIG_FRAME at the entry point of procedure X. Thereafter, all the stack pointer modifications – by constant or non-constant values – are represented as adjustments of these variables. Allocation of a single array representing the original local frame guarantees the correctness of stack arithmetic inside the procedure X.

In some procedures, it might not be possible to obtain a definite expression for the maximum size of the local stack frame. For example, scoped variable-sized local objects in source code might result in a stack allocation with a non-constant amount, whose expression is not available at the beginning of the procedure. Consequently, a single array ORIG_FRAME of a definite size cannot be allocated. Neither can multiple local arrays, one per such stack increment, be allocated since IR optimizations and compiler backend can modify their relative layout thereby invalidating the stack arithmetic. In such procedures, we do not convert the physical stack to an abstract frame. A physical stack frame is maintained in the IR using inline assembly versions of all the stack modification instructions while the remaining instructions are converted to LLVM IR. The runtime checks mechanism presented in the next section is employed to distinguish the local and ancestor accesses.

Persistent stack modification: Returns from a procedure ordinarily restore the value of stack pointer to the value before the call. However, in some cases, the stack pointer might point to a different location after returning from a procedure call. For example, the called procedure can cleanup the arguments passed through the stack. To represent this stack pointer modification, which persists beyond a procedure call, we introduce the following definition:

Balance Number: The balance number for a procedure is defined as the net shift in the stack pointer from before its entry to after its exit. Four different cases can arise:

Case1: *Balance Number* $= 0$

This is the common case; no modification required.

Case2: *Balance Number* < 0

This case arises when a procedure cleans up a portion of the caller stack frame and is represented as an adjustment of the stack pointer by *Balance Number* amount in the caller procedure after the call. The amount need not be a constant.

Case3: *Balance Number* > 0

This case implies that a procedure leaves its local frame on the stack and the corresponding frame outlives the activation of its procedure. Such procedures are represented by considering their allocation as part of the caller procedure allocation. The *Balance Number* amount is added to the size of ORIG_FRAME array in the caller procedure and the stack pointer is adjusted after the call by this amount.

Case4: *Balance Number* Indeterminable

In such a case, we do not convert the physical frame into abstract frame and represent the stack as a default global variable in the IR, as shown in Fig 1(b). This is an extremely rare case and in fact, it did not appear in our experiments.

3.2 Representing procedure arguments

As per the source-level representation, we aim to represent all the stack-based interprocedural communication through explicit argument framework. We discuss why this is not feasible in all the cases and propose our novel methods based on run-time checks to handle such scenarios.

We use Value Set Analysis (VSA) [3] to aid our analysis. VSA determines an over-approximation of the set of memory addresses and integer values that each register and memory location can hold at each program point. Value Set (VS) of the address expression present in a memory access instruction provides a conservative but correct estimate of the possible memory locations accessed by the instruction. VSA accurately captures the stack pointer modifications and the assignments of stack pointer to other registers.

The stack location at the entry point of a procedure is initialized as the base (zero) in VSA and the local frame allocations are taken as negative offsets. Intuitively, memory

```
1.  function foo:
2.          sub 100, esp           // Subtract 100 from sp
3.          call bar               // call bar

4.  function bar:
5.          sub 10, esp            // Subtract 10 from sp
6.          lea 4(esp),edi         //Move address esp+4 to edi
7.          mov 2, ebx             // Move value 2 to ebx
8.          mov 15, ecx            // Move value 15 to ecx
9.          if (esi ≤ 5) jmp B2    //Conditional Branch

10. B1:     mov 4,ebx              // Move value 4 to ebx
11.         mov 16,ecx             //Move value 16 to ecx

 12.B2:     store 10, ebx[edi]     // Store 10 to indirect offset (edi + ebx)
13.         store 10, ecx[esp]     // Store 10 to indirect offset (esp + ecx)
14.         store 10, edx[edi]     // Store 10 to indirect offset (edi + edx)
...
```

Figure 5. *A small psuedo-assembly code. The second operand in the instruction is the destination*

accesses with positive offsets represent the accesses into the parent frame and constitute the arguments to a procedure. A formal argument is defined corresponding to each constant offset into the parent frame and each such access is directly replaced by an access to the formal argument.

However, the above method for recognizing arguments is suitable only if VS of the address expression is a singleton set. If the VS has multiple entries, it is not possible to statically replace it with a single argument. We introduce the following definitions to ease the understanding:

$CURRENT\_BASE$: Stack pointer at the entry point of a procedure.

$addr_m$: The address expression of a memory access instruction *m*

$VS(addr_m)$:Value Set of $addr_m$

(x, y): Lower and upper bounds, respectively, of the possible offsets relative to $CURRENT\_BASE$ in $VS(addr_m)$

$LOCAL\_SIZE$: Size of local frame variable $ORIG\_FRAME$

$SIZE_i$:Size of $ORIG\_FRAME_i$ of the 'ith' ancestor in the call graph, with the caller being represented as the first ancestor. $SIZE_0$ is defined as value 0.

Fig 5 contains an x86 assembly fragment which will be used to illustrate the handling of interprocedural accesses. Fig 6 shows the output IR that results from Fig 5. Three different cases for memory reference categorization of a memory access instruction *m* arise:

Case 1 $(x, y) \subset (-LOCAL\_SIZE, 0)$
This condition implies that the current memory access instruction strictly refers to a local stack location. In Fig 5, Line 12 corresponds to this case. Instruction at Line 6 moves address *(esp+4)* to register *edi*. Since the size of current frame in *bar* ($LOCAL\_SIZE$) is 10 and the local allocations are taken as negative offsets, this translates to VS of *edi* as $\{CURRENT\_BASE$-6$\}$. The VS of *ebx* at Line 12 is {2,4}; therefore the $VS(dest_m)$ is $\{CURRENT\_BASE$ -2 , $CURRENT\_BASE$ - 4$\}$, which translates as a subset of (-$LOCAL\_SIZE$,0). In this case, we replace the indirect access by an access to the local frame as shown Fig 6 (Line 12).

Case 2 $\exists$ N : $(x, y) \subset (\sum_{i \| i \in (0,N)} SIZE_i, \sum_{i \| i \in (0,N+1)} SIZE_i)$
This case implies that the current instruction exclusively accesses the local frame of Nth ancestor. In such cases, we make the local frame variable of the Nth ancestor procedure, $ORIG\_FRAME_N$, an extra incoming argument to the current procedure as well as to all the procedures on the call-graph paths from the ancestor to the current procedure. The indirect stack access is replaced by an explicit argument access.

```
1.  function foo:
2.          ORIG_FRAME_FOO=alloca i32, 100     // Local frame allocation
3.          call    bar(ORIG_FRAME_FOO)        // call bar

4.  function bar(i32* inArg)
5.          ORIG_FRAME_BAR=alloca i32, 10      // Local frame allocation
6.          edi = ORIG_FRAME_BAR+4
7.          ebx = 2                            // Move value 2 to ebx
8.          ecx = 15                           // Move value 15 to ecx
9.          if (esi ≤ 5) jmp B2

10. B1:     ebx = 4                            // Move value 4 to ebx
11.         ecx = 16                           // Move value 16 to ecx

12. B2:     store  10, ebx[edi]                // Store 10 to local frame
13.         store  10, (ecx-SIZE-BAR)[inArg]   // Ancestor Store
14.         if((edx+edi - ORIG_FRAME_BAR) ≤ SIZE-BAR)  //Run Time Check
15                  store 10, edx[edi]                 //Local Store
16.         else
17.                 store 10, (edx+edi - SIZE-BAR)[inArg]//Ancestor Store
...
```

Figure 6. *IR of the psuedo-assembly code. SIZE-BAR is size of ORIG_FRAME_BAR, register names are pure IR symbols*

Line 13 in Fig 5 represents this case. Here, VS of *ecx* is {15,16} which translates to the stack-offset range (5,6) which is subset of (0,$SIZE_1$). Line 13 in Fig 6 shows the adjusted offset into the formal argument *inArg*.

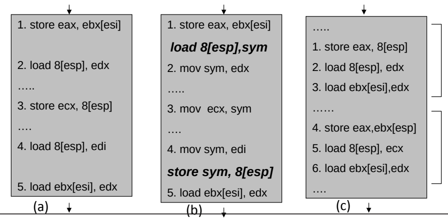
Case 3
$\exists$ N: $\{\ \{(x, y) \cap (\sum_{i \| i \in (0,N)} SIZE_i, \sum_{i \| i \in (0,N+1)} SIZE_i)\ \neq\ \phi\} \wedge \{(x, y) \not\subset (\sum_{i \| i \in (0,N)} SIZE_i, \sum_{i \| i \in (0,N+1)} SIZE_i)\ \}\ \}$
This case arises when VSA cannot bound the memory access exclusively to the local frame of one ancestor or to the local frame of the current procedure. It also includes cases where VS of the target location is *TOP* (*i.e.*, unknown).

We propose a run-time-check-based solution to represent such accesses in the IR. We define all the possible ancestor stack frames in the call graph as arguments to this procedure. Further, at the indirect stack access, a run-time check is inserted in the IR to dynamically translate the access to the local frame or to one of the ancestor stack frames.

Line 14 in Fig 5 represents this case. Suppose edx is data-dependent and hence its VS is TOP. Line 14 in Fig 6 shows the run-time check inserted based on this value. Depending on this check, we either access the local frame (Line 15) or the incoming argument (Line 17).

The cases where $LOCAL\_SIZE$ is not statically known fall naturally under Case 3. The cases where the stack-frame size in the caller at the point of call is smaller than $LOCAL\_SIZE_i$, are handled by considering the actual size at the point of call instead of $LOCAL\_SIZE_i$. We have neglected the size of return buffer in our calculations for the ease of understanding. It is easily considered in our model by adding the return buffer size to each ancestor's local frame size.

In the case of dynamically linked libraries (DLLs), the procedure body is not available; hence the above method for handling the arguments cannot be applied. In order to make



```
1. store eax, ebx[esi]

2. load 8[esp], edx
.....
3. store ecx, 8[esp]
....
4. load 8[esp], edi

5. load ebx[esi], edx
```

```
1. store eax, ebx[esi]
load 8[esp],sym
2. mov sym, edx
.....
3. mov  ecx, sym
....
4. mov sym, edi
store sym, 8[esp]
5. load ebx[esi], edx
```

```
.....
1. store eax, 8[esp]
2. load 8[esp], edx
3. load ebx[esi],edx
......
4. store eax,ebx[esp]
5. load 8[esp], ecx
6. load ebx[esi],edx
....
```

Figure 7. *Symbol promotion. Second operand in the instruction is the destination of the instruction.*

sure that the external procedures access arguments as before, LLVM code generator is minimally modified to allocate the abstract frame, ORIG_FRAME, at the bottom of the stack in each procedure in the rewritten binary. Since external procedures are not aware of the call hierarchy inside a program, their interprocedural references are usually limited to only the parent frame. When the prototypes of these external procedures are available (such as for standard library calls), this stack maintenance restriction is avoided altogether by employing the solution presented for any other procedure.

4. Translating memory locations to symbols

Sec 3 presented methods for deconstructing the physical stack frame into individual abstract frames, one per procedure. Even though this representation allows unrestricted modification of the stack frame, accesses to local variables appear as explicit memory references to locations within this array, which are not amenable to standard data-flow analysis. In this section, we propose our methods for translating these memory operations to symbol operations in the IR.

4.1 Motivation for partitions

As presented in Sec 2, maintaining the data-flow consistency of the underlying memory locations across the whole program is imperative while promoting memory accesses to symbolic accesses. Fig 7(a) shows a small example with three direct accesses to location *(esp+8)* at Lines 2,3,4; the remaining two are unbounded indirect accesses. The simplest method for maintaining the data-flow consistency across the program is to load the data from the memory location into the symbol just after each aliasing definition, store the symbol back to the memory location just before each aliasing use and promote each candidate stack access to a symbolic access, as shown in Fig 7(b). The load inserted just after the aliasing definition is referred to as a *Promoting Load* and store just before the aliasing use is referred to as a *Promoting Store* (shown as bold in Fig 7(b)). Although this method ensures correct data flow propagation, it results in a large number of promoting loads and stores which might overshadow the benefit of symbol promotion.

Fig 7(c) illustrates this unprofitable case. In this example, suppose VS of *ebx* is TOP. Consequently, the instructions at Line 3, 4 and 6 are aliasing indirect accesses to the stack location *(sp+8)*. In order to promote the direct memory accesses at instructions 1, 2 and 5, we need to insert Promoting Stores just before instruction 3 and instruction 6 and a Promoting Load just after instruction 4. Hence, promoting three direct memory operations entails the insertion of three extra memory operations, nullifying the benefit.

Statement s	gen[s]	kill[s]

d:store x,mem[reg]	if([sp+addr]∈VS(mem+reg))
	d
	else { }
		if([sp+addr]∈VS(mem+reg))
		defs(addr) - d
		else { }

d: store y,addr[sp]	d	defs(addr) - d
d: z = load mem[reg]	{}	{}
d: z = load addr[sp]	{}	{}

Memory location loc : $[sp + addr]$
mem: Non-constant access
$addr$: Constant
$defs(addr)$: Set of instructions defining the memory location [sp+addr]
$in[n]$: Set of definitions that reach the begining of node n
$out[n]$: Set of definitions that reach the end of node n
$pred[n]$: Predecessor nodes of node n
$in[n] = \cup_{i|i \in pred[n]}\{(out[p])\}$
$out[n] = gen[n] \cup (in[n] - kill[n])$

Figure 8. *The reaching definition description. Definitions are propagated across the control flow of program*

We propose a novel partition-based symbol promotion algorithm where we divide the program into a set of non-overlapping promotional lifetimes for each memory location. It serves as a fine-grain framework where the symbol promotion decision can be made independently for each lifetime (a partition) instead of the entire program at once. Not doing symbol promotion in a partition does not affect the correctness of the data-flow in the program. The symbol promotion can be selectively performed in only those partitions where it is provably beneficial. Fig 7(c) shows an intuitive division of the current example into two safe partitions.

4.2 Reaching definition framework

We define a new reaching definition analysis on *memory locations* for computing the partitions. This is different from the standard reaching definitions on *symbols* well-known in compiler theory. For each memory location *loc*, this analysis computes the set of instructions defining the memory location *loc* that reach each program point. The set of definitions includes stores to the memory location *loc* using direct addressing mode as well as possibly aliasing stores.

Fig 8 formulates the reaching definition in terms of VS of the memory accesses. These reaching definitions are propagated across the control flow of the program, similar to the standard compiler dataflow propagation, allowing the partitions to be formed across basic blocks. The interprocedural version of VSA implicitly takes into consideration a local pointer passed to a procedure through an argument.

4.3 Symbol promotion algorithm

The candidates for symbol promotion in a procedure P, represented by a set LOC, are computed as follows:

M: Set of memory accesses in P

DM: Statically determinable memory accesses, $\bigcup_{d \in M}\{d| \|VS(addr_d)\| = 1\}$

LOC: Statically determined stack locations in P, $\bigcup_{d \in DM}\{m|m \in VS(d)\}$

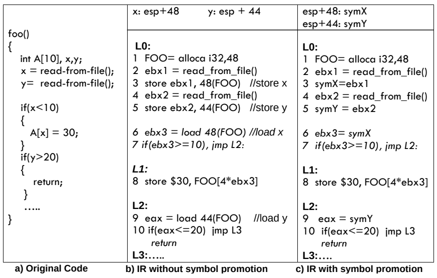
Mathematically, for a stack location *loc*, a single partition constitutes three sets of memory accesses: *DirectAcc*, *BeginSet* and *EndSet*. *DirectAcc* contains statically determinable

Algorithm 1: *Algorithm for computing partitions for a location* loc *in a procedure P*

```
1  L: Set of loads in P; S: Set of stores in P
2  DL:$\bigcup_{l\in L}\{l|\{loc\} = VS(addr_l)\}$ //Direct Loads
3  IL:$\bigcup_{l\in L}\{l|\{loc\} \subset VS(addr_l)\}$ //Indirect Aliasing Loads
4  DS:$\bigcup_{s\in S}\{s|\{loc\} = VS(addr_s)\}$ //Direct Stores
5  IS:$\bigcup_{s\in S}\{s|\{loc\} \subset VS(addr_s)\}$ //Indirect Aliasing Stores
6  Processed: Set of elements processed
7  while DS != $\Phi$||IS != $\Phi$ do
8      define new Partition P, define new list $ActiveList$
9      if DS != $\Phi$ then
10         s = DS.begin; add s to P.DirectAcc
11     else
12         s = IS.begin; add s to P.BeginSet
13     add s to ActiveList
14     while ActiveList.size!=0 do
15         s = $ActiveList.top$; Add s to Processed
16         for $dl \in DL$ do
17             if s $\in$ in[dl] then
18                 add $dl$ to P.DirectAcc
19                 for $s' \in$ in[dl] do
20                     add s' to $ActiveList$ if s' $\notin$ Processed
21                 remove $dl$ from DL
22         if $s \in IS$ then
23             continue /* No need to store symbol back */
24         for $il \in$ {IL,IS} do
25             if s $\in$ in[il] then
26                 add $il$ to P.EndSet
27             for $s' \in$ in[il] do
28                 add s' to $ActiveList$ if s' $\notin$ Processed
29             remove $il$ from $IL$ if $il \in IL$
```

accesses to the location *loc* and constitutes the potential candidates for symbol promotion. *BeginSet* constitutes the indirect stores that may-alias with *loc* and have a control flow path to at least one element of the set DirectAcc. *EndSet* consists of all the aliasing accesses such that there is a control flow path from some element of BeginSet to these accesses. Intuitively, program points just after the elements in BeginSet represent the locations for inserting Promoting Loads. Similarly, program points just before the elements of EndSet are the locations for inserting Promoting Stores.

Algorithm 1 provides a formal description of the method for computing partitions for a memory location *loc*. We begin with an empty partition. We analyze a store instruction, say *ds*. If *ds* is a direct addressing mode instruction then it is added to the DirectAcc set; otherwise it is added to BeginSet (Line 9-12). Load instructions using direct addressing where *ds* is one of the reaching definitions are added to the DirectAcc set of the partition (Line 16-18). The remaining reaching definitions at these load instructions are added to the analysis list (Line 19-20). If *ds* uses a direct addressing mode, indirect load and store instructions with *ds* as one of the reaching definitions are added to the EndSet (Line 24-26). For indirect stores, the symbol need not be stored back to the memory (Line 22-23). As with the direct loads, the rest of the reaching definitions are added to the analysis list (Line 27-29). This analysis is applied repeatedly until the analysis list is empty. At that point, we have one independent partition. We repeatedly obtain new partitions until there are no more direct stores or indirect stores to analyze.



```
foo()
{
   int A[10], x,y;
   x = read-from-file();
   y=  read-from-file();

   if(x<10)
   {
     A[x] = 30;
   }
   if(y>20)
   {
      return;
    }
    .....
}
```

a) Original Code

```
x: esp+48        y: esp + 44

L0:
1  FOO= alloca i32,48
2  ebx1 = read_from_file()
3  store ebx1, 48(FOO)   //store x
4  ebx2 = read_from_file()
5  store ebx2, 44(FOO)   //store y

6  ebx3 = load 48(FOO) //load x
7 if(ebx3>=10), jmp L2:

L1:
8  store $30, FOO[4*ebx3]

L2:
9  eax = load 44(FOO)    //load y
10 if(eax<=20)  jmp L3
     return
L3:.....
```

b) IR without symbol promotion

```
esp+48: symX
esp+44: symY

L0:
1  FOO= alloca i32,48
2  ebx1 = read_from_file()
3  symX=ebx1
4  ebx2 = read_from_file()
5  symY = ebx2

6  ebx3= symX
7  if(ebx3>=10), jmp L2:

L1:
8  store $30, FOO[4*ebx3]

L2:
9   eax = symY
10  if(eax<=20)  jmp L3
      return
L3:....
```

c) IR with symbol promotion

Figure 9. *A small source code example*

```
Constraints:
 A1=write(FOO,48,ebx1)
 A2= write(A1,44,ebx2)
 read(A2,48)<10
 A3 = write(A2,
          4*read(A2,48),30)
Solve:
read(A3,44) <= 20
```

Figure 10. *Constraints for Fig 9(b)*

We implement a simple benefit-cost model to determine whether the symbol promotion should be carried out for a particular partition. In a partition, the size of DirectAcc set is the number of memory accesses replaced by symbol accesses. We define $Freq_i$ as the statically determined execution frequency at program point i. Hence, the benefit of symbol promotion in terms of eliminated memory references:

$$Benefit = \sum_{i|i\in DirectAcc} \{(Freq_i)\} \tag{1}$$

One promoting load/store is needed for each element of BeginSet and Endset, consequently, the cost:

$$Cost = \sum_{i|i\in BeginSet} \{(Freq_i)\} + \sum_{i|i\in EndSet} \{(Freq_i)\} \tag{2}$$

We calculate the net benefit of each partition as *Benefit - Cost*. Symbol promotion is carried out in a partition only if the net benefit is positive.

5. Symbolic Execution

Symbolic execution [9] is a well-known technique for automatically detecting bugs and security vulnerabilities in a program. Among various challenges facing symbolic execution, handling symbolic memory addresses (addresses derived from user-input) is an important one. There are two primary approaches for handling symbolic memory. Previous symbolic executors for executables [12, 33] make simplifying and unsound assumptions by concretizing the symbolic memory reference to a fixed memory location. On the other hand, popular source-level tools like EXE [9] and KLEE [8] employ logical constraint solvers to reason about possible locations referenced by a symbolic memory operation. Even though the expressions involving symbolic memory become more sophisticated, these tools outperform the former approaches in terms of path exploration and bug detection.

The presence of a physical stack and the lack of symbols in an executable pose a difficult challenge in efficiently extending the logical solver based approach for representing symbolic memory in executables. The most straightforward representation of the memory would be a flat byte array. Unfortunately, the constraint solvers employed in existing source-level symbolic execution tools would almost never be able to solve the resulting constraints [8].

The segmented memory representation in our framework, obtained by abstract stack and symbol promotion, drastically improves the efficiency of such constraint solvers by enabling them to only consider the constraints related to the segments referenced by the current memory address expression and ignore the remaining segments.

Fig 9 illustrates this case. Fig 9(a) contains a symbolic memory store to array A. Fig 9(b) and Fig 9(c) show the pseudo IR obtained from an executable corresponding to Fig 9(a), without and with the application of symbol promotion. Fig 10 shows the constraints and query generated at Line 10 while symbolically executing the path L0→L1→L2 in Fig 9(b). Here, *read(A,i)* returns the value at index i in array A and *write(A,j,v)* returns a new array with same value as A at all indices except j, where it has value v.

However, in Fig 9(c), symbol promotion has segmented the array FOO in different segments and references to variables x and y do not refer the segment FOO. Hence, the solver only needs to solve the following simplified query:

$$Solve : symY \leq 20 \tag{3}$$

This example only shows the simplification of constraints with symbol promotion. The presence of an abstract stack also results in a similar simplification of constraints by segmenting the memory space within each procedure.

6. Practical Considerations

Indirect calls and branches :SecondWrite implements various mechanisms, as proposed by Smithson et. al. [32], to address code discovery problems and to handle indirect control transfers. Here, we briefly summarize the mechanism.

A key challenge in binary frameworks is discovering which portions of the code section in an input executable are definitely code. Smithson et. al. [32] proposed *speculative disassembly*, coupled with *binary characterization*, to efficiently address this problem. SecondWrite speculatively disassembles the unknown portions of the code segments as if they are code. However, it also retains the unchanged code segments in the IR to guarantee the correctness of data references in case the disassembled region was actually data.

SecondWrite employs *binary characterization* to limit such unknown portions of code. It leverages the restriction that an indirect control transfer instruction (CTI) requires an absolute address operand, and that these address operands must appear within the code and/or data segments. The code and data segments are scanned for values that lie within the range of code segment. The resulting values are guaranteed to contain, at a minimum, all of the indirect CTI targets.

The indirect CTIs are handled by appropriately translating the original target to the corresponding location in IR through a runtime translator. Each recognized procedure (through speculative disassembly) is initially considered a possible target of the translator, which is pruned further using alias analysis. The arguments for each possible target procedure (Sec 3.2) are unioned to find the set of arguments to be passed to the translator; a stub inside the translator populates the arguments according to the actual target.

Above method is not sufficient for discovering indirect branch targets where addresses are calculated in binary. Hence, various procedure boundary determination techniques, like ending the boundary at beginning of next procedure, are also proposed [32] to limit the possible targets.

Memory Consistency :Our framework mimics the assumptions behind all standard software transformation tools with regards to memory consistency. A majority of compilers (gcc, LLVM, Visual Studio) and popular binary frameworks like PLTO [30], DynamoRIO [6], PIN [26], iSpike [22], Diablo [35] reorder code without taking memory consistency into account. Since synchronization is highly multiprocessor specific, most programmers are expected to write synchronized programs using standard synchronization libraries [27]. The presence of synchronization primitives legalizes the applications of all software optimizations.

Recently, the research community is exploring the possibility of preserving memory consistency in software transformation tools [27]. The key idea is to selectively invalidate the transformations for possibly shared memory locations. Currently, our framework can preserve consistency by declaring all possibly shared memory regions as *volatile* in the IR, we plan to work on more detailed methods in future.

Limitations : Following are the limitations of our current framework, we plan to look at them in future.

- **Self Modifying Code** Like most static binary tools, we do not handle self modifying code. Various tools [37] statically detect the presence of self-modifying code in a program. Such a tool can be integrated in our front-end to warn the user and to discontinue further operation.
- **Volatile Memory** Stripped executables have no information about volatile variables. Existing binary tools [6, 11, 13, 17, 30, 35] ignore their occurrences, instead, we support most common cases of their occurrences. Externally visible volatile variables appear in dynamic symbol table of executables and memory mapped volatile variables can be detected through system calls like *mmap*. Such variables are declared volatile in IR. Volatile variables for exception handling calls (e.g. *setjmp*) are handled by declaring the abstract frame in the corresponding procedure as volatile. Other usages like lock-free variables fall under the umbrella of memory consistency discussed above.
- **Obfuscated Code** We have not tested our techniques against binaries with hand-coded assembly or with obfuscated control flow.

7. Results

Fig 11 lists all the benchmarks which have been successfully evaluated with SecondWrite prototype. It includes the complete SPEC2006 benchmark suite, benchmarks from other suite and a real world program, Apache server. As evident from Fig 11, SecondWrite is able to correctly handle binaries from three different programming languages (C, C++, Fortran), two different compilers (gcc and Visual studio) and two different OS (Linux and Windows), demonstrating the robustness of our techniques. Benchmarks on Linux are compiled with gcc v4.4.1 (O0 (No optimization) and O3 (Full optimization) flags) without any symbolic or debug information. Results are obtained for *ref* datasets on a 2.4GHz 8-core Intel Nehalem machine running Ubuntu. Windows benchmarks are compiled with Microsoft Visual Studio compiler (O0 (No optimization) and O2 (Maximum optimization) flags). *Only the C and C++ programs are included for Windows since VisualStudio does not compile Fortran*. In the case of Apache server, we tested a *http* server with the rewritten binary. Apache server, being an interactive program, is not amenable to performance analysis like other benchmarks; hence we only provide its static results.

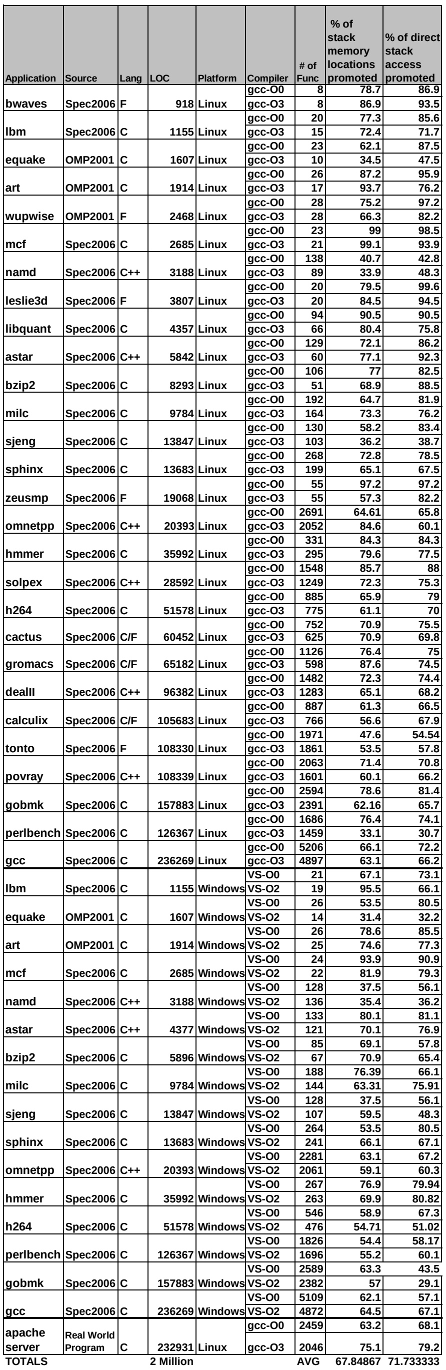
7.1 Static characteristics

Fig 11 displays the static characteristics of symbol promotion techniques - the percentage of stack memory locations promoted to symbols and the percentage of direct stack memory accesses promoted to symbol accesses in the IR and in the recovered source-code. The results depict that on average, 67% of stack locations are promoted to symbols resulting in promotion of 72% of direct stack accesses. For the remaining memory operations, the net benefit for promotion didn't meet the corresponding threshold. Theoretically, our framework can achieve *100% symbol promotion* if the promotion threshold is ignored, but this leads to high overhead in the rewritten binaries due to *Promoting Loads* and *Promoting Stores*. The development of more advanced alias analysis would improve results of our symbol promotion without adversely affecting the performance.

Fig 12 relates the above promoted symbolic references to the original source-level artifacts. We enumerated the symbolic references in the input program using debug information (employed only for counting the references) and compared how many of these symbolic references are restored in the IR and the source code. It shows that our techniques are able to restore 66% of the original symbolic references.

Fig 13 presents an insightful result regarding our partition algorithm (Alg 1). On average, around 76% of the memory locations have one partition, 18% have two to five, and 6% have five or more partitions. This is not unexpected since large procedures are relatively rare.

Fig 14 lists the programs which hit corner cases during the deconstruction of physical stack. The analysis of the original source code revealed that a physical stack frame was required for procedures that call *alloca()*. Runtime checks are inserted in some procedures which accept a variable number of arguments using the *va-arg* mechanism. Most of the procedures using *va-arg* do not require runtime checks. This result establishes our earlier hypothesis that scenarios requiring run-time checks are extremely rare and consequently, have negligible overhead. Nonetheless, not handling these scenarios prohibits obtaining a functional IR and hence, are imperative for any translation system.



Application	Source	Lang	LOC	Platform	Compiler	# of Func	% of stack memory locations promoted	% of direct stack access promoted
bwaves	Spec2006	F	918	Linux	gcc-O0	8	78.7	86.9
					gcc-O3	8	86.9	93.5
lbm	Spec2006	C	1155	Linux	gcc-O0	20	77.3	85.6
					gcc-O3	15	72.4	71.7
equake	OMP2001	C	1607	Linux	gcc-O0	23	62.1	87.5
					gcc-O3	10	34.5	47.5
art	OMP2001	C	1914	Linux	gcc-O0	26	87.2	95.9
					gcc-O3	17	93.7	76.2
wupwise	OMP2001	F	2468	Linux	gcc-O0	28	75.2	97.2
					gcc-O3	28	66.3	82.2
mcf	Spec2006	C	2685	Linux	gcc-O0	23	99	98.5
					gcc-O3	21	99.1	93.9
namd	Spec2006	C++	3188	Linux	gcc-O0	138	40.7	42.8
					gcc-O3	89	33.9	48.3
leslie3d	Spec2006	F	3807	Linux	gcc-O0	20	79.5	99.6
					gcc-O3	20	84.5	94.5
libquant	Spec2006	C	4357	Linux	gcc-O0	94	90.5	90.5
					gcc-O3	66	80.4	75.8
astar	Spec2006	C++	5842	Linux	gcc-O0	129	72.1	86.2
					gcc-O3	60	77.1	92.3
bzip2	Spec2006	C	8293	Linux	gcc-O0	106	77	82.5
					gcc-O3	51	68.9	88.5
milc	Spec2006	C	9784	Linux	gcc-O0	192	64.7	81.9
					gcc-O3	164	73.3	76.2
sjeng	Spec2006	C	13847	Linux	gcc-O0	130	58.2	83.4
					gcc-O3	103	36.2	38.7
sphinx	Spec2006	C	13683	Linux	gcc-O0	268	72.8	78.5
					gcc-O3	199	65.1	67.5
zeusmp	Spec2006	F	19068	Linux	gcc-O0	55	97.2	97.2
					gcc-O3	55	57.3	82.2
omnetpp	Spec2006	C++	20393	Linux	gcc-O0	2691	64.61	65.8
					gcc-O3	2052	84.6	60.1
hmmer	Spec2006	C	35992	Linux	gcc-O0	331	84.3	84.3
					gcc-O3	295	79.6	77.5
solpex	Spec2006	C++	28592	Linux	gcc-O0	1548	85.7	88
					gcc-O3	1249	72.3	75.3
h264	Spec2006	C	51578	Linux	gcc-O0	885	65.9	79
					gcc-O3	775	61.1	70
cactus	Spec2006	C/F	60452	Linux	gcc-O0	752	70.9	75.5
					gcc-O3	625	70.9	69.8
gromacs	Spec2006	C/F	65182	Linux	gcc-O0	1126	76.4	75
					gcc-O3	598	87.6	74.5
dealll	Spec2006	C++	96382	Linux	gcc-O0	1482	72.3	74.4
					gcc-O3	1283	65.1	68.2
calculix	Spec2006	C/F	105683	Linux	gcc-O0	887	61.3	66.5
					gcc-O3	766	56.6	67.9
tonto	Spec2006	F	108330	Linux	gcc-O0	1971	47.6	54.54
					gcc-O3	1861	53.5	57.8
povray	Spec2006	C++	108339	Linux	gcc-O0	2063	71.4	70.8
					gcc-O3	1601	60.1	66.2
gobmk	Spec2006	C	157883	Linux	gcc-O0	2594	78.6	81.4
					gcc-O3	2391	62.16	65.7
perlbench	Spec2006	C	126367	Linux	gcc-O0	1686	76.4	74.1
					gcc-O3	1459	33.1	30.7
gcc	Spec2006	C	236269	Linux	gcc-O0	5206	66.1	72.2
					gcc-O3	4897	63.1	66.2
lbm	Spec2006	C	1155	Windows	VS-O0	21	67.1	73.1
					VS-O2	19	95.5	66.1
equake	OMP2001	C	1607	Windows	VS-O0	26	53.5	80.5
					VS-O2	14	31.4	32.2
art	OMP2001	C	1914	Windows	VS-O0	26	78.6	85.5
					VS-O2	25	74.6	77.3
mcf	Spec2006	C	2685	Windows	VS-O0	24	93.9	90.9
					VS-O2	22	81.9	79.3
namd	Spec2006	C++	3188	Windows	VS-O0	128	37.5	56.1
					VS-O2	136	35.4	36.2
astar	Spec2006	C++	4377	Windows	VS-O0	133	80.1	81.1
					VS-O2	121	70.1	76.9
bzip2	Spec2006	C	5896	Windows	VS-O0	85	69.1	57.8
					VS-O2	67	70.9	65.4
milc	Spec2006	C	9784	Windows	VS-O0	188	76.39	66.1
					VS-O2	144	63.31	75.91
sjeng	Spec2006	C	13847	Windows	VS-O0	128	37.5	56.1
					VS-O2	107	59.5	48.3
sphinx	Spec2006	C	13683	Windows	VS-O0	264	53.5	80.5
					VS-O2	241	66.1	67.1
omnetpp	Spec2006	C++	20393	Windows	VS-O0	2281	63.1	67.2
					VS-O2	2061	59.1	60.3
hmmer	Spec2006	C	35992	Windows	VS-O0	267	76.9	79.94
					VS-O2	263	69.9	80.82
h264	Spec2006	C	51578	Windows	VS-O0	546	58.9	67.3
					VS-O2	476	54.71	51.02
perlbench	Spec2006	C	126367	Windows	VS-O0	1826	54.4	58.17
					VS-O2	1696	55.2	60.1
gobmk	Spec2006	C	157883	Windows	VS-O0	2589	63.3	43.5
					VS-O2	2382	57	29.1
gcc	Spec2006	C	236269	Windows	VS-O0	5109	62.1	57.1
					VS-O2	4872	64.5	67.1
apache server	Real World Program	C	232931	Linux	gcc-O0	2459	63.2	68.1
					gcc-O3	2046	75.1	79.2
TOTALS			2 Million			AVG	67.84867	71.733333

Figure 11. *Benchmarks Table*

7.2 Un-optimized input binaries

Fig 15 shows the normalized run-time of each rewritten binary compared to an input binary produced using gcc with no optimization (-O0 flag). Fig 17 shows the corresponding run-time for binaries produced using Visual Studio compiler with no optimization (-O0 flag). We obtain an average improvement of 40% in execution time for binaries produced

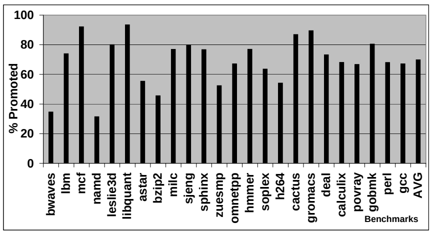


Figure 12. *Percentage of original symbolic accesses recovered in IR*

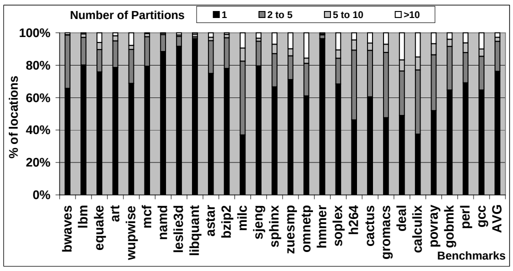


Figure 13. *Partition algorithm visualization*

Program	Version	Physical stack	Run Time Checks
gcc	gcc-O0,VS-O0	117	0
gcc	gcc-O3, VS-Ox	117	10
tonto	gcc-O0, gccO3	20	0

Figure 14. *Programs demonstrating corner cases of our analysis*

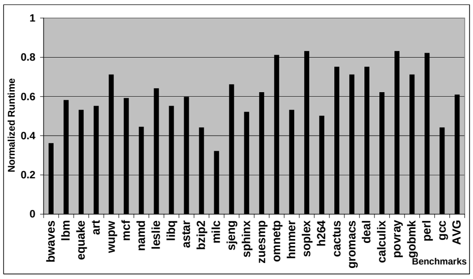


Figure 15. *Normalized runtime of rewritten binary as compared to unoptimized gcc input binary (=1.0)*

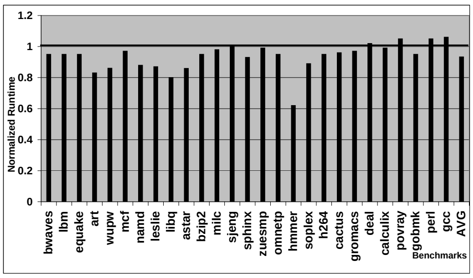


Figure 16. *Normalized runtime of rewritten binary as compared to optimized gcc input binary (=1.0)*

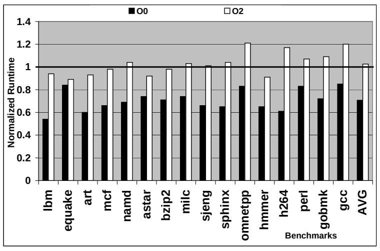


Figure 17. *Normalized runtime of rewritten binary as compared to its corresponding input version (=1.0) compiled by visual studio*

by gcc and 30% for binaries produced by Visual Studio, with an improvement of over 65% in some cases (*bwaves*). In fact, our tool brings down the normalized runtime of unoptimized input binaries from 2.2 to close to the runtime (1.25) of gcc-optimized binaries. (Graph not shown due to lack of space)

7.3 Optimized input binaries

Fig 16 shows the normalized execution time of each rewritten binary compared to an input binary produced using gcc with the highest-available level of optimization (-O3 flag). In this case, we obtain an average improvement of 6.5% in execution time. It is interesting that we were able to obtain this improvement over already optimized binaries without any custom optimization of our own. One of our rewritten binaries (hmmer) had a 38% speedup vs the input binary. Although GCC -O3 is known to produce good code, it missed the creation of few predicated instructions whereas LLVM did this optimization, explaining the speedup. Fig 17 shows the corresponding run-time for binaries produced using Visual Studio compiler with full optimization flag (-O2). As evident, our framework was able to retain the performance of these binaries, with a small overhead of 2.7% on average.

7.4 Impact of symbol promotion

Next, we substantiate the impact of symbol promotion on the run-time of rewritten binaries. Fig 18 and Fig 19 show the normalized improvement in execution time obtained by applying only LLVM optimizations and by applying our symbol promotion techniques. It shows that symbol promotion is responsible for improving the average performance of rewritten binary from 30% to 40% in the case of unoptimized binaries (produced by gcc) and from 1% to 6.5% in the case of optimized binaries (produced by gcc). Since our cost metric is based on static profiling, we observed a small slowdown with symbol promotion in *bzip2 O3*.

It is important to note that these results only measure the impact of symbol promotion. The impact of our method to convert physical frames to abstract frames is not measured above. However, we can infer that number since without obtaining abstract frames, none of the existing LLVM passes would run at all, leading to zero run-time improvement.

7.5 Symbolic Execution

KLEE is efficiently designed to obtain a high code coverage on source programs. We run KLEE in our framework on a set of 50 coreutils binaries and compare the resulting code coverage with that achieved by KLEE on the corresponding source code in Fig 20. Our framework achieves a code coverage of 73% on average compared to 76% obtained by KLEE on source programs.

KLEE has been shown to detect various bugs in a particular version of coreutils (6.10). Fig 21 lists the test cases generated through our framework and those reported in the original KLEE paper[2]. Note that the original KLEE paper detected these bugs from source code, but our framework detected the same bugs from binaries. Both these results demonstrate the unique ability of our framework to efficiently employ source level research directly for executables.

Recall from Sec 5, symbol promotion enables our framework to efficiently reason about symbolic memory accesses. We run symbolic execution on IR produced from binaries for 30 minutes and compare the result in two scenarios: with and without symbol promotion. Fig 22 shows that the presence

[2] The original KLEE paper shows 10 bugs in coreutils but latest version of KLEE only detects five of these bugs

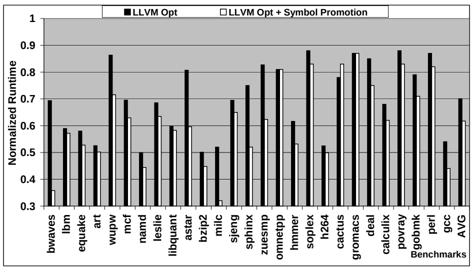


Figure 18. *Impact of symbol promotion on runtime of rewritten binary v/s unoptimized input binary (=1.0)*

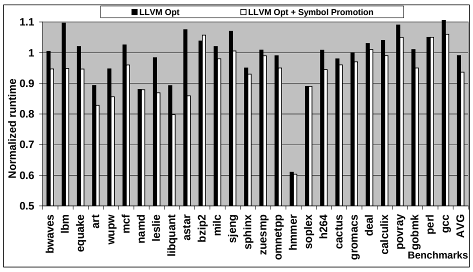


Figure 19. *Impact of symbol promotion on runtime of rewritten binary v/s optimized input binary (=1.0)*

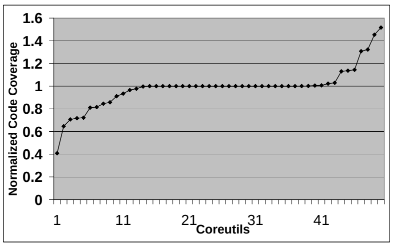


Figure 20. *Normalized code coverage in coreutils binaries as compared to source*

mkdir -Z a b mkfifo -Z a b mknod -Z a b p seq -f %0 1 paste -d\\ abcdefghijklmn	mkdir -Z @@ - mkfifo -Z @@ - mkdir -Z @@ - p@ seq -f %1 1 paste -d\\ @@@@@@@@ @@@@@@

Figure 21. *Testcases for crashes (a)Source code [8] (b) Binaries*

Binary	No Promotion	With Promotion
htget	980	4671
cut	1301	5103
split	1623	4104

Figure 22. *Improvement in constraints processing with symbol promotion*

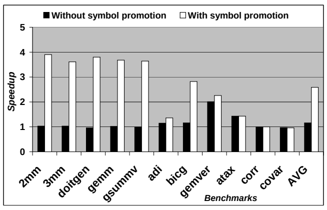


Figure 23. *Automatic parallelization*

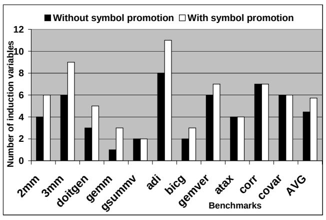


Figure 24. *Number of induction variables recognized*

of symbol promotion greatly improves the number of constraints processed by the solver, resulting in a much higher code coverage in the same amount of time. Program *htget* has been shown to have a bug [10] and we were able to detect that bug within 5 minutes in the presence of symbols as opposed to 25 minutes without symbol promotion.

Further, the presence of a rewriting path in our framework enables us to remedy the above detected bugs in binaries. We analyzed the dump for one of the coreutil binary (*mkdir*), fixed the corresponding behavior in IR and obtained a rewritten bug-free binary.

7.6 Automatic Parallelization

Kotha et al [24] presented a method for automatic parallelization for binaries. Here, we substantiate the impact of symbol promotion on their methods for a subset of *PolyBench* and *Stream* suite. Fig 23 shows that symbol promotion increases the speedup by 2.25x for 4 threads.

The underlying reason for the speedup is that symbol promotion enables discovery of more induction variables. Many induction variables for outer loops are often present on the stack instead of registers. Consequently, such induction variables are not detected by compiler methods, resulting in parallelization of inner loops, which have high synchronization overhead. As shown in Fig 24, symbol promotion enables the discovery of more induction variables, enabling the parallelization of more beneficial outer loops.

8. Related Work

Binary rewriting: Binary rewriting research is being carried out in two directions: static and dynamic. None of the previous dynamic rewriters, PIN [26], DynInst [20], FX!32 [11], DynamoRIO [6], Valgrind [31] and others, employ a compiler IR. Chipounov [15] presented a method for dynamically translating x86 to LLVM using QEMU, reporting a huge slowdown of 35x over the baseline QEMU x86-to-x86 translator. Unlike our approach, it converts blocks of code to IR on the fly which limits the application of LLVM analyses to only one block at a time. Further, they do not provide any methods for stack deconstruction or symbol promotion.

Existing static binary rewriters related to our approach include Etch [28], ATOM [17], PLTO [30], FDPR [18], Diablo [35] and UQBT [13]. All these rewriters define their own low-level custom IR as opposed to using a compiler IR. These IR are devoid of features like abstract frames, symbols and maintain memory as a black-box; the limitations of which have already been discussed in Sec 1. PLTO implements stack analysis to determine the use-kill depths of each function [30]. However this information is used only for low-level custom optimizations like load/store forwarding rather than obtaining a high-level IR. UQBT [13] employs function prototypes in its IR, but relies on user to provide this information, instead of determining it automatically from a binary like we do. This severely limits its applicability since only the developers have access to that information.

Virtual machines [1] implement stack-walking techniques to determine the calling context by simply iterating over the list of frame pointers maintained as metadata in the dynamic framework; making it orthogonal to our run-time checks mechanism which statically inserts checks in the IR.

Binary Analysis: King et al. [23] provide a comprehensive survey of several binary analysis tools. The analysis related to our methods are presented by [3, 4, 38]. Balakrishnan et al [3, 4] present Value Set Analysis for analyzing memory accesses and extracting high level information like variables and their types. As presented in Sec 2, analyzing variables does not guarantee promotion to symbols in IR. Zhang et al [38] present techniques for recovering parameters and return values from executables but they do not consider the scenarios where the information cannot be derived.

Shastry et al [29] present methods for register promotion in source programs. Their method relies on memory locations being represented in SSA form with all the aliases exposed, which is a definite non-starter for binaries. Jianjun et al [25] promote stack variables to registers in a dynamic rewriter, relying on hardware mechanism for memory disambiguation. In contrast, we provide theoretical formulations for symbol promotion without using any hardware support.

Symbolic execution : KLEE [8], EXE [9] are example of source-level tools and cannot be applied directly to executables. Previous binary-only symbolic execution tools like BitBlaze [33], S2E [12] do not represent symbolic memory. MAYHEM [10] proposes a new index-based memory model for simulating symbolic memory, while our techniques enable application of existing solver based models to binaries.

9. Conclusions

We present a binary framework which decompiles the input binary into LLVM IR, allowing the application of source-level complex transformations and advanced symbolic execution strategies on executables and enabling functional source-code recovery. In future, we plan to extend our framework for various platform-specific optimizations.

References

[1] B. Alpern and et. al. The jalapeno virtual machine. *IBM Systems Journal*, 39(1):211 –238, 2000.

[2] Announcement for Binary Executable Transforms. http://www07.grants.gov/.

[3] G. Balakrishnan and T. Reps. Analyzing memory accesses in x86 executables. In *IN CC*, pages 5–23. Springer-Verlag, 2004.

[4] G. Balakrishnan and T. Reps. Divine: discovering variables in executables. In *Proceedings of the 8th international conference on Verification, model checking, and abstract interpretation*, VMCAI'07, pages 1–28, Berlin, Heidelberg, 2007.

[5] Boomerang Decompiler. http://boomerang.sourceforge.net/.

[6] D. Bruening. *Efficient, Transparent, and Comprehensive Runtime Code Manipulation*. PhD thesis, MIT, 2004.

[7] W. R. Bush and et. al. A static analyzer for finding dynamic programming errors. *Softw. Pract. Exper.*, 30(7):775–802, June 2000.

[8] C. Cadar, D. Dunbar, and D. Engler. Klee: unassisted and automatic generation of high-coverage tests for complex systems programs. In *Proceedings of the 8th USENIX conference on Operating systems design and implementation*, OSDI'08, pages 209–224, Berkeley, CA, USA, 2008. USENIX Association.

[9] C. Cadar and et. al. Exe: automatically generating inputs of death. In *Proceedings of the 13th ACM conference on Computer and communications security*, CCS '06, pages 322–335, NY, USA, 2006.

[10] S. K. Cha, T. Avgerinos, A. Rebert, and D. Brumley. Unleashing mayhem on binary code. In *IEEE Symposium on Security and Privacy*, pages 380–394, 2012.

[11] A. Chernoff and et. al. Fx!32 - a profile-directed binary translator. *IEEE Micro*, 18:56–64, 1998.

[12] V. Chipounov, V. Kuznetsov, and G. Candea. S2e: a platform for in-vivo multi-path analysis of software systems. *SIGPLAN Not.*, 46(3):265–278, Mar. 2011.

[13] C. Cifuentes and M. V. Emmerick. Uqbt: Adaptable binary translation at low cost. *IEEE Computer*, 33(3):60–66, March 2000.

[14] C. Cowan and et al. Stackguard: automatic adaptive detection and prevention of buffer-overflow attacks. In *Proceedings of the 7th conference on USENIX Security Symposium*, pages 63–78, 1998.

[15] Dynamically Translating x86 to LLVM using QEMU. http://infoscience.epfl.ch/record/149975.

[16] H. ETO and K. Yoda. Propolice: Improved stack-smashing attack detection. *IPSJ SIGNotes Computer Security 14 (Oct 26)*, 2001.

[17] A. Eustace and A. Srivastava. Atom: a flexible interface for building high performance program analysis tools. In *TCON'95: Proceedings of the USENIX 1995 Technical Conference*, pages 25–25, Berkeley, CA, USA, 1995.

[18] G. Haber and et al. Optimization opportunities created by global data reordering. In *Proceedings of the International symposium on Code generation and optimization*, Washington DC, USA, 2003.

[19] Hex-Rays Decompiler. http://www.hex-rays.com/.

[20] J. K. Hollingsworth, B. P. Miller, and J. Cargille. Dynamic program instrumentation for scalable performance tools. Scalable High Performance Computing Conference, May 1994.

[21] IDAPro disassembler. http://www.hex-rays.com/idapro/.

[22] C. keung Luk and et. al. Ispike: A post-link optimizer for the intel itanium architecture. In *In IEEE/ACM International Symposium on Code Generation and Optimization*, pages 15–26, 2004.

[23] A. King and et. al. Analysis of executables: Benefits and challenges. *Dagstuhl Reports*, pages 100–116, 2012.

[24] A. Kotha and et al. Automatic parallelization in a binary rewriter. In *Proceedings of the International Symposium on Microarchitecture 2010*. ACM Press.

[25] J. Li, C. Wu, and W. Hsu. Dynamic register promotion of stack variables. In *Proceedings of the International symposium of Code generation and optimization,March 2011*.

[26] C.-K. Luk and et al. Pin: building customized program analysis tools with dynamic instrumentation. In *Proceedings of the ACM conference on Programming language design and implementation*, pages 190–200, 2005.

[27] D. Marino and et. al. A case for an sc-preserving compiler. In *Proceedings of the 32nd ACM SIGPLAN conference on Programming language design and implementation*, PLDI '11, pages 199–210, New York, NY, USA, 2011. ACM.

[28] T. Romer and et al. Instrumentation and optimization of win32/intel executables using etch. In *In Proceedings of the USENIX Windows NT Workshop*, 1997.

[29] A. V. S. Sastry and R. D. C. Ju. A new algorithm for scalar register promotion based on ssa form. *SIGPLAN Not.*, 33:15–25, May 1998.

[30] B. Schwarz and et. al. PLTO: A link-time optimizer for the intel ia-32 architecture. In *In Proc. Workshop on Binary Translation*, 2001.

[31] J. Seward and N. Nethercote. Valgrind, an open-source memory debugger for x86-linux. http://developer.kde.org/~sewardj/.

[32] M. Smithson and R. Barua. Binary Rewriting without Relocation Information. *USPTO patent pending no. 12/785,923*, May 2010.

[33] D. Song and et al. Bitblaze: A new approach to computer security via binary analysis. In *Proceedings of the 4th International Conference on Information Systems Security*, pages 1–25, Berlin, 2008.

[34] N. Vachharajani and et. al. Speculative decoupled software pipelining. In *PACT '07: Proceedings of the 16th International Conference on Parallel Architecture and Compilation Techniques*, pages 49–59, Washington, DC, USA, 2007.

[35] L. Van Put and et. al. Diablo: a reliable, retargetable and extensible link-time rewriting framework. In *Proceedings of the 2005 IEEE International Symposium On Signal Processing And Information Technology*, pages 7–12, Athens, December 2005.

[36] D. Wagner and et. al. A first step towards automated detection of buffer overrun vulnerabilities. In *In Network and Distributed System Security Symposium*, pages 3–17, 2000.

[37] X. Wang and et. al. Still: Exploit code detection via static taint and initialization analyses. In *Computer Security Applications Conference, Annual*, pages 289 –298, 2008.

[38] J. Zhang and et. al. Parameter and return-value analysis of binary executables. In *Proceedings of the 31st Annual International Computer Software and Applications Conference*, pages 501–508, Washington, DC, USA, 2007.

[39] C. Zilles and G. Sohi. Master/slave speculative parallelization. In *Proceedings of the 35th annual ACM international symposium on Microarchitecture*, pages 85–96, CA, USA, 2002.