Compiler assisted memory management for safety-critical hard-real time applications

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Arguably the major hurdle in adoption of memory managed languages, such as Java, in the real-time systems community is the incorporation of garbage collection (GC) within the real-time application. It has recently been shown that the statically estimated worst case execution time (WCET) value of tasks invoking GC is pessimistic to such a great extent that incorporating them within the hard real-time scheduling framework is infeasible. We develop a compiler assisted memory management technique for safety critical hard real-time applications developed in the synchronous subset of the formal globally asynchronous locally synchronous programming language called SystemJ. The SystemJ model of computation allows us to partition the heap into two distinct areas and perform compile time memory allocation. The new memory reclaim procedure is a simple pointer reset, which is guaranteed to complete within a bounded number of clock-cycles, thereby alleviating need for large pessimistic WCET bounds obtained due to GC cycle times. Other than being amenable to formal verification and tight WCET analysis, results show that our proposed approach to memory management is approximately $3\times$ faster compared to standard real-time GC approaches.

Categories and Subject Descriptors: D.3.4 [**Processors**]: Compilers

General Terms: Compiler, Static Analysis, WCET, SystemJ, Garbage Collection

Additional Key Words and Phrases: Wireless sensor networks, media access control, multi-channel, radio interference, time synchronization

ACM Reference Format:
ACM Trans. Embedd. Comput. Syst. , , Article (November 2014), 25 pages.
DOI:http://dx.doi.org/10.1145/0000000.0000000

1. INTRODUCTION

Synchronous programming languages [Berry and Gonthier 1992] allow building correct by construction safety critical hard real-time applications, because they are based on formal mathematical semantics, which allows the compiler itself to act as a model-checker [Jagadeesan et al. 1995]. The major benefit of the synchronous *Model of Computation* (MoC) is that it is based on the concept of a *Finite State Machine* (FSM). Especially, languages such as Esterel [Berry and Gonthier 1992] and its derivation SystemJ [Malik et al. 2010] have been designed to reduce the state space explosion problem exhibited during formal verification, by adhering to a very strict programmer specified state demarcation approach. The programming model for these languages can be described as follows: a synchronous program is quiescent until one or more events occur from the environment. Upon detecting the event, the synchronous program *re-*

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DOI:http://dx.doi.org/10.1145/0000000.0000000

acts instantaneously in zero time to respond to these inputs and becomes quiescent until the next input events arrive. The start and the end of this reaction transition demarcate states of the program – note that internal data updates do not lead to change in the program state thereby reducing the state space explosion problem. Furthermore, the reaction being logically in zero-time is by definition *atomic* and always faster than the incoming input events, thereby guaranteeing that none of the incoming events are missed. In reality, the reaction does take time, one needs to find out the *Worst Case Reaction Time* (WCRT) of any given synchronous program. This WCRT value determines the shortest inter-arrival time for input events from the environment.

Many techniques exist for computing WCRT values for synchronous programs providing support for data-computation via a non-managed language (primarily 'C') [Boldt et al. 2008; Roop et al. 2009]. These WCRT estimation approaches suffer from a fundamental problem; they cannot incorporate complex data-computations within the WCRT analysis framework, because of the many undefined behaviors, type unsafety and the general lack of formal operational semantics inherent to low-level non-managed languages. Hence, in this paper we are interested in the WCRT analysis of synchronous programs that support data computation via managed-languages like Java, which provides a well defined operational semantics. The key hurdle to the adoption of managed languages in the safety critical hard real-time application domain is *Garbage Collection* (GC) for automatic memory management. On the one hand GC guarantees no memory leaks, thereby providing functional safety, but on the other hand, it has been shown [Puffitsch 2013] that using even a state of the art real-time GC results in very pessimistic worst case execution time estimates for real-time tasks, [1] compared to the observed task execution times, thereby rendering worst case execution/reaction time analysis unfeasible in the general case with a GC.

The main **contribution** of this paper is to present a new memory management approach that is amicable to static WCRT analysis of synchronous programs intertwined with managed runtime environments for data-computation. More concretely our contributions can be refined as follows:

— *Programming model inspired memory organization*: In this paper we present a new memory organization for Java and an accompanying static WCRT analysis based on the synchronous MoC.
— *Compiler supported memory allocation*: We present the compiler transformations that are needed to statically guarantee the *Worst Case Memory Consumption* (WCMC) and *allocation*.
— *Real-time analyzable back-end code generation and garbage collection*: We present a strategy to replace the object allocation bytecodes with real-time analyzable alternatives. Furthermore, we also present an object placement strategy, which allows for O(1) heap accesses in all cases.

We use the SystemJ [Malik et al. 2010] programming language, because it provides the synchronous MoC as a subset along with Java data-driven computation support. Furthermore, there exist efficient and real-time analyzable execution architectures [Nadeem et al. 2011] and WCRT analysis tools [Li et al. 2014] that support subset of SystemJ programs; those without GC invocation.

The rest of the paper is arranged as follows: Section 2 gives the preliminary information needed to read the rest of the paper. Section 3 gives the motivating example to explain the SystemJ language and its WCRT analysis. Section 4 gives the overview of the new memory management scheme. Section 5 describes the main data-flow anal-

[1]The worst case execution time was estimated to be in seconds due to the GC, but the observed time was in milli-seconds.

noop	do nothing
pause	complete a logical tick
[**input**][**output**][**type**]**signal** σ	declare signal σ
emit σ[**value**]	broadcast signal σ
abort (σ) **s**	preempt statement s if σ is true
suspend (σ) **s**	suspend statement s for one tick if σ is true
present (σ) **s1 else s2**	do s1 if σ is true else do s2
s1; **s2**	do s1 and then s2
s1||**s2**	do s1 and s2 in lockstep parallel
while(**true**) **s**	do s forever
jterm	Java data term

Fig. 1: Syntax of the synchronous subset of the SystemJ language

ysis algorithm used for compile time memory allocation. Section 6 explains the back-end code generation procedure. Section 7 gives the quantitative results comparing the presented technique with real-time garbage collection alternatives. Section 8 gives a thorough comparison of the work described in this paper with current state of the art. Finally, we conclude in Section 9.

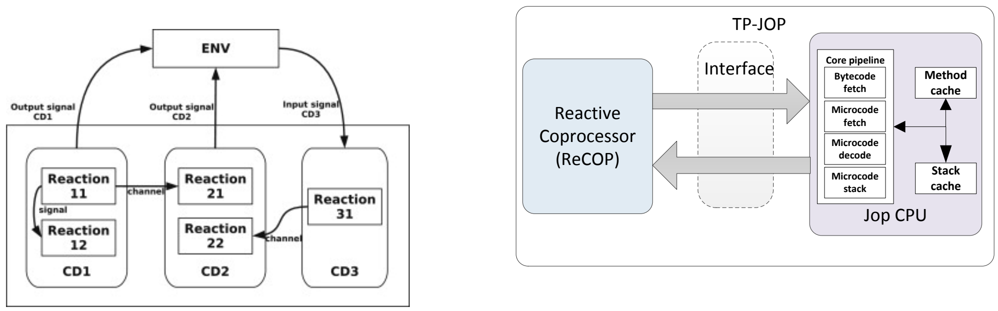
2. PRELIMINARIES

Before we present the key contributions of the paper, we dedicate this section to the description of the preliminary information needed by the reader.

2.1. The SystemJ programming language

SystemJ [Malik et al. 2010] extends the Java programming language by introducing both synchronous and asynchronous concurrency. SystemJ integrates the synchronous essence of the Esterel [Berry and Gonthier 1992] language with the asynchronous concurrency of *Communicating Sequential Processes* (CSP) [Hoare 1978]. SystemJ is grounded on rigorous mathematical semantics and thus is amenable to formal verification. We will use the graphical representation of a SystemJ program in Figure 2a for explaining the SystemJ MoC. On the top level, a SystemJ program comprises a set of clock-domains (CD1, CD2, CD3) executing asynchronously. Each clock-domain is a composition of one or more synchronous parallel reactions (Reaction11, Reaction12, etc), which are executed in lock-step with the logical tick of the clock-domain. Reactions within the same clock-domain exploit synchronous broadcast mechanism on signals to communicate with each other. Reactions can themselves be composed of more synchronous parallel reactions, thereby forming a hierarchy. Finally, every reaction in the clock-domain communicates with the environment through a set of interface input and output signals. At the beginning of each clock-domain tick the input signals are captured, a reaction function processes these input signals and emits the output signals. These output signals are presented to the environment at the end of the clock-domain tick. Inter-clock-domain communication is achieved by implementing CSP style rendezvous mechanism through point-to-point unidirectional channels. A complete list of SystemJ kernel statements is shown in Figure 1. In this work we are only interested in the synchronous subset of SystemJ, i.e., a single clock-domain.

Signals are the primary means of communication between reactions of a clock-domain and between a clock-domain and its environment. A signal in SystemJ can be typed or untyped. Every untyped signal in SystemJ is called a pure signal and only consists of a Boolean status, which can be set to `true` or `false` via signal emission. A typed signal is called a valued signal, and consists of a value (of any Java data-type) in addition to the status. The value of a typed signal can also be set, optionally, via the `emit` statement. Once emitted, the status of a signal is `true` only for a single tick, but the value of a signal is persistent over ticks. Emission of a signal, broadcasts it across



(a) A graphical representation of a SystemJ program

(b) Overview of the TP-JOP architecture

Fig. 2: A generic SystemJ program and its execution architecture

all reactions within the clock-domain. Moreover, the updated status and value of the signal, upon emission, is only visible in the next tick.

SystemJ provides mechanisms to describe reactivity via statements such as **present**, **abort**, and **suspend**, which change the control-flow of a SystemJ program depending upon signal statuses. The control-flow of a program can also be altered via standard Java conditional constructs. Synchronous (lock-step) parallel execution of reactions within a SystemJ program is captured using the synchronous parallel (||) operator [2]. Furthermore, shared variable communication between synchronous parallel reactions is not allowed in SystemJ. The shared memory communication model is replaced by the signal emission based communication mechanism. State boundaries are demarcated explicitly by programmers using the **pause** construct. Finally, there are two types of loops in SystemJ: temporal; consisting of a **pause** statement, loops that execute forever and bounded data-loops like in standard Java.

2.2. The real-time execution architecture

Every static WCRT analysis technique needs intimate knowledge of the hardware that executes the program. For our purpose, we choose a multi-core time predictable execution platform called *Tandem-Java Optimized Processor* (TP-JOP) [Salcic and Malik 2013; Nadeem et al. 2011]. There are two main reasons we chose this processor architecture: (1) efficiency – it has been shown previously that the TP-JOP processor architecture is much more efficient for executing SystemJ programs compared to general purpose Java processors [Park et al. 2014a] and (2) there already exists tools for statically estimating tight WCRT bounds for synchronous SystemJ programs on the TP-JOP architecture [Li et al. 2014].

Figure 2 gives an overview of the TP-JOP architecture. The architecture consists of two time predictable cores. The first core is termed *Reactive Co-processor* (ReCOP) that executes the control flow instructions of a SystemJ program. The Java data computations are dispatched to the JOP core on a as needed basis by the ReCOP. ReCOP has a multi-cycle data-path, and each native ReCOP instruction is executed in 3 clock cycles. ReCOP core uses a single small private memory for both: data and program.

The JOP [Schoberl 2003] core is a hardware implementation of the Java Virtual Machine (JVM). The JOP core is a four stage pipelined processor. Every Java bytecode is translated into one or more microcodes, which are the native instruction of the JOP

[2]Standard Java threading model is not allowed within SystemJ. Instead, synchronous and asynchronous parallel operators need to be used.

processor. The JOP pipeline has been designed to execute each microcode in one clock cycle and is guaranteed to be stall free. Furthermore, to guarantee time predictability, a novel cache architecture is implemented in JOP. There are two caches in JOP. The first is the stack cache that acts as a replacement for the data cache found in general purpose processors. The second is the method cache that acts as a replacement for program cache. A complete Java method is loaded into the method cache before execution and hence, there are only two program points: the *invoke* and the *return* bytecodes that can lead to a method cache miss, which can be accounted for in the static WCRT analysis procedure with ease. Finally, the interface connecting the ReCOP and JOP has single clock cycle latency as described in [Salcic and Malik 2013].

The overall program execution on the TP-JOP architecture proceeds as follows: the ReCOP leads the program execution since it executes the control-flow graph of the SystemJ program. When a data computation node is encountered, a call is dispatched to the JOP core with a method ID. Upon dispatch, the ReCOP itself stops processing further instructions until a result is returned from JOP. The JOP core polls continuously for an incoming request from ReCOP. Once a request is received, JOP decodes the method ID and loads the method into the method cache. Once loaded, the requested method is executed and upon completion of execution, the result is returned back to ReCOP.

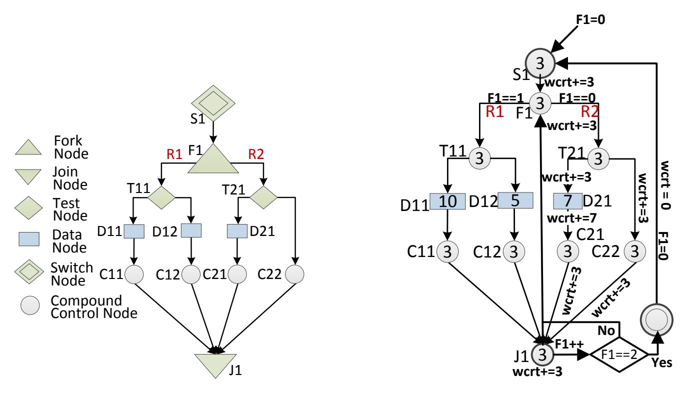
3. STATICALLY ESTIMATING WCRT OF SYSTEMJ CLOCK-DOMAINS

In this section we use a pedagogical SystemJ synchronous program to describe the static analysis technique used to estimate the WCRT of SystemJ clock-domains. The WCRT estimation technique is not the main contribution of this paper and hence, we only give a brief overview of the approach. The reader is referred to [Li et al. 2014] for a more detailed description of the technique.

Figure 3a shows a pedagogical SystemJ synchronous program. This program declares 2 input signals (lines 3 and 4) that are used to capture events generated from the external environment and 1 output signal (line 2) used to produce events back to the external environment. Two concurrent reactions (lines 6-16 and 18-26) execute in lock-step parallel, composed using the synchronous parallel operator (||) on line 17. Reaction R1 checks for an incoming event on the input signal INPUT_A (line 9). If an event

```
1   SysJ_Example(
2     output signal OUTPUT;
3     input Integer signal INPUT_A;
4     input Integer signal INPUT_B;
5    )->{
6    {
7      int count_a = 0;
8      while(true){
9        present (INPUT_A) {
10         if (#INPUT_A == 0)
11         //...data computation
12         else ++count_a;
13       } else --count_a;
14       pause;
15     }
16   }
17   ||
18   {
19     int count_b = 0;
20     while(true){
21       present(INPUT_B)
22         ++count_b;
23       else emit OUTPUT;
24       pause;
25     }
26   }
27  }
```

(a) Example synchronous SystemJ clock-domain



(b) GRC of the SystemJ clock-domain in Figure 3a

(c) Uppaal representation of GRC

Fig. 3: Example SystemJ program with its GRC and Uppaal representations

is detected, a counter is incremented, else, it is decremented. This logic is carried out forever using an infinite loop (line 8). The pause construct (line 14) explicitly demarcates the end of program transition. Similarly, a counter `count_b` is incremented upon reception of an event on signal `INPUT_B` in reaction R2, else signal `OUTPUT` is emitted to the external environment.

3.1. GRC representation of the clock-domain

The semantics of a synchronous program described in SystemJ is captured using an intermediate format called *GRaph Code* (GRC) [Malik et al. 2010]. The GRC is a directed acyclic graph that explicitly describes the control flow of a synchronous program through primitive nodes.

Definition 3.1. *Clock-domain transition (or tick)*: A single clock-domain transition also called a clock-domain tick is a single traversal of the GRC from the source node (node with no incoming edges) to the sink node (node with no outgoing edges).

Java statements in SystemJ are represented as data nodes. State encoding and selection are captured with enter and switch nodes, respectively. Present statements are denoted by test nodes. Concurrency embodied by parallel reactions is delineated using fork and join nodes. A fork node spawns parallel reactions, while a join node instructs that the parent reaction waits until all child reactions finish processing, which results in synchronized lock-step execution. Each node in the GRC is a schedulable unit, which can be either: (a) a java data node (JDN) representing data-driven computation implemented as Java methods, or (b) a control node (CN) representing all nodes other than JDNs, implemented on the ReCOP. All control nodes between two data nodes are grouped together to form a single compound control node in order to simplify the GRC. The GRC of the SystemJ program shown in Figure 3a is depicted in Figure 3b. The fork node F1 creates two parallel reactions R1 and R2 at the start of the program. The present statements (lines 9 and 21) are represented as nodes T11 and T21, respectively in the GRC. The Java data computations at lines 10-12, 13, and 22 are represented as nodes D11, D12, and D22, respectively. Finally, the infinite loops are simply infinite executions of the GRC

3.2. WCRT estimation using the Uppaal model-checker

Finding the WCRT of a SystemJ clock-domain is equivalent to finding the longest execution path from source node to the sink node in the GRC. Finding this longest path is non-trivial, because of concurrent execution of the synchronous parallel reactions, the execution of different control branches depending upon the incoming input events from the external environment, and the state decoding within the switch nodes. We use the Uppaal [Behrmann et al. 2004] model-checker to perform an exhaustive state space exploration of the GRC, in the process emulating the execution of the GRC on the TP-JOP architecture to find the WCRT of the SystemJ clock-domain.

There are three steps involved in finding the WCRT of a clock-domain using Uppaal:

(1) *Annotating the GRC nodes with Worst Case Execution Time (WCET) values*: The GRC is annotated with timing information in order to compute the WCRT of the clock-domain. All CNs are executed on the ReCOP and hence, their timing information can be easily computed, as all native ReCOP instructions take 3 clock cycles to execute. The JDNs are dispatched, encapsulated within individual Java methods, for execution on JOP. One needs to compute the *Worst Case Execution Time* (WCET) for each of these JDNs, using the JOP provided worst case analysis tool (WCA) [Schoeberl 2005]. The computed worst case execution times for both: CNs

and JDNs are back annotated onto the GRC. The numerical annotations on the nodes in Figure 3c show these back annotated WCETs.

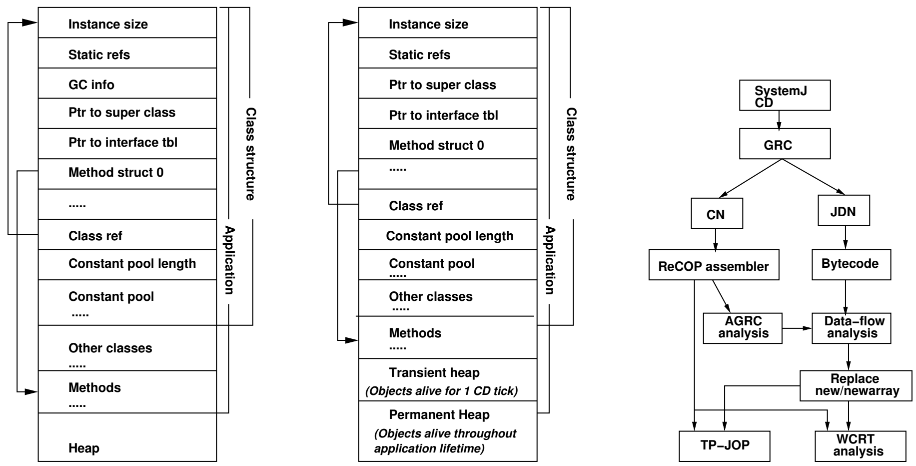
(2) *Translating the GRC into an Uppaal timed automaton*: Figure 3c shows the GRC of the pedagogical SystemJ clock-domain captured as a *Timed Automaton* (TA) in Uppaal. All nodes and edges in the GRC are mapped as locations and edges in the TA. Given the TP-JOP architecture, we know that all control nodes are executed only on a single ReCOP. Consequently, all synchronous parallel reactions can only be executed sequentially. In order to enforce this sequential execution of the synchronous parallel reactions, we introduce extra integer type variables for each fork node, e.g., $F1$ in Figure 3c. These fork variables are initialized to 0. Furthermore, a back-edge is added from the corresponding join node to the fork node, which increments the fork variable ($F1$ in Figure 3c). Upon reaching the join-node, during state space exploration, the back-edge is taken as long as the value of the fork variable is not equal to the number of synchronous parallel reactions (2 in Figure 3c). Thereby, guaranteeing sequential execution of the synchronous parallel reactions. An integer type variable $wcrt$, is added to the TA. No additions are needed for the test nodes, since the model-checker implicitly backtracks and explores all branches of the test nodes. The $wcrt$ variable is incremented by the WCET of individual nodes in the TA at every edge. Finally, to emulate the execution of multiple ticks, and consequently explore all branches of the switch nodes, a back-edge is also added from a dummy sink node back to the source node. The $wcrt$ and fork values are reset upon execution of this back-edge.

(3) *Using a Computational Tree Logic (CTL) property to obtain the WCRT*: We model the WCRT estimation problem as verifying a CTL property in the Uppaal model-checker. The WCRT value lies in the bounded range denoted by $[WCRT_{lb}, WCRT_{ub}]$. $WCRT_{ub}$ is a safe upper bound on WCRT obtained by applying Max-Plus algebra to the GRC, which is essentially summing up the maximum tick execution time of every reaction in the clock-domain. Similarly, we calculate a lower-bound of WCRT, termed $WCRT_{lb}$, by summing up the minimum tick execution time of every reaction in the clock-domain. After translating GRC to TA, we check the validity of the WCRT estimate of the clock-domain, termed $WCRT_{est}$, between $[WCRT_{lb}, WCRT_{ub}]$ by verifying a CTL property upon the TA using the Uppaal model checker. The CTL property is written as $A[](wcrt \leq WCRT_{est})$, meaning that the value of $wcrt$ variable in the TA is less or equal to the $WCRT_{est}$ for every path in the TA starting from the initial location. Upon violation of this property we are guaranteed that we have found the tightest $WCRT_{est}$ value.

4. A NEW MEMORY MANAGEMENT APPROACH FOR SYSTEMJ

The WCRT approach described above is suitable *only* for programs that do not invoke *Garbage Collection* (GC), since the GC execution time is unaccounted for in the aforementioned WCRT analysis framework. The main reason for this lack of GC incorporation is that the GC cycle time cannot be *practically* bounded; as the number of heap allocations in a Java program depends upon the application and during garbage collection a linked list of allocated objects needs to be traversed as one of the collection phases. Puffitsch [2013] shows that a WCRT value can be obtained for programs, which invoke the GC, by pessimistically bounding the collection phase during static analysis. But, the resultant WCRT value is in seconds, orders of magnitude larger than real execution time (usually in micro-seconds), which makes garbage collection unsuitable for hard real-time applications.

Our solution to the above problem is to eliminate garbage collection altogether. More concretely, we divide the heap space into two parts: a bounded permanent heap and a



(a) Original JOP runtime memory organization

(b) Proposed JOP runtime memory organization

(c) The tool-chain flow

Fig. 4: The tool-chain flow along with the Java runtime memory organization in JOP, before and after transformation. The Java stack is on separate physical memory as shown in Figure 2.

transient heap with bump pointer based object allocation and a simple pointer reset based memory reclaim, which are both, time and space bounded and efficient.

Signals (valued or pure) are the only communication primitive within a SystemJ clock-domain. Signals, internally themselves implemented as Java objects, are alive throughout the lifetime of the application. The proposed heap organization is based on one **key insight**; there are two types of Java objects in a SystemJ clock-domain:

— *Permanent objects*: Java objects that are emitted via signals. These objects like signals are potentially alive throughout the application lifetime.
— *Transient objects*: Java objects that are created internally for computation, but are never emitted via signals. Transient objects are only alive during a single clock-domain transition and can be garbage collected at the end of the transition.

The basic idea is to allocate the permanent objects, we consider signals and the values associated with them to be permanent objects as well, within a special memory area called *permanent heap* and allocate all transient objects within a *transient heap*. The transient heap gets reclaimed at the end of every clock-domain tick transition, by simply resetting the allocation pointer to the start of the transient heap space. With this change, the proposed runtime Java memory organization is compared with the original RTGC based memory organization in Figure 4.

Both runtime memory organizations; original (Figure 4a) and proposed (Figure 4b) consists of a number of common elements required to execute every Java application such as, the constant pool, method structure table for instance method invocation, references to static objects, etc. The major point of difference is that the *GC info* word does not exist in the proposed runtime memory organization and the *heap* space is divided into two parts; the transient heap and the permanent heap. Looking at the proposed runtime memory organization two requirements become very important:

(1) There should be no pointer pointing from the permanent heap to the transient heap, else an application might terminate at runtime with a `NullPointer` exception – since the transient heap may be reclaimed or overwritten at the end of the clock-

```
1  int signal S;
2  int a = 0;
3  a = a + 1;
4  emit S(a);
5  pause;
6  /*Variable values can
7   only be carried
8   over tick boundaries
9   via signals */
10 a = #S;
11 a = a + 1;
12 emit S(a);
```

(a) SystemJ code snippet

```
public class FruitSorterController {
 //signal S declaration
 private static Signal S;
 private static int a;
 public static void init () {
  S = new Signal();
 }
 public static void main(String args){
  init();
  while(true){
   //poll on the method call request from ReCOP
   int methodnum = Native.rd(Const.METHOD_NUM);
   switch(methodnum){
    case 0: ...
    ...
    //Call method for
    //lines 2 - 4 (Figure~5a)
    case 1: Native.wr(RESULT_REG,
    MethodCall1_0());
    //Call method for
    //lines 11 - 12 (Figure 5a)
    case 2: Native.wr(RESULT_REG,
    MethodCall2_0());
   }
  }
 }
 private static boolean MethodCall1_0(){
   a = 0;//line 2, Figure 5a
   a = a + 1;//line 3, Figure 5a
   S.setValue(new Integer(a));
 }
 private static boolean MethodCall2_0(){
   a = S.getPreValue(); //get value from previous tick via S
   a = a + 1;
   S.setValue(new Integer(a));
 }
}
```

(b) Produced Java code

Fig. 5: Example showing the need for GRC and data-flow analysis of a SystemJ program

domain transition. Or in the much worse scenario, an incorrect value might be read, in which case the application will be functionally incorrect. Both these scenarios are unacceptable for safety-critical systems.

(2) The permanent heap is never reclaimed and hence, permanent heap space should be judiciously allocated. In fact, we need to bound the permanent heap space at compile time.

In the rest of the paper we describe compiler transformations that accomplish the described memory management technique.

5. STATIC ANALYSIS FOR COMPILE TIME MEMORY ALLOCATION

The complete compiler tool-chain flow in shown in Figure 4c. The SystemJ clock-domain is first compiled into the intermediate GRC format. Next, the control nodes (CN) and the Java data nodes (JDN) are split and compiled separately into ReCOP native assembly and Java bytecode, respectively, for execution on the TP-JOP platform. Two types of compiler analysis are then performed on the produced native ReCOP assembler and Java bytecode, respectively: (1) GRC analysis is used to determine the *must* and *may* Java reachable methods from any given Java program point and (2) *forward* and *backward* data-flow analysis is performed to find the objects and arrays that need to be allocated to the permanent heap. Finally, these object allocation bytecodes need to be replaced by their time predictable alternatives.

The need for these two types of analysis can be explained using the simple example in Figure 5a. This SystemJ code snippet declares a valued signal `S` (line 1), and an integer variable `a` (line 2). Variable `a` is incremented (line 3) and emitted via signal `S` (line 4). After emission, the tick expires as indicated by the `pause` statement. In the second clock-domain transition, variable `a` is initialized again, since in SystemJ the only values that are persistent across ticks are signal values. Variable `a` is first initialized to the value held in signal `S` (from the previous tick) and then incremented (lines 10-11). The result of this increment is emitted via signal `S`. Two methods are produced in the back-end Java code for execution on JOP (Figure 5b). The first method `MethodCall1_0` (lines 27-31) initializes `a` and then increments it. Finally, it sets the value of the signal `S` as the *object* `a`. The second method call `MethodCall2_0` (lines 32-36) increments the variable `a` and changes the object reference of signal `S`.

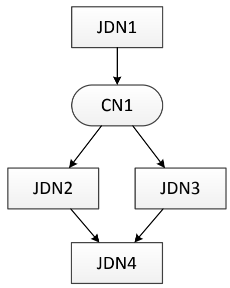
Our objective is to find out all the *program-points* that allocate a new object, whose returned reference is set as a signal value via the `setValue` virtual method call on the signal object. We perform data-flow analysis (ReachDef analysis [Muchnick 1997] to be exact) to find out such program points in the Java program generated for execution on JOP (e.g., Figure 5b). Fields or method local object references might be set as signal values and hence, our data-flow analysis is a context and flow-sensitive intra-procedural and inter-procedural analysis, i.e., the data-flow analysis not only examines each Java method, but also traverses the call-graph of the generated Java program.

A call-graph generated from *just* the Java program is incomplete, since method calls that *must happen before* a given program point *appear* as *may happen before* a given program point. Consider the SystemJ program in Figure 5a, we know for sure that program point at line 3 *must* be executed before program point at line 11. But, in the generated Java program (Figure 5b), the sequential control-flow of the *SystemJ program* has been turned into branching control-flow (lines 13-24). Just looking at the Java program, one can only state with certainty that method `MethodCall1_0` *may* be called before method `MethodCall2_0`. More disconcertingly, just looking at the Java program, one can even say that method `MethodCall2_0` may be called before `MethodCall1_0`, since the dependence edge, which is clear in the SystemJ program is lost in the generated Java program. The GRC analysis step produces the may and must method callers for any callee in the generated Java code. Note that we cannot produce a single Java method call coalescing the two Java methods: `MethodCall1_0` and `MethodCall2_0`, because there is an intertwined control construct, `pause`, which needs to be executed on ReCOP. We describe these GRC and the data-flow analysis steps in the next sections.

5.1. GRC analysis

GRC is a *directed graph* $G = (V, E)$, where V is the set of vertices (nodes of GRC) and E is the set of ordered pairs of vertices. Each edge $e = (v_i, v_j), \forall e \in E, v_i \in V, v_j \in V$ is a tuple denoting that the edge is directed from v_i to v_j. Then a lambda function $\lambda : v \rightarrow E_i$, $E_i \subseteq E$, maps *each* vertex to a set of edge(s) where $\forall (v_i, v) \in E_i$ and E_i may be $\emptyset$. When $|E_i| = 1$ we call v_i *must* happen before v whereas when $|E_i| > 1$ we say any $v_i \in E_i$ *may* happen before v.

In this section, we illustrate how the data-structure, which contains information on the *may* and *must* relationships between the nodes, is obtained from the GRC. Consider a snippet of GRC graph shown in Figure 6a. Here, the root node, `JDN1`, has no parent whereas it has an immediate child `CN1`. `CN1` has two children, `JDN2` and `JDN3`, and both of these JDNs have the same child `JDN4`. This graph when given as an input to our GRC analysis tool, returns the result as a S-expression as shown in Figure 6b. The resultant recursive data-structure consists of a set of nodes called `mNode`, which has three fields: (1) the node name of type string, (2) `Must` field of type `mNode` and (3) `May` field, which is a set of type `mNode`. This data-structure can be used to identify potential callers of each `JDN` in the GRC graph. For example, since `JDN1` has no parents, it's `Must` and `May` fields are both `Null` (line 2). On the other hand, `JDN1` must be called before `JDN2` or `JDN3`, hence `Must` of both `JDN2` and `JDN3` is `JDN1` (lines 4 and 7). Lastly, `JDN4` has two parent nodes; `JDN2` and `JDN3`. Therefore, its `May` field has both `JDN2` and `JDN3` (lines 11-16). It should be noted that we are only interested in callers of `JDNs`, hence `CNs` are



(a) An abstracted GRC graph

```
(mList
 (mNode JDN1 (Must Null) (May Null))
 (mNode JDN2
  (Must (mNode JDN1 (Must Null) (May Null)))
  (May Null))
 (mNode JDN3
  (Must (mNode JDN1 (Must Null) (May Null)))
  (May Null))
 (mNode JDN4 (Must Null)
  (May
   (mNode JDN2
    (Must (mNode JDN1 (Must Null) (May Null)))
    (May Null))
   (mNode JDN3
    (Must (mNode JDN1 (Must Null) (May Null)))
    (May Null))))
```

(b) An output of the GRC analysis

Fig. 6: An example of GRC analysis

not included in the final result (Figure 6b). A complete algorithm of our GRC analysis is given in Appendix B.

5.2. Data-flow analysis

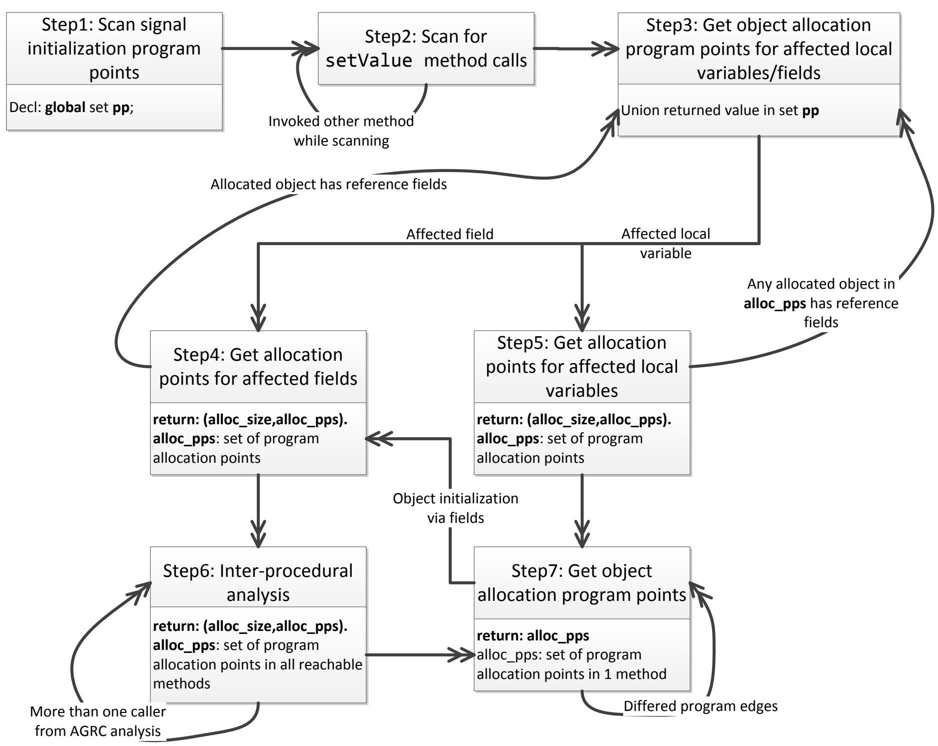


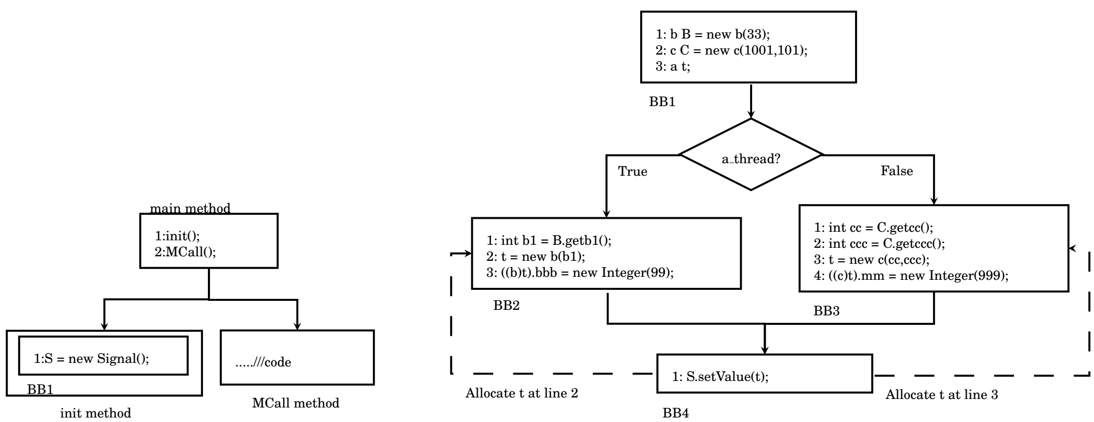
Fig. 7: The flow-chart for the data-flow analysis algorithm

Given the class file(s), the data-flow analysis carried out is shown in Figure 7. We use the flow-chart in Figure 7 to describe the steps of our data-flow analysis procedure. The detailed algorithm for each step is provided in Appendix A. Being a complex procedure we use simple code examples to describe the data-flow analysis procedure.

We divide the presentation of our data-flow analysis procedure into two sections. Section 5.3 presents the *Intra-procedural* data-flow analysis, which is complemented with the *Inter-procedural* data-flow analysis in Section 5.4.

5.3. Intra-procedural analysis

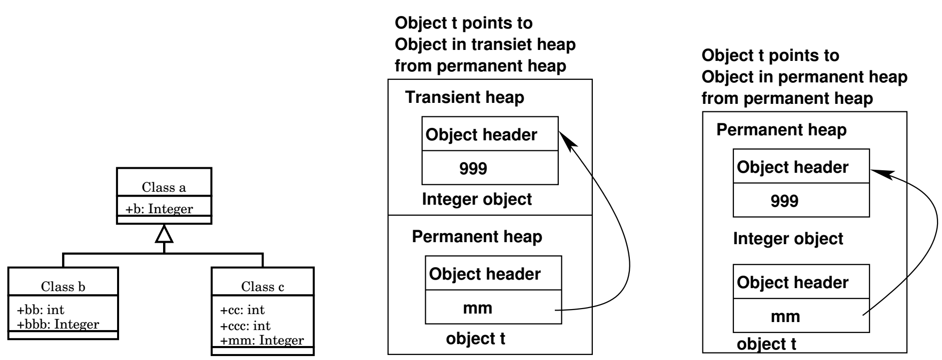
Figure 8a shows a Java program as a call-graph. There are three methods: `main`, `init`, and `MCall`. The `main` method calls the `init` and the `MCall` methods sequentially. The `init` method initializes a static object reference `S` of type `Signal`. Figure 8b shows the control-flow-graph (CFG) of the `MCall` method. It consists of four basic blocks annotated as BB1, BB2, BB3, and BB4, respectively. BB1 initializes objects `B` and `C`, whose



(a) The call-graph for the running example

(b) The control-flow graph of the MCall method

Fig. 8: Running example used to explain the intra-procedural data-flow analysis



(a) UML class hierarchy for classes `a`, `b` and `c` used in Figure 8b

(b) Pointers from permanent heap to transient heap

(c) Points-to only in permanent heap

Fig. 9: The UML representation of the class hierarchy used in Figure 8 and the points-to problem

types `b` and `c` are inherited from a single parent class `a` as shown in Figure 9a. Object `t` (BB1, line 3) of type `a` is initialized to either type `b` or `c` (BB2, line 2, BB3, line 3) depending upon the value of the Boolean `a_thread`. Finally, object `t` is emitted via signal `S` (BB4, line 1).

★ Step1 of the data-flow analysis algorithm starts by scanning the Java program for program points allocating signals (e.g., method `init` in Figure 8a). A globally accessible set pp is defined, which holds all the program points allocating objects that will be placed in the permanent heap. Step1 scans the call graph from the starting program point (usually the `main` method) for signal allocation and places these program points into the set pp. Each individual tuple within the set pp consists of a one or more program points that allocate objects. After the scan phase, the set pp holds $\{(5, \{\texttt{init} : \texttt{BB1} : 1\})\}$ for the example in Figure 8a. The first element of the tuple gives the *instance* size of the `Signal` class, the second element is the allocation program point [3] for the example in Figure 8a. If more than one signal is allocated, all these program points are placed in the set pp. Finally, Step2 is called to find all the signal emission program points.

★ Step2 is a recursive procedure that scans all reachable methods, in the call-graph, for the `setValue` virtual method call on the signal objects. These program points are placed in the set PPC for each reachable method individually. For the example in Figure 8, this involves scanning method `MCall` and placing the program point: 'BB4:line 1', in set PPC. Finally, Step3 of the algorithm is called to obtain all object allocation program points whose returned references are emitted via signals.

★ Step3 analyzes the emission expression for every program point in set PPC. An emitting expression can *only* emit a method local variable or a field. In our current example a local variable (`t`) in Figure 8b, BB4, line 1, is being emitted (or formally called being affected) and hence, Step5 is called.

★ Step5 extracts the variable from the affected expression and calls Step7 for analysis. Step7 is the core of the whole data-flow analysis procedure that performs reachability analysis to identify the program points initializing the objects to be placed in the permanent heap. Before proceeding to Step7, let us assume that we have already obtained the correct program points from Step7 in set $alloc\_pps$. For our current example, given that the emitting program point is 'BB4:line 1', $alloc\_pps = \{\texttt{MCall} : \texttt{BB2} : 2, \texttt{MCall} : \texttt{BB3} : 3\}$. Looking at Figure 8b, it is clear that object reference `t` can be initialized at either in BB2 line 2 or BB3 line 3, depending upon the value of variable `a_thread` during program execution. For each of these program points, Step5 guarantees that the object reference does not *escape* the lifetime of the method [Choi et al. 1999]. [4] If the object reference does not escape the lifetime of the method, then the algorithm checks if there are any reference fields within this object. If so, Step3 is invoked for migrating all object allocations, whose returned reference is held in these fields, into the permanent heap.
Consider the program point in BB3 at line 3. Object `t` is initialized as class `c` (for program point in BB2 line 2, `t` is initialized and analyzed as class `b`), which has three fields (c.f. Figure 9a): two (`cc` and `ccc`) are primitive fields, but the third: `mm` is a reference field of type `Integer`. The `mm` reference field is initialized to an `Integer` object in BB3 line 4. Figure 9b shows the address held in `mm` if Step3 is not called from Step5. Reference field `mm` points to an object in transient heap space, which violates the first requirement in Section 4. Calling Step3, guarantees that

[3] The allocation program point is actually at the bytecode level, but in this paper we keep the program points at the Java source code level for ease of understanding.
[4] Let O be an object reference and M be a method invocation. O is said to escape M, if the lifetime of O may exceed the lifetime of M.

all potential *pointed-to* objects from the permanent heap are also migrated to the permanent-heap. Note that Step3 and Step5 are mutually recursive and hence, perform a chained permanent heap migration. For the running example, the result of calling Step3 from Step5 is shown in Figure 9c.
The final computation that Step5 performs is computing the size, which will be reserved in the permanent heap, of object `t`, given that $alloc\_pps$ may have multiple (two in our current example) elements. The primary idea is to reserve the maximum size from amongst all the different types being initialized. For the running example, `t` can be of type `b` or `c` during program execution. We compute the *instance* size of both these types as 3 and 4 words, respectively [5] and reserve 4 words (plus object header size) in the permanent heap space. Step7 returns multiple program points in set $alloc\_pps$ iff these program points are mutually unreachable. This, guarantees that the reserved permanent heap space can be *reused* at runtime by all program points in set $alloc\_pps$.

⊞ **Design decision**: We have disallowed method local object references from escaping as a trade-off between space utilization and functional correctness. If we allow, for example `t` in Figure 8b, to escape method `MCall`, then any of the reference fields of `t` (e.g., `mm`) might be reinitialized at some other program point. This program point would also then need to be considered and space on the permanent heap would need to be reserved for the object allocation. This can result in the permanent heap becoming too large. Conversely, if the escape analysis is not performed, then there will be pointers pointing from the permanent heap to transient heap, which may result in potentially incorrect program behavior. Simply not allowing any object reference to exceed the lifetime of the method avoids both these problems.

★ Step7, given a program point that affects a variable (also termed the *use* program point) gives one or more program points initializing that variable (also called the defining program point). Step7 first builds the standard *use-def* chains [Muchnick 1997] to obtain the program points defining the variable. If the defining program points are `new` or `newarray` bytecodes [Meyer and Downing 1997], then these program points are simply unioned into the set of program points that will be returned by this function. If the defining program points are themselves affecting variables, then Step7 is called recursively to follow the so called deferred program edges [Burke et al. 1995] to finally reach one or more program points that allocates the object or if the variable is not initialized then gives a compile time `not initialized` error. Note, that this amounts to carrying out a complete alias analysis. If the defining program point affects a field, then Step4 is called, which calls Step7 via Step6, we describe Step4 and Step6 in the next section. If the defining program point is anything but any of the aforementioned statements, then an error is raised, stating that the variable might be initialized outside the method being analyzed. Finally, Step7 makes sure that the program points being returned are not reachable from each other.

⊞ **Design decision**: We have consciously decided to not allow object allocation outside the method that uses that object in Step7. The objective of our dataflow algorithm is to identify all the program points that allocate objects whose returned references are emitted as signal values. We are performing compile time memory allocation, which, as we will see later, requires us to replace the

[5]According to Java semantics, all fields of a parent class are inherited by the children classes and hence, the sizes for `b` and `c` are 3 and 4 words, respectively.

```
1  private static Signal S;
2  public static void main(String args){
3   Integer A = 0;
4   Integer B = 1;
5   S.setValue(A);
6  }
```

(a) Example Java code snippet

```
1  iconst_0
2  invokestatic  #2// valueOf:(I)Ljava/lang/Integer;
3  astore_1
4  iconst_1
5  invokestatic  #2// valueOf:(I)Ljava/lang/Integer;
6  astore_2
7  return
```

(b) Produced Java bytecode

```
1  new #3 // class java/lang/Integer
2  dup
3  iload_0
4  invokespecial #10 // Method "<init>":(I)V
5  areturn
```

(c) The bytecode for method `valueOf` in Integer class

Fig. 10: Inadvertent object sharing problem

`new` and `newarray` bytecodes with a different set of bytecodes (c.f. Section 6). We cannot blindly replace these bytecodes in methods other than the ones being analyzed, because, consequently *any* other program point, related or unrelated to signals, using the same object initialization program point would overwrite the object value as it inadvertently shares the same object. A simple example elucidating the situation is shown in Figure 10.

In Figure 10a, we initialize two `Integer` object type variables; `A` and `B` to constant `int` values. Then we emit `A` via signal `S`. The bytecode produced from this Java program is shown in Figure 10b. First, the constant 0 is loaded onto the stack and the method `valueOf` in the `Integer` class, from the Java standard library, is invoked, which allocates memory and initializes the returned reference field to 0, the returned reference is then stored on the stack (lines 1-3), same is done for the object `B` in lines 4-6. The actual object initialization is carried out inside the `valueOf` method, Figure 10c, line 1. If we were to replace this `new` bytecode, to allocate some memory in the permanent heap, then both; `A` and `B` will point to the same place in permanent heap. Consequently, first a constant 0 will be written in the allocated memory (Figure 10b line 3) and then this value will be overwritten to 1 (Figure 10b, line 5), resulting in functionally incorrect code. The *only* solution to this problem is to *inline* the called method, `valueOf` in this case. But, extreme caution is required when in-lining methods, because the method might be too large to fit in the method cache, or more than one such methods might need to be in-lined (in case a method itself calls another method, which does the actual object allocation and so on and so forth). To put the cache size restriction into perspective, all our benchmarks have a very limited cache size of only 1KB. Thus, we have decided to disallow, object allocation outside of the method being analyzed. As a consequence of this design decision, the SystemJ programmer now needs to make explicit deep copies as shown in Figure 8b BB2 line 3 and BB3 line 4.

5.4. Inter-procedural analysis

Until now we have only described those parts of our data-flow analysis that handle local method references being emitted via signals. Step4 and Step6 (c.f. Figure 7) described in this section work in conjunction with the previously described steps in order to handle cases where Java field references may be affected by signal emissions. We again use a simple example shown in Figure 11 to describe Step4 and Step6 of our data-flow analysis.

★ Step4 is the dual of Step5. It works on Java fields rather than method local variables like Step5. The other difference between the two is that Step5 calls Step7 via Step6, which performs inter-procedural analysis.

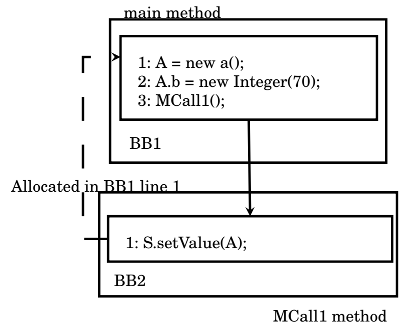


Fig. 11: Inter-procedural analysis with fields. `A` and `S` are static fields

★ Step6 can be conceptually partitioned into two parts. Part-1 returns the program points allocating objects within the same method that the field is being used. The more interesting part is part-2, which carries out *inter-procedural* analysis if the affected field being analyzed is not allocated in the same method.

Let us consider the Java program call-graph in Figure 11 to explain the inter-procedural analysis. The `main` method initializes two objects: `A`, a static field, of type `a` (c.f. Figure 9a) and its `Integer` type reference field `b`. Next, `main` calls method `MCall1`. Field `A` is emitted via signal `S` in method `MCall1`. Our data-flow analysis algorithm needs to locate the program points in BB1 lines 1 and 2, so that these object allocations can be moved to the permanent heap space.

Part-1 of Step6 calls Step7 passing it method `MCall1` to find out the object `A` and its field's allocation program points. Step7 being an *intra-procedural* analysis step returns back an empty list ($alloc\_pps = \emptyset$). In such a case, Step6 looks up the call-graph tree to find out the caller methods for the current callee, this lookup procedure also includes looking up all the potential callers indicated by the GRC-analysis (c.f. Section 5.1). Once the caller method is identified in the call-graph, Step6 recursively calls itself to analyze the caller. This procedure is continued until the program point allocating the affected field is identified or the analysis reaches the top of the call graph, in the latter case a `Not initialized field` error is thrown by the compiler. In case of Figure 11, there is a single caller: the `main` method in the call-graph for callee `MCall1`, which is analyzed to find the program point allocating field `A`, this program point is returned by Step6. Note that multiple callers might call the same callee, due to *may* happen before results produced by the GRC-analysis. In such cases, all the callers need to be analyzed in order to identify the program points allocating the affected field and to make sure that such allocations exist in all paths, else, the field would be uninitialized in some run of the SystemJ program.

This finishes the treatment of the data-flow analysis algorithm to identify the program points that return an object allocation reference, which may be emitted via signals. Now that we have identified these program points, we next give the procedure to generate the back-end code that: (1) replaces the object allocation bytecodes with time predictable alternatives and (2) performs compile time memory allocation in the permanent heap space.

6. BACK-END CODE GENERATION

The back-end code generation is a two step-procedure as shown in Figures 12a and 12b, respectively. The memory allocation algorithm (Figure 12a) takes as input the maximum memory address, the program points to replace and the whole program itself as input. Two variables: $allocPtr$ and $header\_size$ are initialized to the maximum memory

```
Input: maxMem: Maximum memory address
/* defs is the same set as pp in Figure 7 */
Input: defs: The program points to replace
Input: program: The program
1  let allocPtr ← maxMem;
2  let header_size ← 2;
3  foreach (size, dd) ∈ defs do
4  |  let dd ← sortUniqueDescending (dd);
5  |  let M ← load (dd);
6  |  let
   |  allocPtr ← allocPtr − (size + header_size);
7  |  if allocPtr < 0 then raise No_mem;
8  |  foreach d ∈ dd do
9  |  |  replaceByteCode (allocPtr, d, M);
10 |  end
11 end
```

(a) Bump-pointer memory allocation

```
Input: allocPtr: allocation pointer
Input: pp: program point to replace
Input: M: Method containing the program
       point
1  if pp = new then
2  |  let arg ← getArg (pp);
3  |  replaceNew (pp,M,arg);
4  |  adjustConditionals (M,22);
5  else if pp = newarray then
6  |  let arg ← getArg (pp);
7  |  if arg = primitive then
8  |  |  replaceNewArrayB (pp,M,arg);
9  |  else replaceNewArrayO (pp,M,arg);
10 |  adjustConditionals (M,22);
11 else raise error
```

(b) Replacing the `new`, `newarray`, and `anewarray` bytecodes

Fig. 12: The back-end code generation pseudo-algorithm

address and the constant two, respectively. The $allocPtr$ is decremented by the size of the object to be allocated (obtained from the data-flow analysis algorithm, Figure 7) plus, the object header size (line 6) – this simple *bumping* of the allocation pointer is termed bump pointer memory allocation. If the resultant $allocPtr$ value is less than 0, then we have exhausted total permanent memory allocation and correspondingly an error is raised (line 7). If there is enough memory available then, the object allocation bytecode at the program point is replaced by alternative bytecodes in the second step, shown in Figure 12b.

The algorithm (Figure 12b) used to replace the object allocation bytecodes takes as input three arguments: the allocation pointer, $allocPtr$, the program point to replace pp, and the method containing the program point: M. There are *two* types of object allocation bytecodes as specified in the *Java Virtual Machine* (JVM) specification [Meyer and Downing 1997]: the `new` and `newarray` bytecodes. Moreover, the `newarray` bytecode itself is of two varieties, dependent upon the type of array being allocated: (1) the `newarray` bytecode that allocates space for arrays holding primitive values and (2) the `newarray` bytecode that allocates space for arrays holding object references – called `anewarray`. This distinction is important since the number of bytes used to represent each of these bytecode varies. The `new` and `anewarray` bytecodes are encoded in the class file with three bytes, whereas the `newarray` bytecode only uses two bytes in the class file representation.

Different methods are called to replace each of these allocation variants: Figure 12b, line 3 replacing `new`, line 8 replacing `newarray`, and line 9 replacing `anewarray` bytecodes, respectively. All methods have the same logic, except that some padding (represented as `nop` bytecodes) is needed in case of the `newarray` and `anewarray` bytecodes. This padding is needed to correctly update the conditional instruction's target address after replacement (line 10).

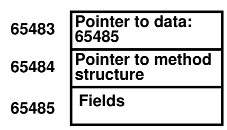
All conditional instruction's, following the object allocation bytecodes, target addresses need to be incremented by 22 bytes after replacement. This 22 byte increment is made obvious by Figures 13a and 13b, which show the Java bytecode snippets for the example program in Figure 11 before and after `new` bytecode replacement, respec-

```
--- constant-pool snippet ----
#2 = Class     #79     //a/a
#3 = Methodref #2.#78  //a/a.''<init>'':()V
#4 = Fieldref  #12.#80 //test/test.A:La/a;
---------- main method snippet --------
0:new #2           // class a/a
3:dup
4:invokespecial #3 // Method a/a.''<init>'':()V
7:putstatic     #4 // Field A:La/a;
```

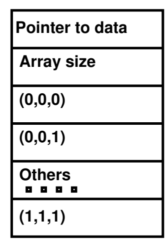
(a) Java bytecode snippet and class constant-pool for example in Figure 11

```
1  --- constant-pool snippet ----
2  #2 = Class     #79     //a/a
3  #3 = Methodref #2.#78  //a/a.''<init>'':()V
4  #4 = Fieldref  #12.#80 //test/test.A:La/a;
5  .... //other constants
6  #154 = Integer           65483
7  #155 = Integer           65485
8  ---------- main method snippet --------
9   0:ldc_w         #154 // int 65483
10  3:dup
11  4:ldc_w         #155 // int 65485
12  7:swap
13  8:invokestatic  #151 // Method Native.wr:(II)V
14 11:dup
15 12:bipush        1
16 14:iadd
17 15:ldc_w         #2   // class a/a
18 18:bipush        5
19 20:iadd
20 21:swap
21 22:invokestatic  #151 // Method Native.wr:(II)V
22 25:dup
23 26:invokespecial #3   // Method a/a.''<init>'':()V
24 29:putstatic     #4   // Field A:La/a;
```

(b) Java bytecode snippet and class constant-pool for example in Figure 11 after `new` bytecode replacement



(c) Java object allocation in permanent heap



(d) Java (2x2x2) array allocation in permanent heap

Fig. 13: The back-end generated code and compacted object representation in heap

tively [6]. Figure 13a shows the bytecode produced by the Java compiler for the `A` object allocation in the `main` method. The very first bytecode, `new`, takes as argument the index, 2, into the constant pool that holds the type of object to allocate. The returned object reference is then duplicated (`dup` instruction). Next, the constructor of the class `a` is invoked to initialize the returned reference and finally, the returned reference is put into a static field of the class. The `new` bytecode is timing *unpredictable*, because it might invoke the GC, which has an unbounded collection cycle. It might also generate a runtime out of memory exception, thereby violating safety criticality. We replace this `new` bytecode with safe timing predictable bytecodes as shown in Figure 13b.

A single `new` bytecode (consuming 3 bytes in the class file) is replaced with a list of bytecodes lines 9-21 (line numbers are given on the right) consuming 25 bytes in the class file, hence the 22 byte increment of target addresses of the conditional bytecodes. The core idea is very simple; since we have already allocated memory for the object (Figure 12a), we can simply replace the `new` bytecode with a bytecode that loads the allocation pointer ($allocPtr$) onto the stack (line 9) as the object reference. The rest of the bytecodes setup the object header for the allocated object. Our object header

[6] In both these figures, the line numbers are shown on the right of the program. The numbers adjacent to the bytecodes are the starting bytes for each of the bytecodes as obtained by the `javap -c -v` command

is two words (the object structure for the running example is shown in Figure 13c), the first word contains the pointer to the start of data, in the current example it is the address 65485. The first word of the object header is initialized in lines 11-13 in Figure 13b. The second word of the object header is the pointer to the start of the method structure of the class (c.f. Figure 4b). The second word of the object header is initialized in lines 14-21 in Figure 13b.

In case of `newarray` and `anewarray` bytecodes, the array layout (shown in Figure 13d) differs from standard real-time JVM implementations. For array allocation, we again use a 2 word object header. The first word contains the pointer to the start of array data. The second word of the object header contains the size of the array. Other than the smaller object header size compared to standard real-time JVM object header size (usually 8 words), which uses spines [Pizlo et al. 2010] or handles [Schoeberl 2005] for arrays and regular objects, respectively, we also structure our arrays differently. We have decided to place arrays in a row major order and contiguously irrespective of the array dimensions (as in 'C') rather than using linked lists for array traversal in order to allow $O(1)$ array accesses.

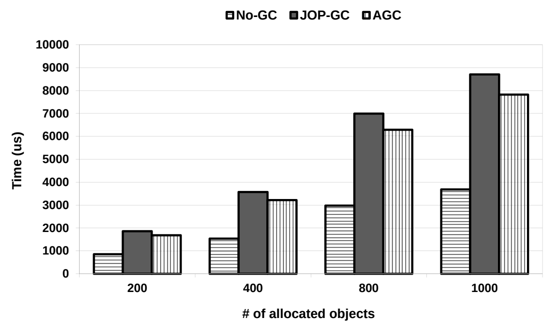
7. EXPERIMENTAL RESULTS

We have carried out two sets of experiments: micro-benchmarks and static WCRT analysis on well established real-time benchmarks to study the efficacy of the proposed memory management scheme compared to two different real-time garbage collector (RTGC) implementations [Schoeberl 2005; Pathirana 2013]. The first is the RTGC proposed in [Schoeberl 2005], which is a concurrent version of Cheney's copy collector developed by Baker [Baker Jr 1978]. In the rest of the section we call it JOP-GC. The second one proposed in [Pathirana 2013] is a variation of the implementation of the allocation and real-time collector described in [Pizlo et al. 2010], in the rest of the section we term it AGC. All our micro-benchmark experiments are run in ModelSim, simulating the real-time hardware implemented JVM called JOP, running at 100 MHz, with 256KB of main memory, 1KB of method cache, and 4KB of stack cache.

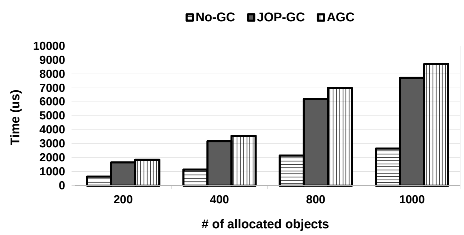
7.1. Micro benchmarks

We carry out three micro-benchmarks: (1) simple object allocation, where we allocate, in a tight loop, 2 objects with 1 word primitive field each. (2) Simple array allocation, where we allocate, in a tight loop, an object with an array type reference field, which is itself initialized to a 20 word 1D array of primitive `int` type. (3) Complex object allocation: in this case, we use a nested 2D loop. In every outer loop iteration, a *reference* type 1D array of size 20 is allocated. In every iteration of the inner loop, 2 objects are allocated and the second object's reference is set as the array value. The example benchmarks are available from [Pathirana 2013].

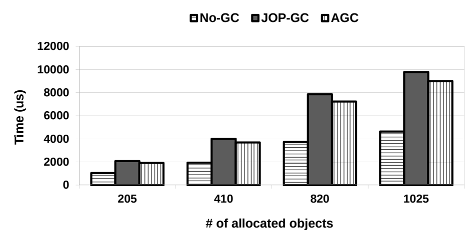
The runtime comparison between the three approaches for each of the micro-benchmarks is shown in Figures 14 and 15. Figures 14 and 15 give the times for completion of a memory allocation request in two cases: (1) when the GC is not invoked, i.e., there is enough memory available to allocate the requested space and (2) when the GC is invoked to collect the garbage when enough memory is not available upon a memory allocation request from the mutator, respectively. The proposed approach is, on average, approximately $3\times$ faster compared to the RTGC based approaches. In many cases (see Figure 15) JOP-GC could not perform garbage collection correctly, as it ran out of handles resulting in null-pointer exceptions. There is a greater speedup in case of array allocations compared to simple object allocations. The runtime memory allocation time difference is significant in the case of GC invocation when compared to the case of no GC invocation. In case of memory allocation time when no GC is invoked, there are multiple reasons for the speedup: (1) there are no read/write barriers when



(a) Simple object memory allocation runtime without GC invocation.



(b) Simple array memory allocation runtime without GC invocation.



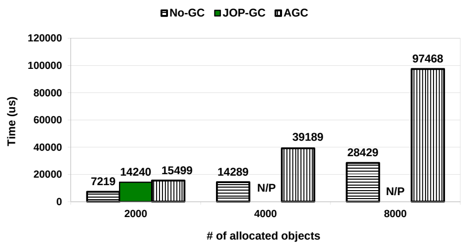
(c) Complex object memory allocation runtime without GC invocation.

Fig. 14: Runtime for micro-benchmarks without GC invocation.

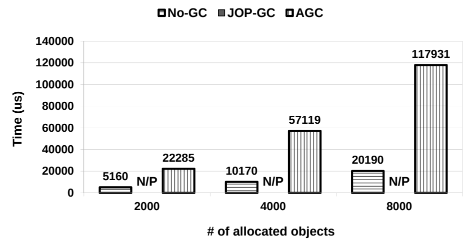
Table I: Comparison of memory words allocated for the complex object allocation. Each word is 4 bytes

# of loop iterations	No-GC (Words)	JOP-GC (Words)	AGC (Words)
100	28	2145	2995
200	28	4290	5590
400	28	8580	11180
500	28	10725	13975
1000	28	21450	27950
2000	28	42900	55900
4000	28	85800	111800

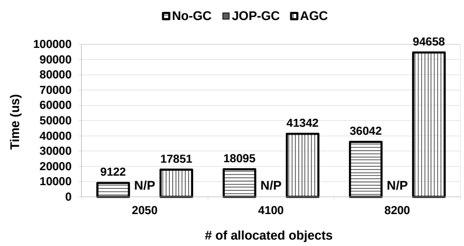
allocating memory, which are needed by both: JOP-GC and AGC. (2) The current `new`, `newarray`, and `anewarray` bytecodes are implemented in software, which means, on *every* execution of these bytecodes, a method may be loaded into the method cache, which results in increased program latency. These runtime numbers are important, because they influence the WCRT as we will see in the next section. (3) In case of the `newarray` and `anewarray` bytecode execution, array spines need to be initialized for AGC [Pizlo et al. 2008; Pathirana 2013], which leads to further slow down in memory allocation times.



(a) Simple object memory allocation runtime with GC. **N/P** stands for `NullPointer` exception



(b) Simple array memory allocation runtime with GC. **N/P** stands for `NullPointer` exception



(c) Complex object memory allocation runtime with GC. **N/P** stands for `NullPointer` exception

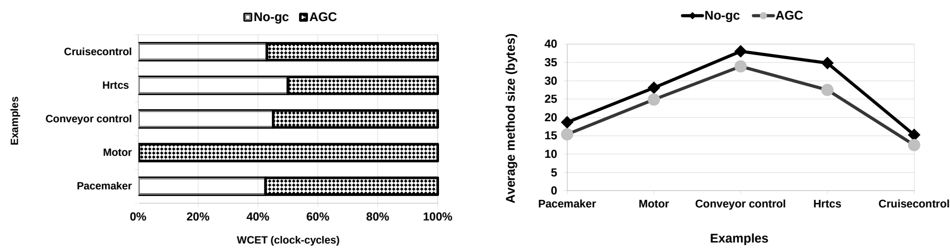
Fig. 15: Runtime for micro-benchmarks with GC invocation

Table I gives the number of words allocated for the complex object allocation benchmark for all the three approaches. In the proposed approach, every iteration of the loop reuses the space allocated on the permanent heap space (same allocation figures would be achieved if the program was recursive rather than iterative). The RTGC approches on the other hand, allocate memory on each `new`, `newarray`, and `anewarray` bytecode execution. Our compile time memory allocation approach simply replaces these memory allocation bytecodes as described in Section 6 and hence, we do not keep on allocating more words. On the other hand, the RTGC approaches, inherently allocate memory on every `new` call. Another important point to note is that AGC although faster when it comes to memory allocation runtime compared to JOP-GC, allocates more words, because AGC uses block based allocation, which leads to internal memory fragmentation. Whereas JOP-GC uses object replication strategy with semi-space partitioning [7].

7.2. WCRT comparisons

In this section we compare the static WCRT obtained by applying the technique described in Section 7.2 on a set of real-time benchmarks. To keep ourselves honest, we chose the benchmarks whose memory allocation requests fit within the allocated heap space without GC invocation. Consequently, we also manually-deleted all code related

[7]In the numbers in Table I, we have not included the semi-space reserved by JOP-GC.



(a) Statically analyzed WCRT values for benchmarks without GC invocation

(b) Statically analyzed method sizes for benchmarks

Fig. 16: WCRT and method size comparison: No-GC vs. AGC

to GC. This primarily included deleting the complete garbage collection method call itself. This was a necessary step, because one cannot *practically* bound the GC cycle time for purpose of static analysis. Puffitsch et al. [2013] have already shown that GC cycle times are orders of magnitudes larger than the execution times of real-time tasks.

We chose 5 real-time synchronous programs for bench-marking. `Cruise control` [Charles 1996] is the cruise speed controller found in cars. `Hrtcs` [Park et al. 2014b] is a human response time gathering system. `Conveyor controller` is a fruit sorter controller, which controls placement of fruits on a conveyor belt using image processing a more detailed description of the application is provided in [Li et al. 2014]. `Motor` [Bourke12 and Sowmya 2009] is a stepper motor control system for a printer. Finally, `Pacemaker` [Park et al. 2014a], is a real-time pacemaker used to control arrhythmia. These benchmarks are compiled to the TP-JOP architecture for execution and their WCRT values are analyzed. The results are shown in Figure 16a. The JOP-GC WCRT results are not shown, because JOP-GC always results in null pointer exceptions for all these examples.

Ignoring the `Motor` example, on average, the proposed No-GC approach's WCRT is around 23% shorter compared to AGC. The `Motor` example is an outlier, with a 4000% improvement in WCRT, because a number of object allocations are performed when emitting events to control the motor coils. In other examples, object allocations are used in computations on the transient heap space, with relatively fewer object allocations on the permanent heap. This requires, deep copying objects from the transient to the permanent heap space, which balances out the WCRT times between the No-GC and the AGC approaches. Note that, the method cache load times and synchronized heap access overheads are still present in the AGC approach.

Finally, Figure 16b gives the average method size comparisons for the compiled SystemJ programs. Since we replace a single object allocation bytecode `new` or `newarray` with a list of other bytecodes, the proposed approach increases the method size. Consequently this also has a penalty on the computed WCRT value, in case of the proposed approach, as the method cache load time increases, but it is not as significant as loading the methods implementing the object allocation for every `new`/`newarray` bytecode execution in the AGC approach.

8. RELATED WORK AND DISCUSSION

Real-time garbage collectors (RTGCs) such as the ones proposed in [Gestegard-Robertz and Henriksson 2003; Kalibera et al. 2009] and the ones presented in [Pizlo et al. 2008], aim to achieve good garbage collection throughput while providing analytic solutions to the worst case execution times of a *single* GC cycle. The major drawback of these approaches is that the equations capturing the GC cycle times depend upon the memory allocation characteristic of the application. Application specific memory allocation patterns can only be gauged via profiling and are *not* exhaustive. In the general case, assuming the worst case memory allocation patterns results in a very pessimistic GC worst case bound as shown in [Puffitsch 2013]. In this paper we are trying to remedy this problem by removing GC altogether.

We are not the first to advocate a GC-less development approach for hard real-time applications with managed languages, especially Java. There exist previous works like the Safety Critical Java (SCJ) [scj 2013] specification and Ravenscar-Java [Kwon et al. 2005] that present a memory organization approach similar to the one presented in this paper. SCJ and Ravenscar-Java both; divide the backing store into multiple parts. A memory region called scoped memory is allocated per real-time task (called a handler in SCJ terminology), which is automatically reclaimed, using pointer reset like us, once the task completes execution. There also exists a permanent memory area similar to the permanent heap advocated in this paper. Any object allocated in the permanent memory area lives throughout the lifetime of the application. One major difference between the presented memory partitioning approach and the memory partitioning in SCJ/Ravenscar-Java is that in the presented approach the same transient heap space (similar to scoped memory in SCJ) can be used by multiple synchronous parallel reactions, whereas every single task gets allocated its own scoped memory in SCJ. This is because, SCJ allows preemption of tasks during program execution, whereas in SystemJ a clock-domain tick is atomic. Furthermore, in this paper we also provide a compile time memory allocation approach, whereas in the SCJ approach, the programmer needs to explicitly and carefully allocate objects to the correct memory areas manually. We believe that our approach helps speed up development times.

The proposed algorithms also have a lot in common with a number of tools proposed for easing development of programs in SCJ. The worst case memory analyzer tool [Andersen et al. 2013] developed for SCJ is able to provide the maximum permanent memory and scoped memory sizes needed for each task automatically, like us. But, the programmer is still responsible for reserving the appropriate sized permanent memory and scoped memory areas in the backing store. Compiler driven memory allocation is a non-trivial task for SCJ applications, because SCJ allows free use of the Java language, consequently bringing all the disadvantages such as: pointer aliasing, shared memory communication, etc, which we restrict in the SystemJ MoC.

Finally, we would like to mention that the memory organization presented in this paper is applicable to JVMs other than JOP. All one needs to provide is an array based backing store, which can be used to implement the memory organization for SystemJ as described in this paper.

9. CONCLUSION AND FUTURE WORK

In this paper we have presented a new memory organization and model for hard real-time and safety-critical applications programmed in formal synchronous languages with data support being provided by managed language runtimes, primarily Java. The presented approach utilizes the formal synchronous programming model provided by SystemJ, to divide the Java objects into two types: (1) transient objects, which are allocated for a single clock-domain transition *only* and are collected by simple pointer reset. (2) Permanent objects, which are alive throughout the life time of the application. Adhering to the synchronous model of computation, allows us to perform compile time memory allocation, which means that our programs are guaranteed to be free of execution time out of memory exceptions. Furthermore, we are able to replace object allocation bytecodes, with real-time analyzable alternatives, which makes our approach amenable to non-pessimistic worst case execution time analysis.

All the aforementioned qualitative improvements are further accompanied by improvements in the program throughput (approximately $3\times$ faster memory allocation times) and worst case execution times, as synchronization barriers necessary in real-time garbage collection approaches become unnecessary in the proposed approach. Not to mention the improvements in the number, size and compacted layout of allocated Java objects and arrays.

However, the proposed approach is not a penance. The proposed memory management technique, does indeed require the programmer to carefully design their real-time applications, as deep object copies need to be made explicitly by the programmer. Furthermore, we see many opportunities for future optimization, especially relaxing the object escapement restrictions currently enforced by the proposed data-flow analysis. We also see an opportunity to *reuse* permanent heap memory allocations across clock-domain tick boundaries, which we plan to address in the future.

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Online Appendix to:
Compiler assisted memory management for safety-critical hard-real time applications

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A. DATA-FLOW ANALYSIS PSEUDO-ALGORITHMS

Figure 17 gives the pseudo-code implementing the data-flow analysis algorithm for compile time memory allocation.

B. THE GRC ANALYSIS ALGORITHM

The algorithm for obtaining *may* and *must* relationships between the GRC nodes is shown in Algorithm 1. The input to this algorithm is an GRC, which is a tuple (V, E) as explained previously. The first part of the algorithm, lines 1-5, groups all the nodes, which are belonging to the same clock-domain. For instance, lines 3-4 initializes a set V_r, which consists of a root node of all clock-domains in GRC. The algorithm then traverses every clock-domain graph, starting from these root nodes via a recursive function `collect_cd_nodes`(v, E) (line 5). A result is a set of set of GRC nodes V_{cd}. Next *May* and *Must* analysis is performed (lines 6-10). Every node in V_{cd} is given as an input to the function `create_node`(v, E) (line 9), which traverses the graph backward starting from the input until it reaches the root node. Function `create_node` first maps the input node to a set of edges E_i using the lambda function (line 14). It is then divided into three parts:

(1) If the current node type is an `JDN` node (line 15) and
  (a) the number of element in E_i is 1, then the only parent node v_i *must* be called before v_j. Then a new node of type $mNode$ is created with its field $Must$ initialized to the result of recursive call of `create_note`.
  (b) the number of element in E_i is greater than 1, then any parent nodes $v_i \in E_i$ *may* be called before v_j. A set of $mList$, consisting return results of `create_node` of each parent node $v_i \in E_i$, is assigned to the field May of a newly created $mNode$.
  (c) the number of element in E_i is 0, then a new $mNode$ is created with both its $Must$ and May fields initialized to $Null$.
(2) If the current node type is `Join`, this means that there are two or more parallel reactions forked at the `Fork` node, which *must* be found in one of the recursive parents of the current node. Hence, the algorithm must continue the backward traversal from this `Fork` node. During compilation, the compiler has already constructed a Hashtable which maps each `Join` node to its corresponding `Fork` node as shown in line 28.
(3) If the current node type is other than `JDN` (line 31) and
  (a) the number of parent node ($|E_i|$) is 0 , then *Null* is returned
  (b) the number of parent node is 1, then a result of `create_node` is returned.

DOI: http://dx.doi.org/10.1145/0000000.0000000

```
Input: program: Program in RTL format
/* pp is the set of program points */
let global pp ← ∅;
let M ← main;
let M ← ReachableMethods (main) ∪ {M};
let size ← sizeof (signal class);
foreach m ∈ M do
    let pp ← (size, get in m, program points
    initializing signals);
end
Step2 (program,pp,main);
```

(a) Step1:Scan signal initialization

```
Input: program: Program in RTL format
Input: pp: Set of program points
Input: method: Method to scan for signal
       emissions
let S ← signalSet (pp);
let C ← code (method);
let PPC ← ∅;
foreach c ∈ C do
    if emits (c,S) then
        PPC ← PPC ∪ {c};
    end
end
Step3 (program,PPC,C);
foreach c ∈ C do
    if invokesMethod (c) then
        Step2 (program,c,getInvoked (c));
    end
end
```

(b) Step2: Scan `setValue` method calls

```
Input: program: Program in RTL format
Input: PPC: Set of program points affected
Input: C: code of the method containing PPC
foreach ppc ∈ PPC do
    if affectsFields (getExpr (ppc)) then
        let pp ← pp ∪ Step4
        (program,ppc,getExpr (ppc),C);
    else if affectsVars (getExpr (ppc)) then
        let pp ← pp ∪ {Step5
        (program,ppc,getExpr (ppc),C)};
    else raise error
end
```

(c) Step3: For every affected field/variable get the object allocation program points

```
Input: program: Program in RTL format
Input: ppc: Affected program point
Input: pexpr: Expression affecting field
Input: C: code of the method containing ppc
let f ← fieldOf (pexr);
let (size,defs) ← Step6 (program,ppc,f,C);
if size ≠ 0 then return (size,defs);
foreach d ∈ defs do
    if !escape (program,C,d) then
        if hasRefFields (f) then
            let PPC ← getAffectedFields
            (getFields (f));
            Step3 (program,PPC,C);
        end
    else raise error;
end
let size ← max (sizeof (defs));
return (size,defs);
```

(d) Step4: Get object allocation program points for affected field

Fig. 17: Data-flow analysis pseudo-algorithm

```
Input: program: Program in RTL format
Input: ppc: Affected program point
Input: pexpr: Expression affecting field
Input: C: code of the method containing ppc

let v ← varOf (pexr);
let defs ← Step7 (program,ppc,v,C);
foreach d ∈ defs do
    if !escape (program,C,d) then
        if hasRefFields (v) then
            let PPC ← getAffectedFields
            (getFields (v));
            Step3 (program,PPC,C);
        end
    else raise error;
end
let size ← max (sizeof (defs));
return (size,defs);
```

(e) Step5: Get object allocation program points for affected variables

```
Input: program: Program in RTL format
Input: ppc: Affected program point
Input: f: Field affected
Input: C: code of the method containing ppc

let ofields ← getAllFields (C) \ {f} ;
let pcs ← reachDef (ppc,f);
if pcs ≠ ∅ then
    let defs ← ∅;
    foreach pc ∈ pcs do
        let vf ← getVarSetting (f,pc);
        let tt ← Step7 (program,pc,vf,C);
        if all other field ∈ ofields have same tt
        then
            /* Concat in set defs                  */
            defs ← defs ∪ tt;
        else /* Append a new set in defs           */
            defs ← defs ∪ {tt};
    end
    return (0,defs);
else
    /* Inter-procedural analysis                   */
    /* Get all callers, including from GRC
    analysis                                       */
    let Callers ← getCallers (C);
    if Callers = ∅ then raise Not_init f;
    let defs ← ∅;
    let sizes ← ∅;
    foreach caller ∈ Callers do
        let defs ← defs ∪ Step6
        (program,getCallerPPC (C),f,caller);
        /* Similar to Figure 17d
        lines 4-12, to fill in the sizes
        set.                                       */
    end
    let size ← max(sizes);
    return (size,defs);
end
```

(f) Step6: Search object allocation program points for affected field

```
Input: program: Program in RTL format
Input: ppc: Affected program point
Input: v: Variable affected
Input: C: code of the method containing ppc

let pcs ← reachDef (ppc,v);
if pcs = ∅ then raise Not_init;
let defs ← ∅;
foreach pc ∈ pcs do
    if getExpr (pc) is New or getExpr (pc) is
    NewArray then
        defs ← defs ∪ {pc};
    else if getExpr (pc) is AffectField then
        let expr ← getExpr (pc);
        let (_,defs) ← Step4
        (program,pc,expr,C);
    else if getExpr (pc) is AffectVar then
        let v ← varOf (getExpr (pc));
        let (_,defs) ← Step7 (program,pc,v,C);
    else
        /* Cannot handle new outside
        current method allocation          */
        raise Not_handled
    end
end
/* Make sure that definitions are not
reachable from each other                  */
if |defs| > 1 then
    assert(notReachable (defs));
end
return (defs);
```

(g) Step7: Search object allocation point(s) for method local affected variable

Fig. 17: Data-flow analysis pseudo-algorithm

Input: $GRC : (V, E)$
Output: $mList$: a set of set of $mNode$
Data: V_{cd}: a set of set of clock-domain nodes
Data: $mNode$ = [$Name$:string,$Must$:$mNode$,May:set of $mNode$]
`/* Grouping nodes for each clock-domain */`
let $V_r = \emptyset \cup V$
let $V_{cd} = \emptyset$
foreach $v \in V$ **do**
> **foreach** $(v_i, v_j) \in E$ **do if** $v_j = v$ **then** $V_r = V_r \backslash \{v\}$

foreach $v \in V_r$ **do** $V_{cd} = V_{cd} \cup$ {`collect_cd_nodes`(v, E)}
`/* Perform May and Must analysis */`
forall the $V \in V_{cd}$ **do**
> **let** $V_{temp} = \emptyset$
> **foreach** $v \in V$ **do**
>> $V_{temp} = V_{temp} \cup$ {`create_node`(v, E)}
>
> $mList = mList \cup \{V_{temp}\}$

Function `create_node`(v, E):
begin
> **let** $E_i = \lambda(v)$
> **if** $v = ActionNode$ **then**
>> **if** $|E_i| = 1$ **then**
>>> **let** $\{(v_i, v_j)\} = E_i$
>>> **let** $newNode = [Name \leftarrow get\_name(v_j); Must \leftarrow$ `create_node`$(v_i, E); May \leftarrow Null]$
>>> **return** *newNode*
>>
>> **else if** $|E_i| > 1$ **then**
>>> **let** $V_{temp} = \emptyset$
>>> **foreach** $(v_i, v_j) \in E_i$ **do**
>>>> $V_{temp} = V_{temp} \cup$ {`create_node`(v_i, E)}
>>>
>>> **return** $[Name \leftarrow get\_name(v_j); Must \leftarrow Null; May \leftarrow V_{temp}]$
>>
>> **else if** $|E_i| = 0$ **then**
>>> **return** $[Name \leftarrow get\_name(v_j); Must \leftarrow Null; May \leftarrow Null]$
>
> **else if** $v = JoinNode$ **then**
>> **let** v = `find_matching_fork`(v)
>> **let** $\{(v_i, v_j)\} = \lambda(v)$
>> **return** `create_node`(v_i, E)
>
> **else**
>> **if** $|E_i| = 0$ **then**
>>> **return** *Null*
>>
>> **else if** $|E_i| = 1$ **then**
>>> **return** `create_node`(v_i, E)

Function `collect_cd_nodes`(v, E):
begin
> **let** $V_{temp} = \emptyset$
> **foreach** $(v_i, v_j) \in E$ **do**
>> **if** $v = v_i$ **then**
>>> $V_{temp} = V_{temp} \cup \{v_j\}$
>>> $V_{temp} = V_{temp} \cup$ `collect_cd_nodes`(v_j, E)
>
> **return** V_{temp}

ALGORITHM 1: May and Must analysis