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Non-termination of Dalvik bytecode *via* compilation to CLP

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Abstract

We present a set of rules for compiling a Dalvik bytecode program into a logic program with array constraints. Non-termination of the resulting program entails that of the original one, hence the techniques we have presented before for proving non-termination of constraint logic programs can be used for proving non-termination of Dalvik programs.

Keywords and phrases Non-Termination, Android, Dalvik, Constraint Logic Programming

1 Introduction

Android is currently the most widespread operating system for mobile devices. Applications running on this system can be downloaded from anywhere, hence reliability is a major concern for its users. In this paper, we consider applications that may run into an infinite loop, which may cause a resource exhaustion, for instance the battery if the loop continuously uses a sensor as the GPS. Android programs are written in Java and compiled to the Google's Dalvik Virtual Machine (DVM) bytecode format [3] before installation on a device. We provide a set of rules for compiling a Dalvik bytecode program into a constraint logic program [5]. Non-termination of the resulting program entails that of the original one, hence the technique we have presented before [6] for proving non-termination of constraint logic programs can be used for proving non-termination of Dalvik programs. We model the memory and the objects it contains with *arrays*, so we compile Dalvik programs to logic programs with array constraints and we consider the theory of arrays presented in [1].

2 The Dalvik Virtual Machine

We briefly describe the operational semantics of the DVM (see [3] for a complete description). Unlike the JVM which is stack-based, the DVM is register-based. Each method uses its own array of registers and invoked methods do not affect the registers of invoking methods. The number of registers used by a method is statically known. At the beginning of an execution, the N arguments to a method land in its last N registers and the other registers are initialized to 0. Many Dalvik bytecode instructions are similar, so we concentrate on a restricted set which exemplifies the operations that the DVM performs.

- *const* d, c Move constant c into register d (*i.e.*, the register at index d in the array of registers of the method where this instruction occurs).
- *move* d, s Move the content of register s into register d .
- *add* d, s, c Store the sum of the content of register s and constant c into register d .

- *if-lt* i, j, q If the content of register i is less than the content of register j then jump to program point q , otherwise execute the immediately following instruction.
- *goto* q Jump to program point q .
- *invoke* S, m where $S = s_0, s_1, \dots, s_p$ is a sequence of register indexes and m is a method. The content r^{s_0} of register s_0, \dots, r^{s_p} of register s_p are the *actual parameters* of the call. Value r^{s_0} is called *receiver* of the call and must be 0 (the equivalent of `null` in Java) or a reference to an object o . In the former case, the computation stops with an exception. Otherwise, a lookup procedure is started from the class of o upwards along the superclass chain, looking for a method with the same signature as m . That method is run from a state where its last registers are bound to $r^{s_0}, r^{s_1}, \dots, r^{s_p}$.
- *return* Return from a void method.
- *new-instance* d, κ Move a reference to a new object of class κ into register d .
- *iget* d, i, f (resp. *iput* s, i, f) The content r^i of register i must be 0 or a reference to an object o . If r^i is 0, the computation stops with an exception. Otherwise, $o(f)$ (the value of field f of o) is stored into register d (resp. the content of register s is stored into $o(f)$).

3 Compilation to CLP clauses

We model a *memory* as a pair (a, i) where a is an array of *objects* and i is the index into this array where the next insertion will take place. An *object* o is an array of terms of the form $[w, f_1(v_1), \dots, f_n(v_n)]$ where w is the name of the class of o , f_1, \dots, f_n are the names of the fields defined in this class and v_1, \dots, v_n are the current values of these fields in o . So, the first component of a memory is an array of arrays of terms and a memory location is an index into this array. Memory locations start at 1 and 0 corresponds to the `null` value.

Our compilation rules are given in Fig. 1–3. We associate a predicate symbol p_q to each program point q of the Dalvik program P under consideration. We generate clauses with constraints on integer and array terms. Our constraint theory combines the theory of integers with that of arrays defined in [1]. Our CLP domain of computation \mathcal{D} (values interpreting constraints) is the union of \mathbb{Z} with the set Obj of arrays of terms of the form $f(i)$ where i is an integer and with the set of arrays of elements of Obj . The read $a[i]$ returns the value stored at position i of the array a and the write $a\{i \leftarrow e\}$ is a modified so that position i has value e . For multidimensional arrays, we abbreviate $a[i] \dots [j]$ with $a[i, \dots, j]$.

Each rule considers an instruction *ins* occurring at a program point q . We let $\tilde{V} = V_0, \dots, V_{r-1}$ and $\tilde{V}' = V'_0, \dots, V'_{r-1}$ be sequences of distinct variables where r is the number of registers used by the method where *ins* occurs. For each $i \in [0, r-1]$, variable V_i (resp. V'_i) models the content of register i before (resp. after) executing *ins*. We let M denote the input memory and M' the output memory. So, \tilde{V} and M (or $[A, I]$) in the head of the clauses are input parameters while M' is an output parameter. We let id denote the sequence $(V'_0 = V_0, \dots, V'_{r-1} = V_{r-1})$ and id_{-i} (where $i \in [0, r-1]$) the sequence $(V'_0 = V_0, \dots, V'_{i-1} = V_{i-1}, V'_{i+1} = V_{i+1}, \dots, V'_{r-1} = V_{r-1})$. By $|\tilde{X}|$ we mean the length of sequence \tilde{X} . For any method m , q_m is the program point where m starts, $reg(m)$ is the number of registers used by m and $sign(m)$ is the set of all the methods with the same signature as m .

Some compilation rules are rather straightforward. For instance, *const* d, c moves constant c into register d , so in Fig. 1 the output register variable V'_d is set to c while the other register variables remain unchanged (modelled with id_{-d}). Rules for *move*, *add* and *goto* are similar. In Fig. 2, we consider method calls. The instruction *invoke* s_0, \dots, s_p, m is compiled into a set of clauses (one for each method with the same signature as m) which

$$\frac{\text{const } d, c}{p_q(\tilde{V}, M, M') \leftarrow \{V'_d = c\} \cup id_{-d}, p_{q+1}(\tilde{V}', M, M')} \quad (1a)$$

$$\frac{\text{if-lt } i, j, q'}{\left\{ \begin{array}{l} p_q(\tilde{V}, M, M') \leftarrow \{V_i < V_j\} \cup id, p_{q'}(\tilde{V}', M, M'), \\ p_q(\tilde{V}, M, M') \leftarrow \{V_i \geq V_j\} \cup id, p_{q+1}(\tilde{V}', M, M') \end{array} \right\}} \quad (1b)$$

Figure 1 Compilation of some simple Dalvik instructions.

$$\frac{\text{invoke } s_0, \dots, s_p, m}{\left\{ \begin{array}{l} p_q(\tilde{V}, M, M') \leftarrow \{V_{s_0} > 0\} \cup id, \\ \text{lookup}_P(M, V_{s_0}, m, q_{m'}), \\ p_{q_{m'}}(\tilde{X}_{m'}, M, M_1), \\ p_{q+1}(\tilde{V}', M_1, M') \end{array} \right\} \left| \begin{array}{l} m' \in \text{sign}(m) \\ \text{and } \tilde{X}_{m'} = 0, \dots, 0, V_{s_0}, \dots, V_{s_p} \\ \text{with } |\tilde{X}_{m'}| = \text{reg}(m') \end{array} \right.} \quad (2a)$$

$$\frac{\text{return}}{p_q(\tilde{V}, M, M') \leftarrow \{M' = M\}} \quad (2b)$$

Figure 2 Compilation of some Dalvik instructions related to method calls.

impose that V_{s_0} (the receiver of the call) is a non-null location (*i.e.*, $V_{s_0} > 0$). Therefore, if $V_{s_0} \leq 0$, the execution of the generated CLP program fails, as the original Dalvik program. If $V_{s_0} > 0$, the lookup procedure begins. For each $m' \in \text{sign}(m)$, this is modelled with the call $\text{lookup}_P(M, V_{s_0}, m, q_{m'})$ which starts from the class of the object at location V_{s_0} in memory M and searches for the closest method m'' with the same signature as m upwards along the superclass chain. If $m'' = m'$, this call succeeds, otherwise it fails. Then, m' is executed, modelled with $p_{q_{m'}}(\tilde{X}_{m'}, M, M_1)$, with some registers $\tilde{X}_{m'}$ initialized as expected. When the execution of m' has finished, control jumps to the following instruction (*i.e.*, $p_{q+1}(\tilde{V}', M_1, M')$). In Fig. 3, we consider some memory-related instructions that we compile to clauses with array constraints.

► **Theorem 1.** *Let P be a Dalvik bytecode program and P_{CLP} its CLP compilation. If there is a computation $p_{q_0}p_{q_1} \dots$ in P_{CLP} then there is an execution $q_0q_1 \dots$ of P .*

More precisely, if there is a finite (resp. infinite) computation in P_{CLP} starting from a query $p_{q_0}(\tilde{v}, [a, i], M')$ (where \tilde{v} , a and i are values in \mathcal{D} and M' is an output variable), then there is a finite (resp. infinite) execution of P , using the same program points, starting from values corresponding to \tilde{v} and a in the DVM registers and memory.

4 Non-termination inference

The following proposition is a CLP reformulation of a result presented in [4].

► **Proposition 2.** *Let $r = p(\tilde{x}) \leftarrow c, p(\tilde{y})$ and $r' = p'(\tilde{x}') \leftarrow c', p(\tilde{y}')$ be some clauses. Suppose there exists a set \mathcal{G} such that formulæ $[\forall \tilde{x} \exists \tilde{y} \tilde{x} \in \mathcal{G} \Rightarrow (c \wedge \tilde{y} \in \mathcal{G})]$ and $[\exists \tilde{x}' \exists \tilde{y}' c' \wedge \tilde{y}' \in \mathcal{G}]$ are true. Then, p' has an infinite computation in $\{r, r'\}$.*

$$\frac{\text{new-instance } d, \kappa}{\begin{array}{l} w \text{ is the name of class } \kappa \text{ and } f_1, \dots, f_n \text{ are the names of the fields defined in } \kappa \\ p_q(\tilde{V}, [A, I], M') \quad \{O[0] = w, O[1] = f_1(0), \dots, O[n] = f_n(0), \\ A_1 = A\{I \quad O\}, V'_d = I, I_1 = I + 1\} \cup id_{-d}, p_{q+1}(\tilde{V}', [A_1, I_1], M') \end{array}} \quad (3a)$$

$$\frac{\text{iget } d, i, f}{p_q(\tilde{V}, [A, I], M') \quad \{V_i > 0, A[V_i, F] = f(V'_d)\} \cup id_{-d}, p_{q+1}(\tilde{V}', [A, I], M')} \quad (3b)$$

$$\frac{\text{iput } s, i, f}{\begin{array}{l} p_q(\tilde{V}, [A, I], M') \quad \{V_i > 0, O = A[V_i], O[F] = f(X), O_1 = O\{F \leftarrow f(V_s)\}, \\ A_1 = A\{V_i \leftarrow O_1\}\} \cup id, p_{q+1}(\tilde{V}', [A_1, I], M') \end{array}} \quad (3c)$$

Figure 3 Compilation of some memory-related instructions.

Consider the Android program in Fig. 4, with the Java syntax on the left and the corresponding Dalvik bytecode P on the right, where v_0, v_1, \dots denote registers $0, 1, \dots$. Method `loop` in class `MyActivity` is called when the user taps a button displayed by the application. Execution of this method does not terminate because in the call to `m`, the objects `o1` and `o2` are aliased and therefore by decrementing `x.i` we are also decrementing `this.i` in the loop of method `m`. We get the following clauses for program points 0 and 14:

$$\begin{array}{l} p_0(\tilde{V}, [A, I], M') \leftarrow \{A[V_1, F] = i(V'_0)\} \cup id_{-0}, p_1(\tilde{V}', [A, I], M') \\ p_{14}(\tilde{V}, M, M') \leftarrow \{V_0 > 0\} \cup id, \text{lookup}_P(M, V_0, \text{Loops} \rightarrow \mathbf{m}(\text{ILoops})V, 0), \\ p_0(0, V_0, V_2, V_1, M, M_1), p_{15}(\tilde{V}', M_1, M') \end{array}$$

Let P_{CLP} denote the CLP program resulting from the compilation of P . The set of *binary unfoldings* [2] of P_{CLP} contains the following clauses

$$\begin{array}{l} r : p_0(\tilde{V}, [A, I], M') \quad \{V_1 > 0, O = A[V_1], O[F] = i(X), X < V_2, \\ O_1 = O\{F \leftarrow i(X + 1)\}, A_1 = A\{V_1 \leftarrow O_1\}, \\ V_3 > 0, O' = A_1[V_3], O'[F'] = i(X'), V'_0 = X' - 1, \\ O'_1 = O'\{F' \leftarrow i(V'_0)\}, A_2 = A_1\{V_3 \leftarrow O'_1\}\} \cup id_{-0}, p_0(\tilde{V}', [A_2, I], M') \\ r' : p_{10}(\tilde{V}, [A, I], M') \leftarrow \{O[0] = \text{loops}, O[1] = i(0), A_1 = A\{I \leftarrow O\}, \\ I_1 = I + 1, I > 0\}, p_0(0, I, 2, I, [A_1, I_1], M_1) \end{array}$$

where r corresponds to the path $0 \rightarrow 1 \rightarrow 3 \rightarrow 4 \rightarrow \dots \rightarrow 9 \rightarrow 0$ and r' to the path $10 \rightarrow 11 \rightarrow 12 \rightarrow 13 \rightarrow 14 \rightarrow 0$ in P . In r' , O corresponds to both o_1 and o_2 , which expresses that o_1 and o_2 are aliased. Note that I , the address of O , is passed to p_0 both as second and fourth parameter, which corresponds in r to V_1 (`this` in method `m`) and V_3 (`x` in `m`). Moreover, when $V_1 = V_3$ in r , we have $O' = O_1$, $F' = F$ and $X' = X + 1$, hence $V'_0 = X' - 1 = X$. Therefore, we have $O'_1 = O$, so $A_2 = A$. The logical formulæ of Proposition 2 are true for the set $\mathcal{G} = \{(\tilde{v}, mem, mem') \in \mathcal{D}^3 | v_1 = v_3\}$. Hence, p_{10} has an infinite computation in $\{r, r'\}$, which implies [2] that p_{10} has an infinite computation in P_{CLP} . So by Theorem 1, P has an infinite execution from program point 10.

```

public class Loops {
    int i;
    public void m(int n, Loops x) {
        while (this.i < n) {
            this.i++;
            x.i--;
        }
    }
}

.method public m(ILoops)V
    .registers 4
0: iget v0, v1, Loops->i:I
1: if-lt v0, v2, 3
2: return-void
3: iget v0, v1, Loops->i:I
4: add-int/lit8 v0, v0, 0x1
5: iput v0, v1, Loops->i:I
6: iget v0, v3, Loops->i:I
7: add-int/lit8 v0, v0, -0x1
8: iput v0, v3, Loops->i:I
9: goto 0
.end method

public class MyActivity extends Activity {
    ...
    public void loop(View v) {
        Loops o1 = new Loops();
        Loops o2 = o1;
        o1.m(2, o2);
    }
    ...
}

.method public loop(Landroid/view/View;)V
    .registers 5
10: new-instance v0, Loops
11: invoke-direct {v0}, Loops-><init>()V
12: move-object v1, v0
13: const/16 v2, 0x2
14: invoke-virtual {v0, v2, v1}, Loops->m(ILoops)V
15: return-void
.end method

```

Figure 4 The non-terminating method loop is called when the user taps a button.

5 Future Work

We plan to implement the technique described above and to write a solver for array constraints. Currently, our compilation rules only consider the operational semantics of Dalvik, a part of the Android platform. We also plan to extend them by considering the operational semantics of other components of Android, for instance *activities* that we have studied in [7].

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