Parallel efficient algorithms and their programming.

Fundation of ATHAPASCAN-1.

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Ce rapport contient les deux premiers chapitres du tutoriel “Parallel Computer Algebra” donné au colloque annuel de calcul formel ISSAC (Juillet 97, Hawaii) par Jean-Louis Roch et Gilles Villard. Il décrit les bases pour la construction, l'analyse et la programmation d'algorithmes parallèles sur des architectures distribuées. Utilisant les techniques à la base de l'algorithmique parallèle synchrone PRAM, il montre comment elles peuvent être appliquées à la construction d'algorithmes parallèles qui conduisent à des programmes performants sur des architectures distribuées asynchrones. Deux points sont alors critiques: la prise en compte des surcoûts de communication et d'ordonnancement.

Le premier chapitre présente les techniques permettant la construction d'algorithmes *efficaces*. Différents critères doivent être minimisés: le nombre d'opérations qui doit rester proche du nombre optimal d'opérations sur une machine séquentielle, le temps parallèle minimal (i.e. sur un nombre infini de processeurs) pour permettre que le temps d'exécution diminue lorsque le nombre de processeurs augmente (on parle d'extensibilité), le volume de communications pour limiter leur surcoût sur une architecture distribuée. Pris séparément, ces différents critéres conduisent à des algorithmes fondamentalement différents. Un algorithme efficace doit alors réaliser un bon compromis entre ces critères, pris deux à deux ou dans leur ensemble; il est souvent basé sur un couplage entre ces différents algorithmes: on parle de *poly-algorithmes* ou d'*algorithmes en cascade*.
Différents exemples illustrent les techniques de base pour construire des algorithmes efficaces réalisant des compromis intéressants. Les algorithmes sont représentés par des graphes de flots de données. Leur programmation est explicitée à partir d'un langage abstrait, ATH (Asynchronous Tasks Handling).

Le deuxième chapitre étudie l'ordonnancement de tels algorithmes sur une architecture distribuée asynchrone (modèle LogP). Le cas le plus général où le graphe est inconnu (les tâches qui le constituent sont construites en cours d'exécution et sont de durées inconnues) est spécifiquement étudié. Un algorithme d'ordonnancement en-ligne qui assure des exécutions optimales pour un algorithme parallèle efficace tel ceux étudiés dans le chapitre 1 est explicité. En conclusion, le langage ATHAPASCAN qui permet l'implémentation d'un tel ordonnancement est présenté. Ce langage (implémenté par une bibliothèque C++) est une réalisation concrète du langage ATH.

Chapter 1

Parallel efficient algorithms

Contents

1.1 PRAM, DFG and cost analysis 4
1.1.1 The PRAM model 4
1.1.2 Execution of a PRAM program and data-flow graphs 7
1.1.3 Describing PRAM algorithms: ATH language 8
1.1.4 Time, work and communication costs 9
1.1.5 Efficient algorithms 11
1.1.6 Example 12
1.1.7 Relations between PRAMs 13
1.2 Increasing granularity 13
1.2.1 Parallel divide and conquer 14
1.2.2 Minimizing communication work 14
1.2.3 Conclusion 16
1.3 Redundancy and cascading divide&conquer 18
1.3.1 DFG of the best sequential algorithm 18
1.3.2 Breaking dependencies 19
1.3.3 Cascading divide&conquer to minimize time 20
1.3.4 Applications in linear algebra 22
1.3.5 Conclusion 22
1.4 Randomization to decrease time or preserve work. 23
1.4.1 Randomization to suppress dependencies 23
1.4.2 Randomization to provide efficiency 25
1.4.3 Conclusion 26
1.5 Parallel time complexity and NC Classification 26
1.6 Conclusion 27

Parallel algorithmic is a successful theory. Several methods, techniques and paradigms, which are presented in several books and surveys [60, 5, 30, 38, 35, 20, 41, 28, 39, 45] have been developed to build powerful theoretical algorithms. Furthermore, they stand as a basis for implementation of performant programs on effective parallel architectures. Those general techniques overflow computer algebra framework even if arithmetic and algebraic computations are of specific interest.

In this chapter, we introduce the main techniques involved in the building of parallel algorithms. They are illustrated on elementary computer algebra problems. The underlying model is PRAM but the data-flow graph representation is also introduced. It is used to describe executions of a parallel algorithm and to define its cost. Three factors are here preponderant: parallel execution time, number of operations and granularity which is related to the required volume of communications. An efficient algorithm realizes a compromise solution between those three factors.

The organization of the chapter is as follows. Section 1 describes the local PRAM model, the data-flow graph representation and cost analysis. Following sections illustrate, using simple examples, the main techniques involved in the building of:

- section 2: a coarse granularity algorithm from a fine grain optimal one;
- section 3: a fast optimal algorithm from a very fast but non optimal one;
- section 4: a very fast optimal randomized algorithm from a deterministic but non optimal one.

Finally, in the last section, we give an overview of parallel time complexity, focusing on boolean-arithmetic circuits which are commonly used in computer algebra.

1.1 PRAM, DFG and cost analysis

The Parallel Random Access Machine (PRAM) [18, 4] is the most common execution model used to build and analyze parallel algorithms. Its major feature is to be independent from the number of processors used. In this section we focus on the local PRAM model introduced in [38]. Cost analysis takes into account both arithmetic and communication complexities.

In the following, A denotes an algorithm and A_n its restriction for input of size $O(n)$.

1.1.1 The PRAM model

A Local Parallel Random Access Machine (PRAM) is set of:

- an (infinite) number of processors $P_0, \ldots, P_N, \ldots$, each indexed by an integer (*processor identifier* or `pid` in short). Each processor is a RAM (Random Access Machine [2]) and gets its own local memory which contains its own pid.
- a global (or shared) memory. Each processor can copy data from the global memory into its own local memory: this operation is called `global read` or `read` in short. Conversely, each processor can copy a data from its own local memory into the global one: this operation is a `write` operation.
  Initially, the input data are available in global memory. At the end of the computation, the output data are also stored there.

- A program that consists in a finite sequence of RAM elementary instructions, extended by the global elementary (i.e. single word location) read and write instructions.

- a global clock that ensures a synchronous mode of computation. After initialization (first top), processors are ready to execute the first instruction of the program. At each top (or *step*), each processor executes the next RAM instruction in the program. Thus it performs either an elementary arithmetic operation within its local memory or an access to the shared memory (read or write).
  The program terminates when processor with pid 0 executes the `halt` instruction.

Note that the program may contain branching instructions eventually depending on the pid value. Due to branching instructions, at a given top, processors may execute different instructions (Multiple Instruction Multiple Data – MIMD – type).

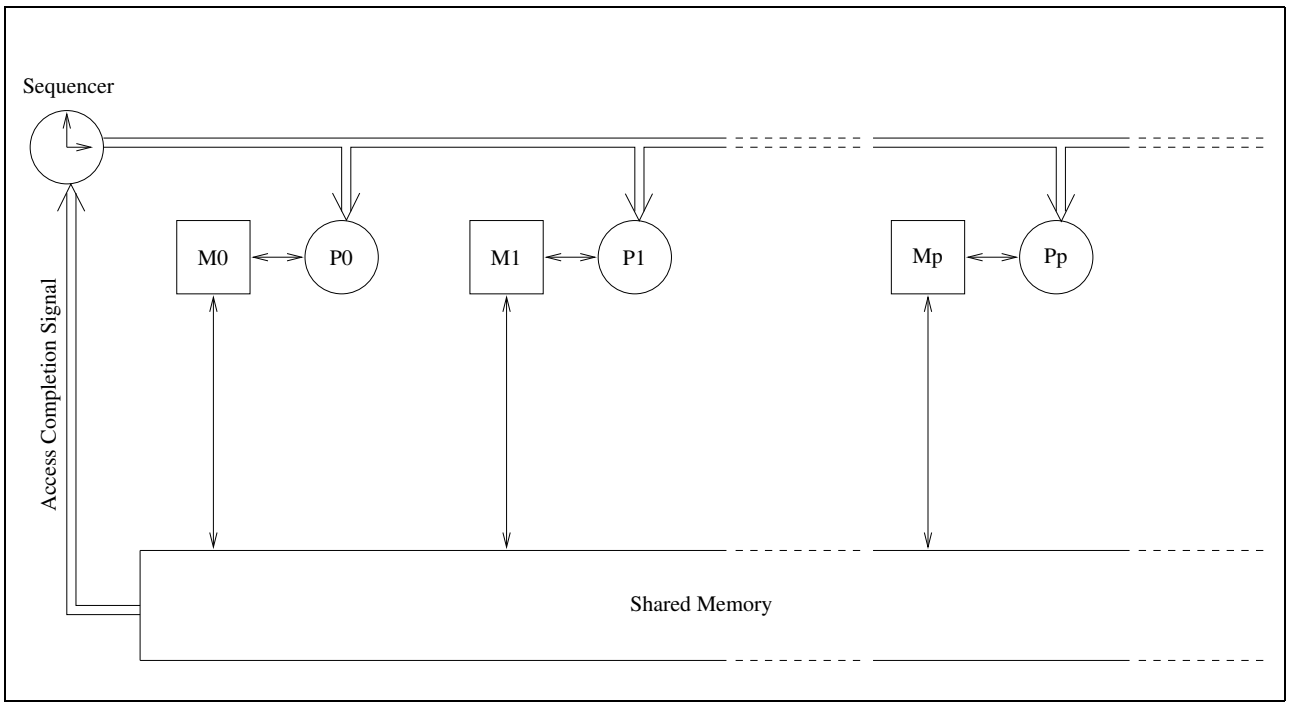

Figure 1.1: The local PRAM execution model

Semantics of access in shared memory. Due to the synchronous mode of computation, semantics of global memory access is simple and only depends on the behavior when, at a same top, several processors concurrently accede to a same single location in the shared memory.

At a same top, two processors can't perform both a read and a write in the same location. But concurrent read (or concurrent write) access may be allowed, depending on the PRAM:

- an EREW-PRAM (Exclusive Read Exclusive Write) does not allow concurrent access to a single location.

- a CREW-PRAM (Concurrent Read Exclusive Write) allows only concurrent read access.
- a CRCW-PRAM (Concurrent Read Concurrent Write) allows concurrent access (all in the same mode, either read or write).

When a concurrent write operation is performed into a single location in the shared memory, different semantics are considered depending on the reduction operation performed to produce the final value:

- COMMON: all processors have to write the same value. If not, an error is produced.
- ARBITRARY: an arbitrary processor writes its value.
- PRIORITY: the processor with the minimum pid writes its value.
- CUMULATIVE: the sum of all the concurrent values is written. The addition operation (defined between single location values) is assumed to be associative. Furthermore, it is assumed to be commutative; this ensures, a semantic independent from the pids of the writing processors likewise concurrent read and common or arbitrary write operations. This concurrent write mode is also called *combining* [41].

As detailed further, those different variants of the PRAM are relatively closed to each others: each one can simulate the other one with small overheads [14, 41, 28].

Dynamic task creation The above definition presents two drawbacks:

- it assumed that, after initialization, an unbounded number of processors start execution;
- dynamic creation of parallelism has to be described in the program using busy-waiting; this means that the scheduling of the program is completely described in the program.

In order to address this second point, in the initial definition from [18], only the processor with pid 0 starts execution of the program. To generate parallelism, an elementary `fork <e>` instruction is defined. When a processor P executes this instruction, an inactive processor P' is reset. The accumulator of P (which may contain an address in the shared memory where some parameters are stored) is first copied into the one of P'. The pid of P' is then put into the accumulator of P. This allows P and P' to later communicate via the shared memory.
At the next step, P executes the following instruction (the one that follows the `fork`) and P' starts the execution of the program at the instruction labeled `e`.
Using `fork`, dynamic task creation is made possible, scheduling (allocation of inactive processors) being ensured by the PRAM machine. However, this modification implies that any PRAM program that uses a polynomial number $n^{O(1)}$ of processors takes a time $\Omega(\log n)$ to be executed, forbidding the building of constant time algorithms; if an algorithm is involved during the execution of a program (e.g. inside the body of a loop), this overhead may easily be avoided. Analysis of costs in this chapter are made under the previous model, thus without taking into account task allocation overhead.

Randomized PRAM To support execution of randomized algorithms, the PRAM is extended in the following way. A new `random` instruction is introduced that allows each processor to generate (in one top) a random bit (or a random number that fits in a single memory location).

Random generations (i.e. `random` instructions) performed by a processor during the execution are assumed to be independent realizations of an uniform law. Moreover, generations performed in parallel at a given top by different processors are also assumed to be independent.

1.1.2 Execution of a PRAM program and data-flow graphs

Being given the input data, the execution of a PRAM program may be represented as a direct acyclic graph. Vertices correspond to instructions that are executed (one vertex, one instruction) and edges to precedence relations between instructions.

Basically, due to the synchronicity of the PRAM, if v (resp. w) is the vertex representing an instruction executed[1] at step i (resp. $i+1$), then there is an edge from v to w. If we forget extra synchronization due to the machine model, synchronizations required by the algorithm itself to ensure correctness of the execution correspond to the ordering of access into a location in memory. This ordering can be represented by the (macro) data-flow graph (DFG) related to the execution. DFG is direct acyclic and bipartite with node sets $J = \{j_1, \ldots, j_n\}$ corresponding to instructions (j meaning *job*) and $T = \{t_1, \ldots, t_m\}$ corresponding to single assignment data (t meaning transition). An edge goes from t_k (resp. j_i) to j_i (resp. t_k) if j_i is a read (resp. write) instruction of the global data related to t_k.

In this DFG, any memory access, either global or local, is represented by an edge between a location (represented by a transition node) and an instruction (a job node) that requires the access. Except for transitions related to input, immediate ancestors of each transition t_k are write instructions: only one on an exclusive-write PRAM, eventually more on a concurrent-write one. Conversely, its immediate successors (except for transitions related to output) are read instructions: only one on an exclusive PRAM, eventually more on a concurrent-read one. This means that when all immediate successors (job nodes) of a transition have been executed, the location related to it in global memory may be garbaged.

Let us consider the DFG related to a tree computation scheme. As an illustration, we consider two algorithms that solve the *iterated product*[2] problem: it consists in computing the product of n elements. In order to exhibit parallelism, multiplication is assumed to be associative and commutative. A balanced binary tree scheme gives an algorithm that works on an EREW PRAM; related DFG is shown in figure 1.2.a. On a CUMULATIVE-ERCW PRAM all products may be performed concurrently and cumulated on a shared location (fig. 1.2.b).

This graph defines a precedence relation, denoted $\prec$, between instruction nodes in J. Let j_1, j_2 be two nodes in J; $j_1 \prec j_2$ if there is a path in DFG from j_1 to j_2. In the following, we will consider the subgraph $DFG_a(J, \prec)$ of DFG, where only arithmetic instructions and their precedence relations are represented.

Remark 1. The data-flow description of the algorithm is roughly equivalent to a *straight-line program* [32].

[1] Instructions corresponding to v and w may be executed by different processors.

[2] also called *iterated sum* when an addition law is considered

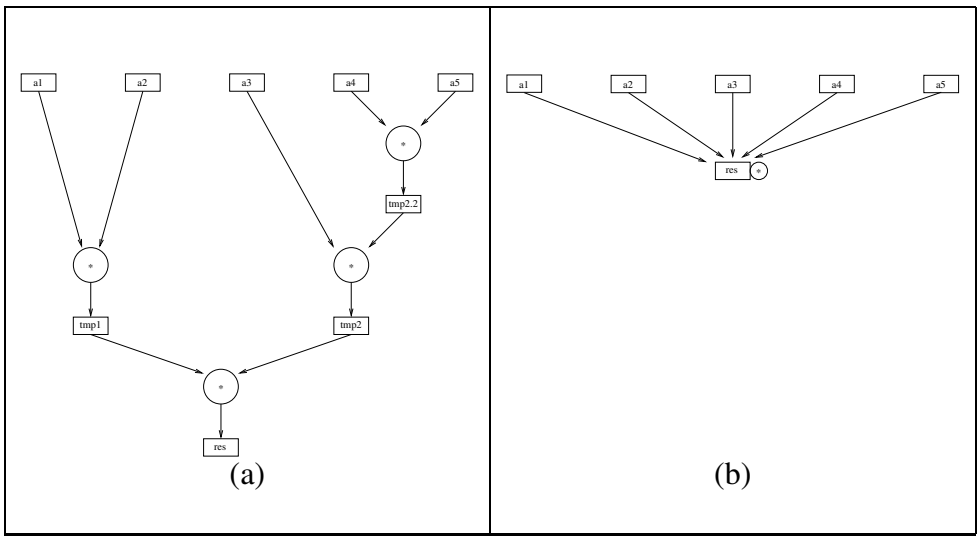

Figure 1.2: DFG of two iterated products: (a) EREW (b) cumulative-ERCW

Remark 2. Note that symmetry of input (resp. output) edges to a transition node assumes commutativity of access. This is verified for any concurrent write (resp. read) access defined on the PRAM.

1.1.3 Describing PRAM algorithms: ATH language

PRAM stands as an abstract model virtualizing any parallel architecture. In order to describe PRAM algorithms, we need an elementary programming language which leads to an easy description of algorithms.

Since the evaluation of a parallel algorithm is directly related to the analysis of DFG, a sequential description should be sufficient thanks to the implicit appearance of data-dependencies: each read access to a location gets the value put by the last write in a sequential execution. However, two characteristics, which do not appear in a sequential description, are to be taken into account:

- two levels of memory access are distinguished: local and global. Global memory access support CUMULATIVE-CRCW semantics.
- the elementary unit of instruction is the block. A block is a sequence of elementary RAM instructions. A block is executed in sequential; it takes benefit of local access.

In the following, we consider an extension of the basic PRAM basic language introduced in [18] based on those two considerations. This abstract language is called *ATH*, an acronym for *Asynchronous Tasks Handling*.

Blocks of instructions are defined as procedures bodies. The execution of such a block is called a *task*. Tasks may be ordered either in sequence using synchronous procedure call or in parallel using asynchronous procedure calls (prefixed by `fork`). In this last case, precedence relations between tasks are defined in a natural way, according to shared-data dependencies that appear

in a sequential execution of the program. Data dependencies concerning local data are then not considered in the relative DFG.

Figure 1.3 gives two different recursive programs for the iterated product using a C++-like langauage. Version (a) works on an EREW PRAM and is related to the DFG presented in figure 1.2.a. Version (b) works on a CUMULATIVE-ERCW; the corresponding DFG is presented in figure 1.2.b.

```
Product(a : in E,
        b : in E,
        c : out E)
begin
  c.Write( a.Read()*b.Read() );
end

IterProd( n       : in  integer,
          a[1..n] : in  array of shared E,
          res     : out shared E)
begin
  if(n==1)
    res.Write( a[1].Read() );
  else
    tmp1i, tmp2 : shared E;

    fork IterProd(n/2, a[1..n/2], tmp1);
    fork IterProd(n-n/2, a[n/2+1..n], tmp2);
    fork Product(tmp1, tmp2, res);
  end if
end
```

(a)

```
IterProd( n       : in  integer,
          a[1..n] : in  array of shared E,
          res     : out shared E )
begin
  if(n==1)
    res.Cumul<*>( a[1].Read() );
  else
    fork IterProd( n/2, a[1..n/2], res );
    fork IterProd( n-n/2, a[n/2+1..n], res );
  end if
end
```

(b)

Figure 1.3: ATH code of two iterated products: (a) EREW, (b) cumulative-ERCW. *Data in shared memory are explicitly declared by the prefix* `shared`. Notation `x.f()` *means that function* `f` is called on the data in shared memory `x`. *In program (b), the function call* `x.Cumul<*>( v )` specifies a cumulative concurrent write on the data in shared memory `x`*; the commutative and associative binary function implementing the operation is* `*`.

1.1.4 Time, work and communication costs

Consider a PRAM program. In the following, n denotes the size of the input. The arithmetic cost is characterized by:

- the *parallel time* $T(n)$ which corresponds to the number of executed steps;
- the *arithmetic work* $W_a(n)$, i.e. the whole number of operations performed.

Those quantities are independent of the number of processors and thus may be defined directly from the DFG description of the execution.

Definition 1 *The parallel time* $T(n)$ *is the maximal depth of DFG(*x*) for any input* x *of size* n*:*

$$T(n) = \max_{x, \|x\|=n} Depth(DFG_a(x)) \tag{1.1}$$

The arithmetic work $W_a(n)$ is the number of instruction nodes of DFG(x) for any input x of size n:

$$W_a(n) = \max_{x,\|x\|=n} \#V(DFG_a(x)) \tag{1.2}$$

The arithmetic cost is denoted:

$$O_a(T(n), W_a(n)) \tag{1.3}$$

Similarly, the communication cost is characterized by two factors:

- the *communication delay*[3] $C_d(n)$, i.e. the maximal number of global memory access performed by a processor;
- the *communication work* $W_c(n)$, i.e. the whole number of global memory access performed.

The PRAM program implements a scheduling of the DFG on an infinite number of processors: any access to the local memory on each processor is not considered as a communication. Thus, the communication cost may vary depending on the number of processors used in the program.
To define communication cost with respect to a parallel algorithm (independent of a number of processors, and so more general than the program that implements it), we will refer to its DFG.

Definition 2 *The communication work $W_c(n)$ is the maximal number of edges for any input of size n:*

$$W_c(n) = \max_{x,\|x\|=n} \#E(DFG(x)) \tag{1.4}$$

The communication delay $C_d(n)$ is the maximal length of a path in DFG from an input data to an output one:

$$C_d(n) = \max_{x,\|x\|=n} Depth(DFG(x)) \tag{1.5}$$

The communication cost is denoted:

$$O_c(C_d(n), W_c(n)) \tag{1.6}$$

In order to compare arithmetic and communication costs, the granularity $g(n)$ is defined.

Definition 3 *The* granularity *$g(n)$ is the ratio between the arithmetic and communication works:*

$$g(n) = \frac{W_a(n)}{W_c(n)} \tag{1.7}$$

Remark. Previous costs are defined at for DFGs with unit time instructions and unit size transitions. For general non unit size DFGs (denoted as macro data-flow graphs), costs are weighted by the size of each node: either the number of elementary instructions for a job node or the size of related data for a transition node. Macro data-flow graphs will be specifically studied in chapter 2.

[3] $C_d(n)$ is called *communication complexity* in [28].

1.1.5 Efficient algorithms

Let $\mathcal{A}$ be an algorithm with cost $T(n), W_a(n), C_d(n), W_c(n)$. Let $W_s(n)$ be the work of the best known (sequential) algorithm that solves the same problem.

The building of a parallel algorithm to solve a given problem may be aimed at different directions:

- either finding the smallest amount of time required to solve a problem. In this context, the class NC of problems that may be solved in parallel time $T(n) = \log^{O(1)} n$ using a polynomial number of processors $W_a(n) = n^{O(1)}$ plays a central role.
- or building an *efficient* program that leads to solve larger problems in a reasonable amount of time taking benefit of the ability to use several processors, let us say p. Here, arithmetic and communication overheads (i.e. $W_a(n)$ and $W_c(n)$) are to be carefully taken into account in order to guarantee efficient executions.

A common trade-off [38] consists in building parallel algorithms that:

- have *polynomial speed-up*, i.e.

$$T(n) = O(W_s(n)^\epsilon) \quad \text{with } \epsilon < 1. \tag{1.8}$$

- are *work-preserving*, i.e.

$$W_a(n) = \Theta(W_s(n)). \tag{1.9}$$

  The *inefficiency* ν measures the arithmetic overhead:

$$\nu(n) = \frac{W_a(n)}{W_s(n)}. \tag{1.10}$$

- require *few communications*, i.e

$$W_c(n) = O(W_a(n)^\epsilon) \quad \text{with } \epsilon < 1. \tag{1.11}$$

  Such an algorithm is also said *local* or *of coarse-granularity* or *with polynomial granularity* (note that $g(n) = \Omega(n^\alpha)$ with $\alpha > 1$).

Definition 4 *$\mathcal{A}$ is said:*

- fast *if it achieves poly-logarithmic parallel time with a polynomial number of operations, i.e.* $T(n) = \log^{O(1)} n$ *and* $W_a(n) = n^{O(1)}$.
- optimal *if it is fast and has constant inefficiency.*
- efficient *if it has a polynomial speed-up and a constant efficiency.*

In order to not absolutely reject fast algorithms involving a small overhead in arithmetic operations, fast algorithms with poly-logarithmic inefficiency will be considered as efficient *also.*

In the following, some main techniques that lead to the building of an efficient *and* of coarse-granularity algorithm are overviewed. It turns out that minimizing time without preserving work (i.e. building NC algorithm) is of specific interest:

- algorithmic techniques involved for both are very close;
- it gives a lower bound on the best parallel time that may be achieved;
- an inefficient but fast algorithm may successfully be coupled to a slower but efficient one to build a faster program.

1.1.6 Example

We illustrate the previous definitions on the iterated sum algorithm presented in figure 1.3.a. Scalar product of two vectors is directly reduced from iterated sum; it may be applied to perform matrix multiplication in a semi-ring.

Iterated sum

For the EREW algorithm presented in figures 1.3.a and 1.2.a (balanced tree computation scheme), we assume $n = 2^m$:

$$\begin{array}{ll} T(n) = \log n & C_d(n) = \log n + 1 \\ W_a(n) = n - 1 & W_c(n) = 2n - 1 \end{array} \tag{1.12}$$

This algorithm is optimal since its cost is – asymptotically – a lower bound.

As a consequence, the scalar product of two vectors is computed on an EREW with cost:

$$O_a(\log n, n) \quad \text{and} \quad O_c(\log n, n). \tag{1.13}$$

On a semi-ring, $+$ is commutative. Thus, on a cumulative-CRCW PRAM, this problem may be computed with parallel cost (fig. 1.2.a):

$$O_a(1, n) \quad \text{and} \quad O_c(1, n). \tag{1.14}$$

However, the description of the computation scheme (cf program in fig. 1.3.b) may require $O_a(\log n, n)$.

Matrix product

Consider the problem of computing a square matrix product $C = AB$ in a semi-ring (i.e. using only $+$ and $\times$ operations).

Let n be the dimension of the matrices: since $C_{i,j} = \sum_{k=1}^{n} A_{i,k} B_{k,j}$, the problem reduces to n^2 independent scalar products. Using 1.13, we obtain a parallel algorithm with cost:

$$O_a(\log n, n^3) \quad \text{and} \quad O_c(\log n, n^3). \tag{1.15}$$

Since $W_s(n) = \Theta(n^3)$ [37], this algorithm is efficient.

However, $g(n) = O(1)$ and it is not coarse-granularity. Besides, it can be seen that, if E is a field (or ring), the above algorithm is not efficient (polynomial inefficiency) neither theoretically since $W_s = O(n^{2.376})$ [15, 45] nor practically since $O(n^{2.81})$ algorithms are of practical use [3, 40, 17]. We will see in following sections how to overcome those problems.

1.1.7 Relations between PRAMs

We consider the cost of the execution of a parallel algorithm (defined on a CUMULATIVE-CRCW PRAM for instance) on a given PRAM with a fixed number of processors and with its own semantics for access in shared memory. Two cases are distinguished: when the number of processors is decreased and when memory access are restricted. We consider here only arithmetic costs. The main consequence is the existence of optimal – within a constant factor – simulations of a CRCW algorithm that uses an unbounded number of processors on an EREW machine with a fixed number of processors.

Theorem 1 Fine grain simulation with fewer processors - Brent's principle *[9, 28]. Let $\mathcal{A}$ be an algorithm that can be implemented to run in (arithmetic) parallel time T and work W_a on a given PRAM with an unbounded number of processors. If each local access corresponds to a global one, then $\mathcal{A}$ can be scheduled on the same PRAM, but with p processors, to run in (arithmetic) parallel time $T_p(n)$:*

$$\left\lceil \frac{W_a(n)}{p} \right\rceil \leq T_p(n) \leq \left\lfloor \frac{W_a(n)}{p} \right\rfloor + T(n) \tag{1.16}$$

It can be noted that this fine grain simulation does not take into account additive cost due to the computation of the schedule [12, 22].

Remark. In chapter 2, theorem 10 gives a more general simulation result with analoguous bounds. It consists in a constructive coarse grain simulation for DFGs where arithmetic nodes may represent a sequence of elementary instructions.

Theorem 2 Simulation with restricted access in global memory *[28, 38]. Let $\mathcal{A}$ be an algorithm that can be implemented to run in (arithmetic) parallel time T_p on a CUMULATIVE-CRCW PRAM with p processor. Then, $\mathcal{A}$ can be implemented on an EREW PRAM with p processors to run in time $O(T_p \log p)$.*

1.2 Increasing granularity

Efficient parallel algorithms require near-optimal work; obviously, the careful analysis of the smallest depth DFG induced by a sequential algorithm among the best is then of practical interest.

As a major example, sequential algorithms based on a partitioning of the problem into – many – independent subproblems have intrinsic parallelism if partitioning and merging (to recover the global solution) steps are either parallel or of neglected cost. This situation appears frequently in numerous divide&conquer algorithms (let us say *parallel divide&conquer*). As a computer algebra instance, modular methods based on Chinese remainder computations [2, 10] amount to this scheme.

Once a fine grain fast parallel algorithm is built, increasing granularity is required to obtain an efficient algorithm with coarse-granularity. In this section, the technique consisting in stopping the recursive splitting is illustrated on the matrix product problem; we prove an optimal granularity for this problem.

1.2.1 Parallel divide and conquer

Let us consider the example of matrix multiplication using a standard bi-dimensional block algorithm:

$$\begin{bmatrix} A_{11} & A_{12} \\ A_{21} & A_{22} \end{bmatrix} \begin{bmatrix} B_{11} & B_{12} \\ B_{21} & B_{22} \end{bmatrix} = \begin{bmatrix} A_{11}B_{11} + A_{12}B_{21} & A_{11}B_{12} + A_{12}B_{22} \\ A_{21}B_{11} + A_{22}B_{21} & A_{21}B_{12} + A_{22}B_{22} \end{bmatrix}. \tag{1.17}$$

All block matrices products, of dimension $n/2$, can be multiplied in parallel. Applying recursively this splitting scheme leads to a parallel algorithm with cost:

$$O_a(\log n, n^3) \quad O_c(\log n, n^3) \tag{1.18}$$

Note that, since coefficient addition is associative, each entry in the output matrix may be computed as an iterated sum of n values. This allows the whole computation to take a time $\log n$ (instead of $\log^2 n$ if additions where performed naively at each step). This remark appears directly on the DFG description for a CUMULATIVE-CRCW PRAM 1.4: all final sums are made in O(1) time. But the splitting process, which involves no arithmetic operation but recursive forks (cf fig. 1.3.b), requires $O(\log n)$ time using recursive forks[4]. Another technique to obtain $O_a(\log n, n^3)$ consists in pipelining additions [1].

Remark. The same strategy applied to Strassen's algorithm leads to a parallel algorithm with cost:

$$O_a(\log n, n^{\log_2 7}) \quad O_c(\log n, n^{\log_2 7}) \tag{1.19}$$

Optimal in work (on a semi-ring), this algorithm has granularity $g(n) = O(1)$: it is roughly equivalent to a recursive version of 1.15). In the next section, we detail how to increase granularity in order to build an efficient algorithm with coarse-granularity.

1.2.2 Minimizing communication work

Obtaining a coarse-granularity algorithm requires to minimize communications. This can be done by stopping the recursive parallel splitting process at a given depth, let us say when sub-matrices are of size lesser than k (i.e. depth $\log \frac{n}{k}$). Operations – resp. sums and products – on matrices of dimension k are then performed sequentially, using an optimal algorithm – resp. in time $O(n^2)$ and $O(n^3)$ –. The cost is then:

$$O_a\left(k^3 + \log n, n^3\right) \quad O_c\left(k^2 + \log \frac{n}{k}, \frac{n^3}{k}\right) \tag{1.20}$$

which gives an algorithm with granularity $g(n) = k$. We thus obtain a parallel efficient algorithm with arbitrary (polynomial) granularity.

Theorem 3 *For any g, $\log^{1/3} n \leq g \leq n$, two $n \times n$ matrices can be multiplied by an algorithm of granularity g with parallel cost:*

$$O_a\left(g^3, n^3\right) \quad O_c\left(g^2 + \log n, \frac{n^3}{g}\right).$$

[4]Note that the brute force program (fig. 1.4) which performs iteratively fork instructions requires $O_a(n^3, n^3)$!

```
CumulProductTerm( a : in  E,
                  b : in  E,
                  c : out E )
begin
  c.Cumul<+>( a.Read()*b.Read() );
end

MatrixProduct( n : in  integer;
               a : in  array[1..n,1..n] of E,
               b : in  array[1..n,1..n] of E,
               c : out array[1..n,1..n] of E )
begin
  i, j, k : local integer;

  for i = 1..n loop
    for j = 1..n loop
      for k = 1..n loop
        fork CumulProductTerm( a[i,k], b[k,j], c[i,j] );
      end loop
    end loop
  end loop
end
```

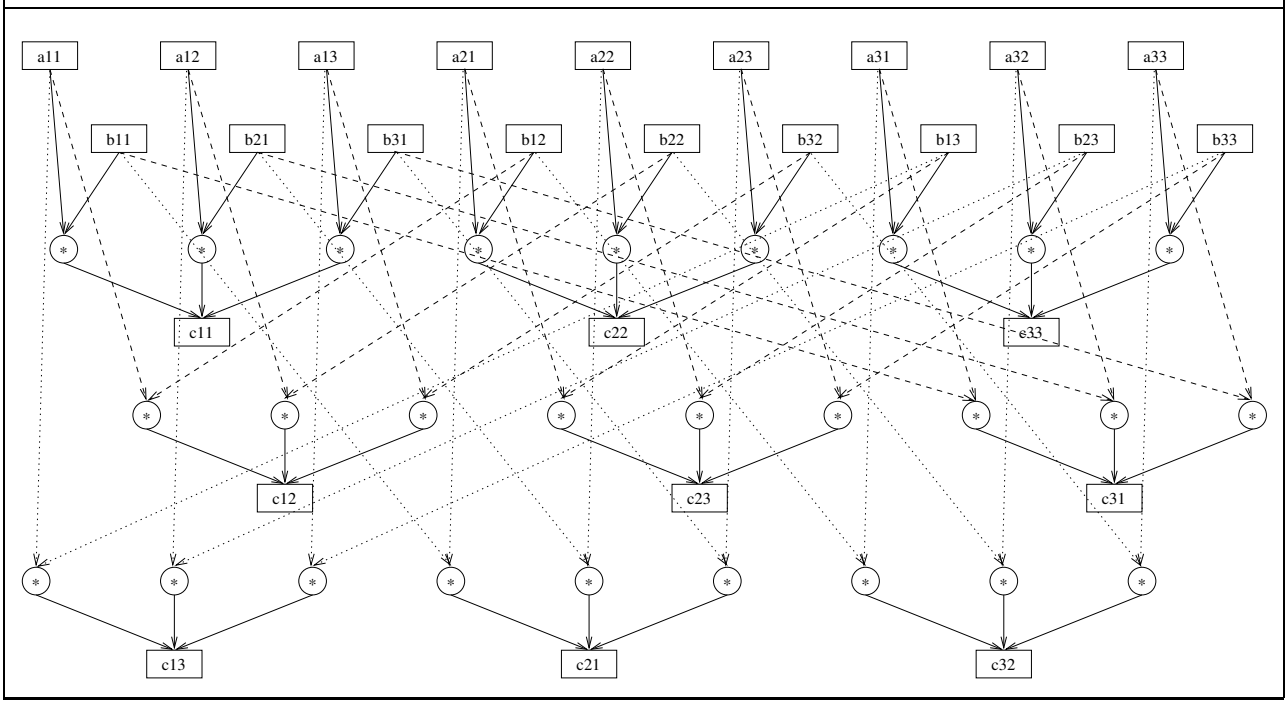

Figure 1.4: DFG of the multiplication of two 3×3 matrix (cumulative-CRCW)

The previous algorithm 1.20 proves the upper bound. □.

The following theorem gives lower bounds for communication costs. It shows that the previous algorithm achieves an optimal communication delay and an optimal granularity among algorithms that achieve an optimal communication delay.

Theorem 4 *Let $\mathcal{A}$ be an efficient parallel algorithm that multiplies two matrices of dimension n in time T using $(+, \times)$ only and performing $\Theta(n^3)$ operations. Then,*

$$C_d = \Omega\left(T^{2/3} + \log n\right) \qquad W_c = \Omega\left(\frac{n^3}{C_d^{1/2}}\right).$$

Since $\mathcal{A}$ is efficient, $T = O(n^\epsilon)$ with $\epsilon < 3$; by reduction from iterative sum, we thus have $C_d = \Omega(\log n)$.

Kerr [37, 1] shows the lower bound $\Omega(n^3)$ on the arithmetic work. Since $\mathcal{A}$ performs $\Theta(n^3)$ operations, its execution can be scheduled in time $\Theta(T)$ using $p = \frac{n^3}{T}$ processors. Let $s_i, 1 \leq i \leq p$, be the number of shared memory access performed by processor i. We then have $W_c = \sum_{i=1}^{p} s_i$ and $C_d \geq \max_{i=1}^{p} s_i$. To obtain a lower bound on W_c and C_d, we use the following lemma [1, 25]: if a processor reads at most s elements of input matrices and computes at most s partial sums of their product, then this processor can compute no more than $s^{3/2}$ multiplicative terms for these partial sums.
Applying this lemma to p_i which reads or writes at most s_i elements and since $\Omega(n^3)$ multiplicative terms are to be computed, we have:

$$\sum_{i=1}^{p} s_i^{3/2} = \Omega(n^3). \tag{1.21}$$

Bounding s_i by C_d and replacing p by $\frac{n^3}{T}$ leads to:

$$C_d = \Omega\left(T^{2/3}\right). \tag{1.22}$$

Noticing that $\sum_{i=1}^{p} s_i^{3/2} \leq C_d^{1/2} \sum_{i=1}^{p} s_i$, we obtain:

$$W_c = \Omega\left(\frac{n^3}{C_d^{1/2}}\right) \tag{1.23}$$

which concludes the proof $\square$.

Recursive multiplication algorithms. A similar study can be applied to other recursive matrix multiplication algorithms (e.g. Strassen). It also lead to efficient parallel algorithms with both polynomial speed-up and polynomial granularity that lead to performant implementations [17].

1.2.3 Conclusion

In this section, we have studied the DFG of a sequential algorithm, based on a divide&conquer scheme, that contains inherent parallelism. By halting the recursive process in order to minimize communications, we have exhibited a family of efficient parallel algorithms with arbitrary coarse-grain granularity.

Due to its practical interest, this technique has been successfully applied to various problems. One of significant interest in computer algebra is the discrete Fourier transform. The direct analysis of the FFT algorithm leads to a parallel algorithm with cost:

$$O_a(\log n, n \log n) \quad O_c(\log n, n \log n).$$

A clustering of elementary instructions (block clustering on the first $\frac{\log n}{2}$ steps and cyclic clustering on the last $\frac{\log n}{2}$ steps, cf fig. 1.5) leads to an algorithm with parallel cost [41, 39]:

$$O_a(\sqrt{n} \log n, n \log n) \quad O_c(\sqrt{n}, n).$$

This algorithm has polynomial speed-up, optimal work and achieves also optimal granularity [1].

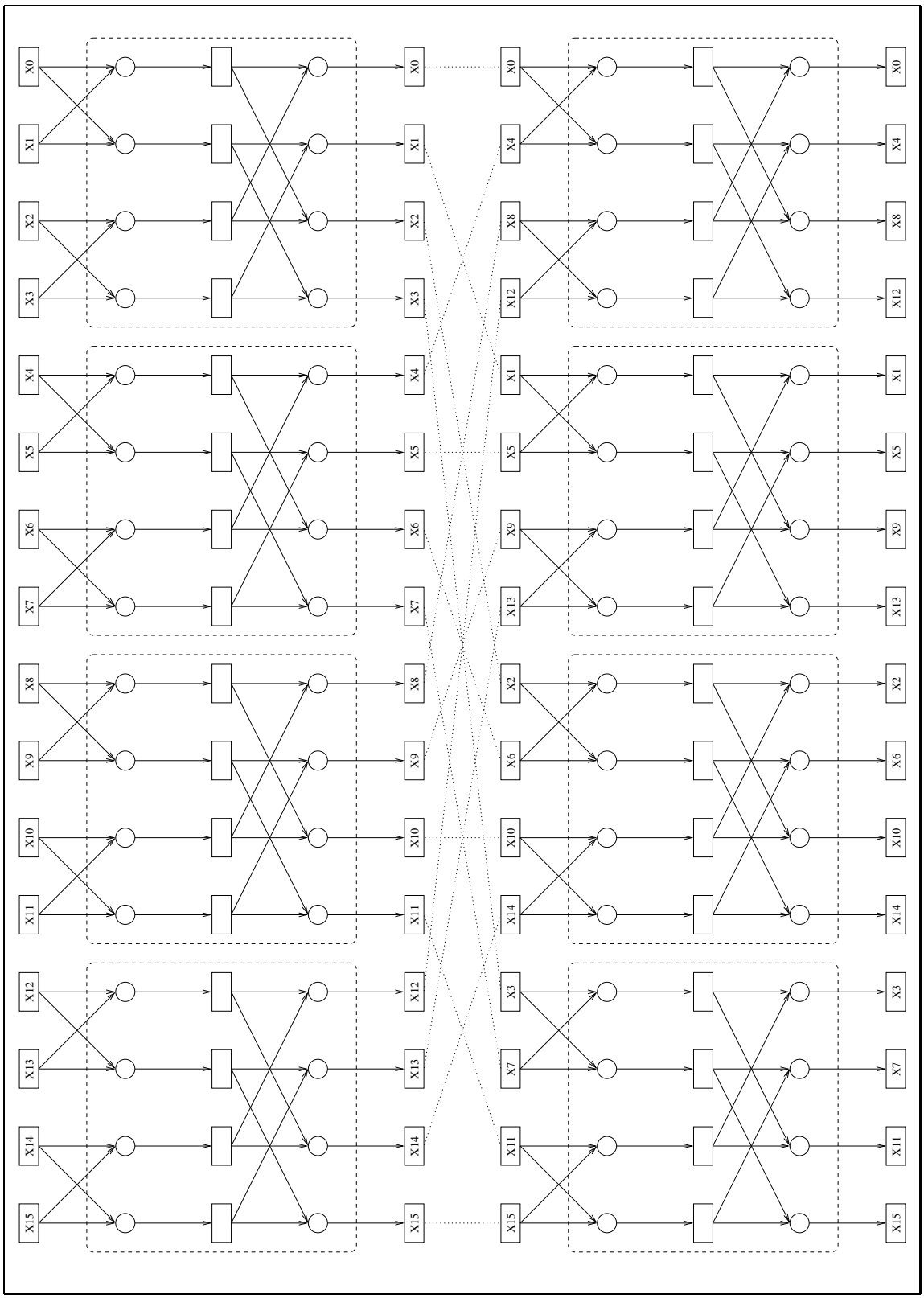

Figure 1.5: DFG of the EREW $O_a(\sqrt{n}\log n, n\log n)$ FFT algorithm of 16 points. There are $2\sqrt{n}$ arithmetic tasks (represented by square boxes embedding elementary operations and local dependencies), each corresponding to a sequential FFT computation on $\sqrt{n}$ points. For any task on the left, shared data dependencies imply a precedence relation with the $\sqrt{n}$ tasks on the right.

The resulting algorithm is based on coupling a very fast parallel algorithm, optimal in time but requiring many communications, to a sequential one which minimizes communication. Such an algorithm is called "poly-algorithm"; the technique that underlies this coupling is called "cascading divide&conquer".

Cascading divide&conquer may be applied in a more general context, by coupling a very fast parallel algorithm, yet requiring many operations, to a slower one which performs an optimal number of operations. This technique makes the building of very fast algorithms attractive even if the required number of operations is larger.

1.3 Breaking data-flow dependencies by redundancy and cascading divide&conquer

It may appear that DFGs related to a sequential algorithm contain data-dependencies that bound parallelism. Introducing redundant computations may then allow to break dependencies in order to minimize parallel time. Cascading divide&Conquer may then be used to obtain an optimal arithmetic work. In this section we illustrate this technique on the computation of the solution of a triangular linear system presented in [46]. We focus on communication costs.

Let A be an $n \times n$ nonsingular triangular matrix with coefficients in a field K. We assume by convenience $n = 2^m$. Let b a vector in K^n. We consider the computation of $x = A^{-1}b$.

1.3.1 DFG of the best sequential algorithm

The simple forward substitution algorithm has sequential cost $W_s(n) = \Theta(n^2)$. Direct analysis of its DFG (see fig. 1.6) gives its parallel cost:

$$O_a(n, n^2) \qquad O_c(n, n^2), \tag{1.24}$$

which leads to an algorithm with polynomial speed-up but small granularity $g(n) = O(1)$.

If entries of A are in global memory after initialization, we have $W_c(n) = \Omega(n^2)$. In a view to minimizing the communications involved by the algorithm itself, in the following we do not consider the access to A in the communication work $W_c(n)$.

In order to increase granularity, we consider a divide&conquer version of this algorithm [7]. Let A, b and x be divided into blocks:

$$A = \begin{bmatrix} A_{11} & 0 \\ A_{21} & A_{22} \end{bmatrix} \qquad b = \begin{bmatrix} b_1 \\ b_2 \end{bmatrix} \qquad x = \begin{bmatrix} x_1 \\ x_2 \end{bmatrix}. \tag{1.25}$$

Here A_{11} is of size $h \times h$, x_1 and x_2 are of size h. We have:

$$A_{11}x_1 = b_1 \quad \text{and} \quad A_{22}x_2 = b_2 - A_{21}x_1. \tag{1.26}$$

where x_1 and x_2 are computed recursively using the same algorithm; $A_{21}x_1$ is computed using a scalar product (see 1.13). Note that the use of a pipeline scheme leads to the previous parallel cost 1.24.

```
Update( x : out E,
        a : in  E,
        y : in E )
begin
  x.Cumul<+>( -a.Read()*y.Read() );
end


FinalDivision( x : in and out E,
               a : in         E )
begin
  x.Write( x.Read() / a.Read() );
end




TriangularSolve ( n : in  integer,
                  a : in  array[1..n, 1..n] of E,
                  b : in  array[1..n] of E,
                  x : out array[1..n] of E )
begin
  i,j : local integer;

  for i = 1..n loop
    x[i].Cumul<+>( b[i].Read() );
    fork FinalDivision(x[i], a[i,i]);
    for j = (i+1)..n loop
      fork Update(x[j], a[j,i], x[i]);
    end loop
  end loop
end
```

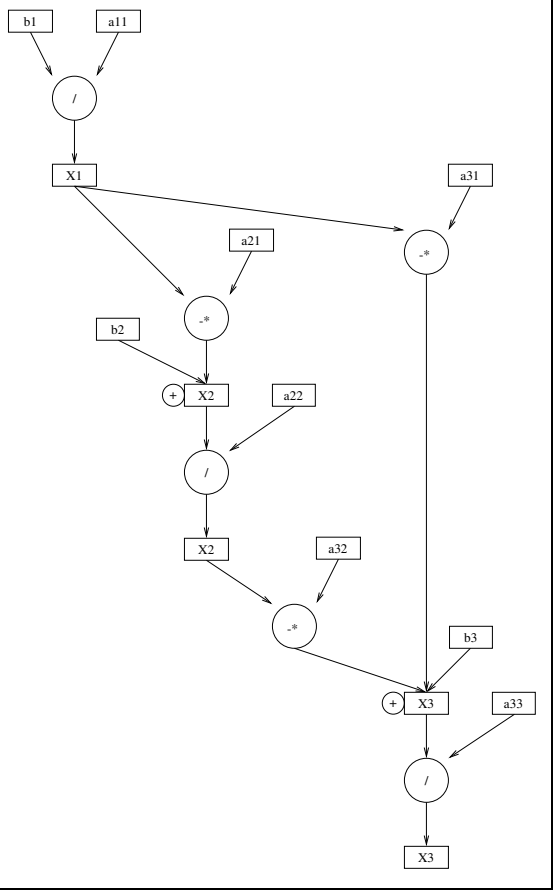

Figure 1.6: DFG for the solving of a 3×3 nonsingular triangular matrix

We may then stop the recursive splitting when matrices are of size $k \times k$, and use sequential algorithms (triangular system inversion and matrix-vector product) on matrices of size lesser than k. The resulting parallel cost is:

$$O_a(nk, n^2) \qquad O_c\left(nk, \frac{n^2}{k}\right) \tag{1.27}$$

which leads to an algorithm with granularity $g(n) = O(k)$.

Theorem 5 *For any $\epsilon < 1$, a triangular nonsingular linear system can be solved by an efficient parallel algorithm of coarse granularity n^ϵ in time $O(n^{1+\epsilon})$.*

Choosing $k = n^\epsilon = o(n)$ in 1.27 proves the upper bound. □.

1.3.2 Breaking dependencies

The linear time lower bound on previous algorithm time comes from the dependency in formula 1.26 between computations of x_1 and x_2. This dependency may be broken by directly computing the inverses of the triangular nonsingular matrices A_{11} and A_{22}.

Consider the matrix A split in four blocks of dimension $n/2$ (1.25 with $h = n/2$). Then we have:

$$A^{-1} = \begin{bmatrix} A_{11}^{-1} & 0 \\ -A_{22}^{-1}A_{21}A_{11}^{-1} & A_{22}^{-1} \end{bmatrix} \tag{1.28}$$

From theorem 3, the product of two matrices of dimension n is computed with parallel cost $O_a(\log n, n^3)$. In the following, we will refer to this cost.

To compute the inverse of A from 1.28, we first compute recursively and in parallel A_{11}^{-1} and A_{22}^{-1}. Then we compute the last block of A^{-1} by performing sequentially two parallel matrix products. The parallel cost for inverting A is then:

$$O_A(log^2 n, n^3) \qquad O_c\left(log^2 n, \frac{n^3}{\log^{1/3} n}\right) \tag{1.29}$$

Once A^{-1} is computed, $x = A^{-1}b$ can be computed with the same cost. However, even if polylogarithmic in time, this algorithm has polynomial inefficiency. In the next paragraph, we use it on A_{11} in 1.26 in order to decrease parallel time.

Remark. The above algorithm is efficient for computing the inverse of a nonsingular triangular matrix. Note that by using fast matrix multiplication, the parallel cost is reduced to $O_a(\log^2 n, n^\omega)$ with $\omega < 2.38$ [46]. Besides, if computations are performed sequentially when the dimensions of the matrices are lesser than $k \geq n^\epsilon$, $(\epsilon < 1)$, the obtained algorithm is efficient and has polynomial speed-up and polynomial granularity.

1.3.3 Cascading divide&conquer to minimize time

The previous algorithm is not efficient but may be combined to the recursive sequential algorithm (formula 1.26). The trick is to use it on small dimension matrices (let us say h) when the overhead $O(h^3)$ due to the fast inversion of such a matrix becomes neglectible compared to coefficients updates (roughly nh). This leads to the following algorithm of Pan&Preparata [46].

Theorem 6 *The solution of a nonsingular triangular system can be computed in*

$$O_a(n^{1/2}\log n, n^2)$$

using a standard n^3 matrix multiplication algorithm.
If a fast n^ω multiplication is used then the parallel cost is:

$$O_a(n^{(\omega-2)/(\omega-1)}\log^2 n, n^2).$$

The following 1.27, let A be split in n^2/h^2 blocks of size $h \times h$. Though, note that a direct computation (see theorem 1.27) leads to a parallel time $O(n^{1/2}\log^2 n)$. To avoid the $\log n$ overhead factor in the parallel time, we proceed by gathering computation on $\log^2 h$ blocks.
Let $k = h\log h$; the matrix A may be seen as split in $(n/k)^2$ blocks, each block consisting in $\log^2 h$ sub-blocks of dimension h (cf fig. 1.7).
We use the sequential iterative algorithm on the $(n/k) \times (n/k)$ coarse grain matrix. At step i, we

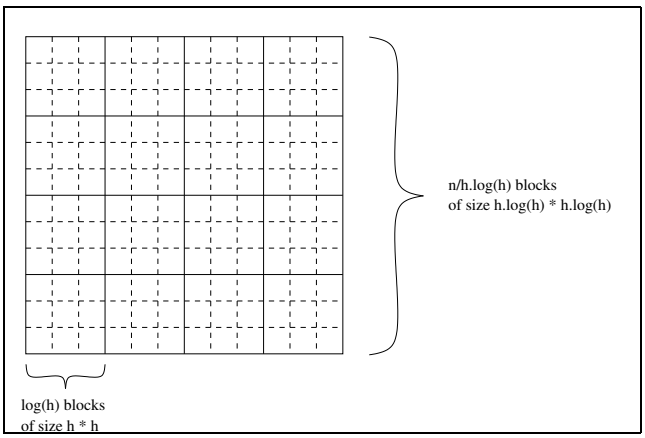

Figure 1.7: Splitting used for $h = 8$, $h \log h = 24$, $n = 96$

have to invert the triangular system corresponding to the diagonal block (i, i). For this computation, we first invert concurrently the $\log h$ diagonal sub-blocks of this block. Then, we update others sub-blocks of x_i. At the end of the step, blocks x_j, for $j > i$, are updated.
The algorithm is the following:

Initialization.

Let A be split into n/k blocks $M_{i,j}$ of dimension k ($k = h \log h$). For $1 \leq j \leq i \leq n/k$, let $M_{i,j}$ be split into $\log h \times \log h$ block $m_{i,j}^{k,l}$ of dimension h.
Let x be initialized to b and split according to A.

for $i = 1..n/k$ do

1. for $j = 1..\log h$ do
   fork $\left(m_{i,i}^{j,j}\right)^{-1} = \text{invert}(m_{i,i}^{j,j})$.
   Using fast inversion and Brent's principle, the cost is $O_a(\log^2 h, h^3 \log h)$.
2. for $j = 1..\log h$ do
   update x_i^j in parallel
   $x_i^j = \left(m_{i,i}^{j,j}\right)^{-1} \left(x_i^j - \sum_{l=1}^{j-1} m_{i,i}^{j,l} x_i^l\right)$
   Scalar product are performed in parallel: thus x_i is computed with a cost $O_a(\log^2 h, h^2 \log h)$.
3. for $j = i + 1..n/k$ fork update x_j in parallel
   $x_j = x_j - M_{j,i} x_i$
   Performing scalar product in parallel, the cost is $O_a(\log h, nh \log h)$.

The final cost is : $O_a(n \log^2 h/k, n/k \max(h^3 \log h, nh \log h))$. Since $k = h \log h$, it reduces to:

$$O_a(n \log h/h, \max(nh^2, n^2)),$$

and the optimal value for h is the larger one that leads to a work $W_a(n) = O(n^2)$. Thus, we choose $h = n^{1/2}$ and we obtain the upper bound.

The same technique is applied to obtain the upper bound when a fast matrix multiplication algorithm is used. □

1.3.4 Applications in linear algebra

Many linear algebra algorithms are based on a Gaussian elimination scheme: linear system solving, normal forms (Hessenberg, Smith, Frobenius, symbolic Jordan). Such a scheme provides parallel algorithms with polynomial speed-up: at each step, a transformation is computed that can then be applied in parallel to each coefficient of the matrix. For instance, solving a non-singular linear system using standard Gaussian elimination leads to a parallel algorithm with cost:

$$O_a(n, n^3) \quad O_c(n, n^2) \tag{1.30}$$

Moreover, very fast deterministic algorithms (polylogarithmic parallel time) are known for most problems [45, 24, 58, 57] but they are often inefficient ($W_a(n) = n^{O(1)} W_s(n)$). For instance, solving a non-singular linear system can be computed in parallel with cost:

$$O_a(log^2 n, n^{3+\alpha}) \tag{1.31}$$

with $\alpha = 1/2$ in characteristic zero [16, 50] and $\alpha = 1$ in the general case [11]. Applying the same cascading divide and conquer strategy leads to sub-linear parallel algorithms with optimal[5] work [46]:

$$O_a(n^{1/5} \log^2 n, n^3). \tag{1.32}$$

Remark. The same technique applied on Strassen formulation [56] (which may take benefit of fast $O(n^{2.376})$ matrix multiplication algorithms), does not succeed in the building of a sub-linear algorithm with parallel time n^β, $\beta < 1$.

1.3.5 Conclusion

In this paragraph, we have used bi-dimensional block matrix partitioning in order to:

- increase the granularity to build polynomial speed-up algorithms with polynomial granularity; the technique used is cascading divide and conquer with a sequential algorithm in order to decrease communication costs.

- decrease parallel time while preserving the work; the technique used is cascading divide and conquer with a very fast but inefficient algorithm in order to make the computation faster.

In [46], the same technique, called *work-preserving speed-up*, is applied to several linear algebra algorithms: LU factorization, inversion, quasi-inversion, solution of linear structured systems.

[5] relatively to the standard $O(n^3)$ sequential algorithm

1.4 Randomization to decrease time or preserve work.

When an algorithm has a bounded degree of parallelism or a polynomial efficiency, randomization may help in order to either decrease time or preserve work, eventually both. This section illustrates both aspects on the computation of the rank of a matrix.

In computer algebra, randomization is most often introduced via the verification of a polynomial identity by evaluation on a random value. Testing whether a polynomial is identically zero can deterministically be solved by evaluating the polynomial, represented as a straight-line program, at a sufficient number of points. However, depending on the degree and on the number of indeterminates, such a deterministic test can require a huge number of evaluations. Following theorem, due to Schwartz [54], uses randomization in order to reduce this number while bounding the probability of failure.

Theorem 7 *[54, 28] Let $P(x_1, \ldots, x_n)$ be a polynomial in the variables (x_i), $1 \leq i \leq n$, over a field K. Let I be a finite subset of K with cardinal c. Let $(\alpha_1, \ldots, \alpha_n)$ be a vector selected at random in K^n. If P is not identically zero then*

$$\textit{Prob}\left(P(\alpha_1, \ldots, \alpha_n)) = 0\right) \leq \frac{\textit{deg}(P)}{c}.$$

Once a problem is reduced to the verification of a polynomial identity, this theorem allows to build a Monte-Carlo algorithm to solve it (for an introduction on Monte-Carlo and Las Vegas algorithms, see [36]). It is sufficient to build a parallel algorithm that evaluates the polynomial at a given input point. By choosing this point at random in a large enough finite subset[6] we obtain a Monte-Carlo algorithm whose probability of error is at most $1/2$. This technique may be applied in a very large framework [36, 28] and is commonly used in computer algebra [45] to build fast algorithms with optimal work. We illustrate it on the problem of computing the rank of a matrix.

In the following, A denotes a matrix of dimension $n \times n$ with coefficients in a field K. For the sake of simplicity, K is assumed infinite.

1.4.1 Randomization to suppress dependencies

The rank of a matrix can be computed using a standard pivoting Gaussian elimination. Similarly to 1.24, this results in an algorithm with parallel cost:

$$O_a(n, n^3) \quad O_c(n, n^2) \tag{1.33}$$

On the contrary to triangular system solving, the computation scheme (DFG) is relatively unknown: coefficients to modify are determined at each step only once the pivot element has been chosen.

In [8], randomization is used in order to reduce the whole problem to a fixed DFG on which parallelization techniques can be applied. The algorithm is based on the following characterization of the rank: rank$(A) = r$ iff there exist two non-singular matrices L and C such that the principal minor of dimension r in LAC is non zero while principal minors of dimension larger than r are

[6] Note that, if K is not large enough, this may require to work in an extension of K [24].

zero. Moreover, L and C can be taken at random with a high probability of success: the use of theorem 7 to evaluate this probability requires to express the problem as a polynomial identity.

Let $\delta_i(L, C)$ denote the principal minor of dimension i of LAC. Due to multi-linearity of the determinant, δ_i is a polynomial of degree $2n$ with indeterminates $L_{i,j}$ and $C_{i,j}$ ($1 \leq i, j \leq n$). Previous rank characterization leads to the following polynomial identities:

$$\begin{array}{ll} \delta_i \neq 0 & 1 \leq i \leq r \\ \delta_i = 0 & r < i \leq n \end{array} \tag{1.34}$$

This suggests the following Monte-Carlo algorithm to compute r:

1. Choose two random non-singular matrices L and C with coefficients in a finite subset of cardinal c of K;
2. Compute: $M = LAC$;
3. For $1 \leq i \leq n$, compute $d_i = \det(M_i)$ and let $d_0 = 1$;
4. Return $s = \text{Max}_{k=0,\ldots,n}\{k/d_k \neq 0\}$.
   (Note that step 3 and 4 may be replaced by a logarithmic search to compute s).

In any case, $s \leq r$. The probability of error, which occurs when $s < r$, corresponds to executions where the evaluation d_r of polynomial δ_r is zero although δ_r, of degree $2n$, is not identically zero. From theorem 7, this probability is bounded by $\frac{2n}{c}$. Choosing $c = 4n$ results in a Monte-Carlo algorithm with probability of error lesser than $\frac{1}{2}$.

Arithmetic cost is dominated by the computation of the n determinants. If Chistov's method [11] is used, this cost is:

$$O_a(log^2 n, n^{\omega+1}) \tag{1.35}$$

In order to improve efficiency, determination of s may be computed using a logarithmic scheme instead of the previous brute force method. Using an efficient randomized algorithm to compute the determinant (for instance the randomized one of Kaltofen and Pan [33], the parallel cost becomes

$$O_a(log^3 n, n^\omega \log n), \tag{1.36}$$

From Monte-Carlo to Las Vegas. The building of a Las Vegas algorithm from a Monte-Carlo one mainly consists in verifying that the output is a correct solution to the initial problem. Such a verification is easy from the previous algorithm; it suffices to verify that all columns (resp. rows) of the matrix $M = LAC$ are linear combinations of s independent columns (resp. rows) in M, s being the output of the algorithm.

Consider the following splitting for M, the first block M_{11} being of size $s \times s$:

$$M = \begin{bmatrix} M_{11} & M_{12} \\ M_{21} & M_{22} \end{bmatrix}. \tag{1.37}$$

M_{11} is a non-singular matrix. Let $X = M_{21}M_{11}^{-1}$ and $Y = M_{11}^{-1}M_{12}$; note that X and Y^t are of size $(n - s) \times s$. Since L and C are non-singular, A is of rank s iff the last $(n - s)$ rows and

columns of M are respectively linear combinations of the s first ones. This relies on the following identities:

$$\begin{cases} [M_{21} \;\; M_{22}] = X[M_{11} \;\; M_{12}] \\ \begin{bmatrix} M_{12} \\ M_{22} \end{bmatrix} = \begin{bmatrix} M_{11} \\ M_{12} \end{bmatrix} Y \end{cases} \tag{1.38}$$

Assuming a Las Vegas algorithm to compute M_{11}^{-1} with parallel cost $O_a(\log^2 n, n^\omega \log n)$ ([33], those identities can be verified with a parallel cost:

$$O_a(\log^2 n, n^\omega \log n). \tag{1.39}$$

This results in an optimal randomized Las Vegas algorithm to compute the rank.

In the above algorithm, randomization is strongly used for preconditionning the input (computation on LAC instead of A) in order to suppress data dependencies that bounds parallelism. A natural question is then the existence of a fast deterministic algorithm, i.e. with few dependencies. In [44], Mulmuley provided such a deterministic algorithm for computing the rank: it achieves parallel time $O(\log^2 n)$ but polynomial inefficiency. Then, randomization is required to provide efficiency.

1.4.2 Randomization to provide efficiency

Based on a generalization of a method developed in [27] for arbitrary fields, Mulmuley algorithm [44] reduces the problem of computing the rank to the computation of a characteristic polynomial in an extension of the ground field K.
In the following, A is assumed symmetric; this is done without loss of generality since

$$\text{rank}(A) = \frac{1}{2}\text{rank}\left(\begin{bmatrix} 0 & A \\ A^t & 0 \end{bmatrix}\right).$$

Theorem 8 *[44] Let A be a square symmetric matrix over a field K and let m be the highest integer such that x^m divides the characteristic polynomial $\xi_{A_Z}(x) = \sum_{i=0}^{n} a_i(z)x^i$ of the matrix A_Z over $K(z)$:*

$$A_Z = \begin{bmatrix} 1 & & & \\ & z & & 0 \\ & & \ddots & \\ 0 & & & z^{n-1} \end{bmatrix} A.$$

Then rank$(A) = n - m$.

Deterministic parallel algorithms for computing the characteristic polynomial in parallel time $O(\log^2 n)$ are known [16, 11] but they have work $O(n^{\omega+1})$. Even if we assume an optimal algorithm for computing the characteristic polynomial with arithmetic work $O(n^\omega)$, due to polynomial arithmetic, the cost of the above algorithm would be:

$$O_a(\log^2 n, n^\omega n log^{O(1)} n) \tag{1.40}$$

Since $a_i(z)$ are polynomials of degree $O(n)$, a way to obtain efficiency is to get rid off polynomial arithmetic on K using evaluation at a random value.

Moreover, efficient $O_a(log^2 n, n^\omega \log n)$ randomized algorithm are known for computing the minimal polynomial. Multiplying A_Z by a random non-singular matrix over K results, with high probability, in a matrix with distinct eigenvalues; then, minimal and characteristic polynomial are equal.
Those two steps of randomization result in the following efficient Monte-Carlo algorithm for computing the rank:

1. Choose a random non-singular matrix P;
2. Choose a random value z in K (or in an extension if K is too small);
3. Compute the minimal polynomial $\xi_{PA_z}(x)$ of the matrix PA_z;
4. Return $n - m$ where m is the highest integer such that x^m divides $\xi_{PA_z}(x)$.

The parallel cost is then:

$$O_a(\log^2 n, n^\omega \log n) \tag{1.41}$$

which results also in an efficient Monte-Carlo algorithm.

Remark. The above algorithm is very close to the one presented in 1.4.1; Mulmuley algorithm can effectively be considered as an inefficient deterministic version of 1.4.1. This is not surprising since both randomized algorithms solve efficiently the same problem. However, we have pointed out two different motivations for the use of randomization.

1.4.3 Conclusion

In the above examples, randomization is used to provide work-optimal computations from either slow or fast but not efficient deterministic algorithms. Due to the fact that only randomized algorithms are known for computing efficiently the inverse of a matrix in polylogarithmic time [33], randomization is an important tool in parallel computer algebra.

1.5 Parallel time complexity and NC Classification

An efficient parallel algorithm achieves polynomial speed-up within an optimal (or near optimal) number of operations. Obtaining bounds on the parallel time required to solve a given problem within a reasonable number of operations is then of fundamental interest. Moreover, as detailed in previous sections, very fast parallel but inefficient algorithms may be of practical interest if they can be coupled to an efficient but slow algorithm.

In the framework of parallel complexity, NC class [13] which includes polynomial sequential time problems that have a polylogarithmic parallel time plays an important role [35]. The parallel model used in the formal definition of NC is log-uniform family of boolean circuits [53]. NC^k is the class of problems that can be solved by such a family with depth $O(\log^k n)$ and $n^{O(1)}$ boolean gates[7]. For instance, integer arithmetic ($+$, $-$, $\times$ and Euclidean division) lies in NC^1. Introduction

[7]Gates compute bounded fan-in boolean operations (*or*, *and* and *not*) and have unbounded fan-out [26]. Extensions to unbounded fan-in gates leads to class AC [29].

of gates that deliver in output a random bit allows to define corresponding randomized classes: RNC for Monte-Carlo circuits and ZNC for Las Vegas ones. Problems P-complete [28, 49, 35] are in NC only iff $NC = P$; among them, the *monotone circuit value problem* (MCVP) consists in the evaluation of a boolean circuit, roughly equivalent to a DFG with boolean nodes as defined in this chapter. The integer greatest common divisor remains an open question; only sub-linear $O(\frac{n}{\log n})$ algorithms are known [34, 35].

The algebraic extension [61] of this primitive model allows to build circuits which gates compute arithmetic operations in an algebraic domain. A gate testing nullity ($? = 0$) is introduced in order to mix boolean and arithmetic operations. For instance NC^k_F (F stands for *field*) is the class of problems that can be solved by log-uniform family of circuits whose gates perform arithmetic operations in any field, i.e. $+$, $-$, $\times$, $/$ and $? = 0$. Complexity of basic computer algebra problems has been extensively studied [8, 13, 59, 60, 35, 45]. Polynomial arithmetic ($+$, $-$, $\times$ and Euclidean division) lies in NC^1_F [45]. An important class is DET_F which contains problems NC^1-reducible to the determinant of a matrix; matrix powering is complete for DET_F. DET_F is included in NC^2_F. Most of linear algebra problems lie in NC^2_F: rank, null-space, minimal and characteristic polynomial, gcd of many polynomials [8, 44], Hermite normal form of polynomial matrices [31], Smith and symbolic Jordan forms [52, 58, 57, 21]. Note that those problems admit an optimal $O_c(\log^2 n, W_s(n))$ parallel algorithm by using randomization [33, 23, 24, 45]. Though, in certain cases, some general techniques are known to remove randomness without increasing the work [42], no work optimal deterministic algorithms with poly-logarithmic time are known for those problems.

As it appears for most computer algebra problems studied in this chapter, parallel algorithms often appear as a restructuration of sequential ones, taking into account algebraic properties of the arithmetic operations involved. Although evaluation of a boolean circuit is P-complete, several algorithms have been developed to evaluate arithmetic DFGs (also called straight-line programs) taking benefit of the underlying structure. In a semi-ring, DFG that are trees can be evaluated in $O(\log n)$ time without increasing the number of operations performed [9]. Any DFG performing n operations in a semi-ring and whose outputs are of arithmetic degree[8] d can be evaluated in $O_a(\log n \log(nd), n^3)$ [32]. This result has been extended to DFGs performing operations in a lattice [51]. A more general simulation of a RAM machine on a PRAM one [43] shows that any DFG can be evaluate in parallel on an unbounded number of processors with polynomial speed-up.

1.6 Conclusion

This chapter overviews the PRAM framework (execution model and main algorithmic techniques) in which parallel algorithms are built and analyzed. The macro data-flow graph (DFG) related to the execution plays a central role: it describes data-dependencies between blocks of instructions.

Abstract measures used to analyze algorithms are *depth* and *work*; *arithmetic* and *communication* costs are distinguished. The one corresponds to operations performed (macro-instructions nodes) while the other to access in the shared memory (data dependencies nodes). Arithmetic work and depth have been used for many years to analyze performances of parallel algorithms

[8] In such a DFG, any output may be equivalently seen as a polynomial whose indeterminates are the inputs. The arithmetic degree is then the maximal degree of polynomials corresponding to the outputs.

[9, 55, 35, 28, 6]. Due to experimental constraints, the relevance of communications costs (i.e. total communication traffic – work - and total communications delay) has been pointed out to obtain practical performant programs [5, 19]. Since minimizing communications overhead and minimizing parallel time are antagonist, good trade-offs have been studied for several common algorithms [47, 1, 48]. The granularity, defined as the arithmetic-to-communication works ratio, appears as a good parameter.

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Chapter 2

Programming models and scheduling

Contents

2.1 Asynchronous distributed architectures 35
2.1.1 Realistic models of distributed architectures 35
2.1.2 Basic programming tools 37
2.1.3 Shared virtual memory 38
2.2 How to schedule a DFG 39
2.2.1 Scheduling cost of a DFG 39
2.2.2 Off-line and on-line scheduling 41
2.2.3 Which scheduling algorithms in computer algebra ? 42
2.3 On-line scheduling algorithms 43
2.3.1 Foundations of on-line scheduling 43
2.3.2 Lower bounds for competitive ratio 44
2.3.3 Communications and scheduling overheads 45
2.3.4 Athapascan: a simulation of the ATH PRAM language 49
2.3.5 The ATHAPASCAN programming model 49
2.3.6 Execution model of ATHAPASCAN 49
2.3.7 An example of ATHAPASCAN program 50
2.4 Conclusion 50

In order to analyze performance of algorithms, a formal model is needed to take the costs into account. The success of the PRAM model is mainly due to the fact that it does not attempt to represent any parallel architecture but can be mapped onto various ones. Moreover, the simulation on a realistic machine can be made efficient (up to a constant related to the granularity), provided many processors of the PRAM are mapped onto a single processor of a host machine. This success is brought to evidence by the fact that most of the tricks used to optimize practical performances when programming on a given architecture are relevant to algorithmic techniques that are theoretically justified on the PRAM model.

Given an algorithm (let us say a macro data-flow graph – DFG – as presented in chapter 1) and a particular multiprocessor architecture, the problem then is reduced to:

- find a good (the best) schedule of the DFG;
- implement the resulting algorithm in a programming language.

Only now, the performance of the program, i.e. the completion time of an execution, may be determined. Assuming fixed the initial algorithm, the machine and the input, this performance depends directly on the scheduling strategy. Tuning the program ammounts to improving the schedule it implements.

This chapter presents the main techniques used to schedule data-dependencies graph (DFG) on a given architecture. As presented in chapter 1, a DFG is the abstract representation of the execution of a particular program on a specific input data x. A fine grain description (elementary instruction, elementary data dependency) is unrealistic for executions requiring hours of computation time. We will thus assume that arithmetic nodes of the DFG correspond to sequence of instructions: each arithmetic node is then weighted by the number of elementary instructions it performs.

Arithmetic depth $T(x)$ and work $W_a(x)$ are evaluated taking into account nodes weights. $T(x)$ is a lower bound of the minimal time required by any schedule ignoring communications times. $W_a(x)$ is the exact number of operations required by a sequential execution of the algorithm. Since the best schedule may replicate some arithmetic nodes in order to minimize completion time, note that $W_a(x)$ is also a lower bound on the number of operations performed by any schedule.

Similarly, transition nodes may correspond to a complex data structure (not a single word); each transition node is weighted by the size of the data it corresponds to. Communication delay $C_d(x)$ and work $W_c(x)$ are also evaluated accordingly. Ignoring arithmetic time, $C_d(x)$ is an upper bound on the minimal communication time required by the best schedule for an infinite number of processors. $W_c(x)$ is an upper bound on the number of remote access (communications) performed by any schedule.

As straightened in the previous chapter, the initial parallel algorithm is assumed efficient, i.e. $W_a(n) = \Theta(W_s(n)$ where $W_s(n)$ is the time of the best known (uniform) sequential algorithm, n being the size of the input. Moreover, in order to make performance evaluation with x in input, we assume that there exists a constant K such that:

$$\forall x, |x| > n_0 : \quad W_a(x) < KW_s(x) \tag{2.1}$$

Note that, for a given input x, DFG_x may be known only after completion: instructions or transitions nodes and edges are dynamically built. In the language ATH introduced in chapter 1,

those nodes are created either by execution of a `fork` instruction or by access to a shared data. Similarly, the cost of any instruction node (resp. size of data related to any transition) is known only after completion of the instruction (resp. communication). In such a general context, DFG_x has to be scheduled using an on-line algorithm. Related to a functional programming model, most of computer algebra algorithms present such a dynamic behavior; we thus focus on on-line scheduling algorithms.

Organization of the chapter is as follows. In the first section, specific characteristics of asynchronous distributed architectures are recalled. Costs of basic operations are modeled by the *LogP* model introduced in [15]. Basic mechanisms allow parallel and distributed programming: communications, threads, remote memory access and synchronizations tools. In the second section, the scheduling of a PRAM algorithm on such a machine is discussed. Approaches may be distinguished in two classes. The first one [54, 28] is based on the simulation of a PRAM machine on a given architecture: the execution of the parallel algorithm is managed via the simulation. Global synchronization and emulation of the shared memory, which are at the basis of the PRAM model, are key points. The second one [26, 51, 38, 50, 5, 19] is based on the direct scheduling of the DFG. The execution of the algorithm is handled by a scheduling algorithm. Both approaches are motivated by the availability of provably good approximation algorithms to solve the underlying theoretical problems (permutation routing [48, 42, 55, 40] or DAG off-line and on-line scheduling [29, 49, 13, 36, 47, 14, 8, 6, 30]).
The last section focuses on on-line scheduling algorithms which are of main interest in computer algebra. We firstly recall upper and lower bounds on the competitive-ratio without taking into account scheduling and communication overheads. As a corollary, we then exhibit a list-scheduling algorithm which achieves optimal simulation of any efficient PRAM algorithm, taking into account those overheads. Finally, we overview some programming languages or libraries based on those approaches, focusing on the one suited to computer algebra algorithms. We describe an effective implementation of the theoretical language ATH introduced in chapter 1, ATHAPASCAN, which achieves provably performances.

2.1 Asynchronous distributed architectures

2.1.1 Realistic models of distributed architectures

There is an apparent convergence in the field of distributed architectures which are similar to a network of workstations. A parallel machine consists in a set of independent processors, each with considerable local memory, linked by an interconnection network. Fundamental differences with the PRAM model are the following (compare 2.1 to 1.1 in 1):

- **asynchrony**: each processor works independently with its own local memory; there are no global synchronization.
- **contention**: the network is a resource with bounded access.

Like the local PRAM introduced in chapter 1, two levels of access may then be distinguished: local and remote access (parallel machines are often called NUMA for non-uniform memory access[1]).

[1]Note that this non-uniformity appears also at the processor level between cache and RAM access.

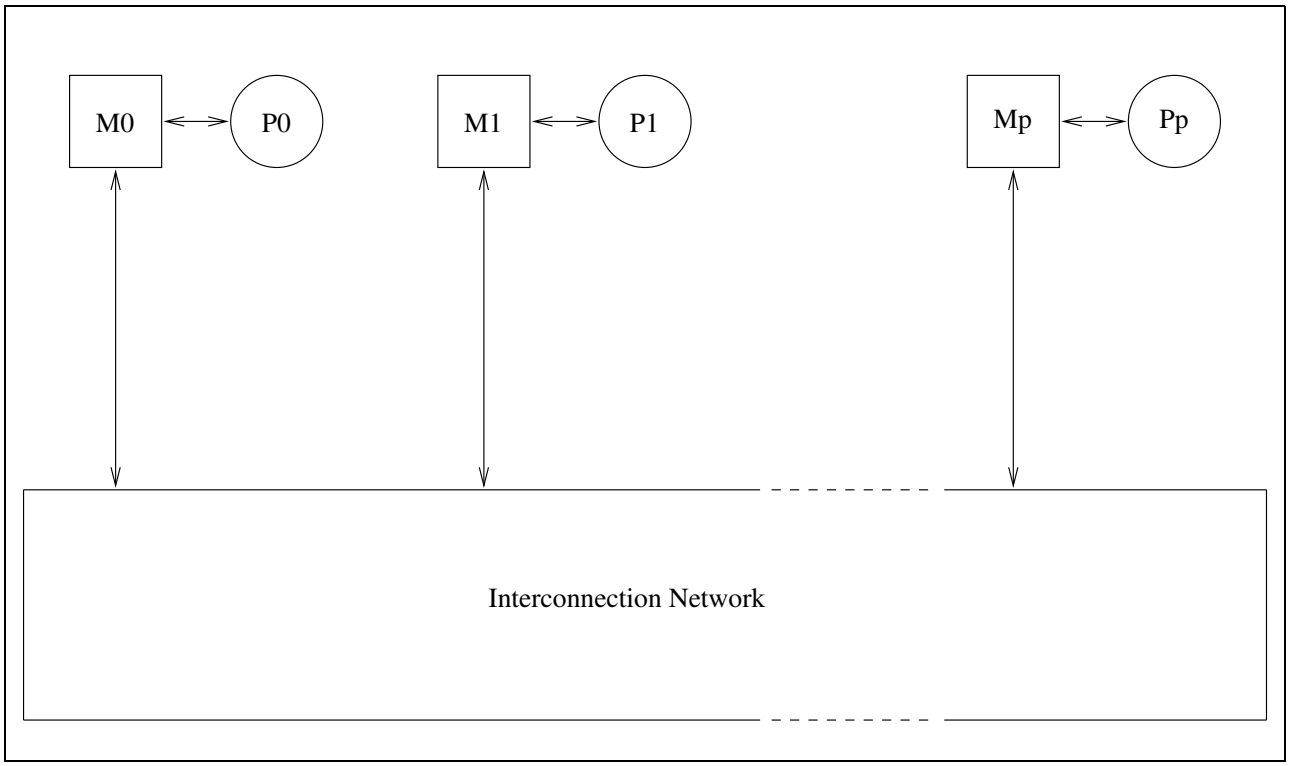

Figure 2.1: General structure of a distributed architecture. *Differences with the PRAM presented in chapter 1 are the absence of a global sequencer and contention for access to the network.*

Costs of remote access are mainly characterized by two factors:

- **bandwidth**: the rate at which each processor can access memory;
- **latency**: the time between making a remote access request and receiving the reply. Latency accounts for resource allocation (solving contention on network) and duration of communication (related to physical distance).

The network bandwidth that is available on recent parallel computers ($>$ 1 GB/s on SGI Power Challenge, Cray T3E, SUN HPC) and even on local networks (typically 1 Gb/s using Mirynet connection or DEC Memory Channel) is becoming large enough compared to the bandwidth to local memory; thus it appears less and less as a bottleneck. However, latency is a more serious problem since it is bounded by physical limits.

Several variations of the PRAM model have been proposed in order to take into account those practical constraints [15]: memory contention [40, 54, 42, 45], asynchrony [27], memory hierarchy [3, 34], latency and bandwidth [47, 1]. Considering that point-to-point communication is a basic primitive, the model *LogP* proposed in [16] characterizes a distributed architecture by the following parameters (fig. 2.2):

L : *latency*: an upper bound on the delay incurred in communicating an unit size data (i.e. a small number of words) from its source to its destination; an extension to longer messages has also been developed [2].

o : *overhead*: the time a processor is engaged in the transmission or reception of a message;

g : *gap*: minimum time interval between consecutive message transmissions or receptions. The reciprocal of g corresponds to the available communication bandwidth per processor; it is denoted σ in [47].

P : the number of *processors*.

This model has been successfully used on different architectures to predict the execution time of some parallel algorithms [16, 20]. As a consequence, classical balanced tree schemes used on the PRAM to perform iterated sum or broadcast appear as non optimal [41].

As a conclusion, the portability of a parallel program cannot be achieved if the characteristics of the target architecture are not taken into account. Notingly, the communication parameters, that are partly modeled by *LogP*, have significant influence on the performances.

2.1.2 Basic programming tools

Reliable message-passing communication is the lowest-level feature required for programming a distributed architecture. It allows both to exchange data between processors (the basic functionality of the PRAM shared-memory) and to express synchronization (the functionality ensured by the sequencer of the PRAM).

Since 10 years, several message basic interfaces have been built on top of the low level ones provided on any specific architectures in order to allow portable programming. Most famous ones are PVM [24] and MPI [53]. MPI has been standardized [18] and is nowadays available

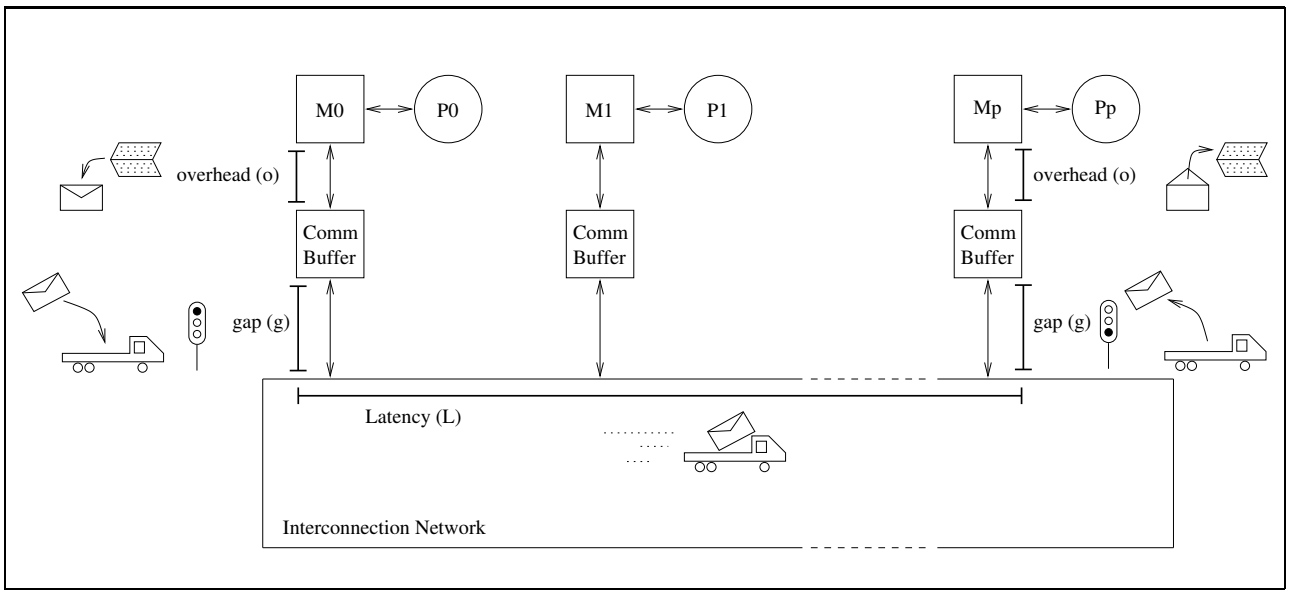

Figure 2.2: Communication cost parameters in the LogP model.

on any distributed architecture or network of workstations. Basic features of MPI are point-to-point and (blocking) collective communications, communication contexts (*communicators*), user-defined data-types. Other extensions concern remote memory access, parallel input and output (MPI-F), active messages and dynamic process control.

In order to hide the communication latency by arithmetic computations, two tools may be used: asynchronous communications and threads. Threads are lightweight processes which require a small overhead for context switching. They are handled directly in the source program: a standard interface, POSIX, has been defined [10]. Threads have firstly been defined for concurrent programming and efficient use of SMPs (Shared Memory Processor) on a single node. Since threads access concurrently the same memory space, synchronization tools are provided for atomicity, such as locks and semaphores (sometimes monitors).

Threads are well suited to hide latency on a distributed architecture: when a thread waits for the result of a communication, it may be preempted and a ready one scheduled. Thus, several portable programming interfaces have been built to couple a message-passing library (usually not thread-safe) and a thread library (available on a single node), providing an easy way to the user for lightweight remote procedure calls or active messages [21, 46, 9].

2.1.3 Shared virtual memory

On many distributed architectures, remote memory access are possible: they provide a virtual shared memory analogous to the one of the PRAM. On such machines, specific hardware allows to load transparently a local or remote data in the cache of a processor. In order to hide the latency of remote access, prefetching and multi-threading is used.

The simulations of the PRAM shared memory on a distributed architecture use hash functions (randomly chosen from a universal class) to map shared memory cells onto the ones of the ar-

chitecture (i.e. memory modules) [48, 42]. The delay of a simulation is the time required for a single access. It is related to the evaluation of the hash function, the memory contention (when several access to a same module occur), and the routing time if the network is not complete. In [48], a simulation with delay $\Theta(\log p)$ of an EREW PRAM on a butterfly network is given. In [40], randomized simulations of EREW and CRCW PRAMs on a distributed architecture with a complete interconnection network (contention is not taken into account) are presented with delay $O(\log \log p \log^\star p)$. Note that, concerning the CRCW PRAM, this simulation is at a factor $\log^\star p$ from optimal.

In order to obtain optimal simulations, such delays are to be hidden by arithmetic computations. The key idea is parallel slackness [42, 55, 40]: it consists in simulating a PRAM with n processors on a distributed architecture with fewer processors $p < n$. The simulation is optimal (*time-processor optimal*) if the delay for an access is proportional to n/p. For instance, the previous mentioned simulation [40] leads to time-processor optimal simulation of an EREW PRAM with $n = p \log \log p \log^\star p$ processors on a distributed architecture with less than p processors. Note that parallel slackness is also involved when using asynchronous communications and threads to hide latency.

On the contrary of communications, remote access to shared memory do not basically allow to synchronize computations. In the PRAM, such a synchronization mechanism is provided by the global sequencer. On distributed architectures, intrinsically asynchronous, synchronization tools classically used are communications, locks and semaphores.

2.2 How to schedule a DFG

Being given an algorithm, the problem considered here is to schedule the DFG related to the execution on input data on a distributed architecture. The goal is to obtain an optimal schedule related to the DFG.

2.2.1 Scheduling cost of a DFG

Computing such an optimal schedule is a difficult problem. Even if communication costs are ignored and the DFG fixed (i.e. no dynamic task creation) with tasks of known duration, computing an optimal schedule is NP-complete and deciding whether the length of the optimal schedule is a given integer l is co-NP-complete [23]. However, on machines with p identical processors, there are several polynomial algorithms with bounded competitive ratio, the most famous one being list-scheduling [29]. Moreover, even on non-uniform machines, approximation algorithms are known [52, 30].

Computing a schedule implies an overhead in the execution time; this scheduling overhead is governed by the time required to compute the schedule itself (i.e. the cost of the scheduling algorithm) and to realize this schedule (i.e. the mapping of tasks, preemption, migration). The scheduling overhead is included in the execution time $T_p(x)$ of the algorithm with input x on the target machine.

Definition 5 (Notation) *Being given a scheduling algorithm s, the execution time of an algorithm with input x on a machine with p identical processors using the schedule delivered by s is denoted*

$T_p^{(s)}(x)$.
The minimum execution time over all scheduling algorithms s is denoted $T_p^\star(x)$

When there is no confusion about s, $T_p^{(s)}(x)$ is denoted by $T_p(x)$.

The cost of computing a schedule is directly related to the size of the DFG, i.e. the number of tasks and dependencies it contains. Note that those costs are different from the arithmetic and communication works considered in the previous chapter which take into account the number of operations performed in each task and the number of communications related to each data dependency (transition).

Definition 6 *Let DFG(x) be the macro data-flow graph corresponding to the execution of a parallel algorithm on an unbounded number of processors. We define the following measures:*

- *$N_a(x)$ is the number of task nodes in DFG$_a(x)$;*
- *$N_t(x)$ is the number of transition nodes in DFG$_t(x)$;*
- *$N_d(x)$ is the maximal degree of a task node in $DFG(x)$; the degree is the number of input and output edges on a task node (to or from a transition node).*

The scheduling cost S_x of DFG(x) is:

$$S(x) = (N_d(x), N_a(x) + N_t(x))$$

Note that other measures may be considered in the analysis of a scheduling algorithm. For instance, other parameters considered in [7] are the maximum number of edges between any pair of nodes and the width of $DFG_a(x)$, i.e. the maximum number of tasks that may be executed concurrently.

The finer the DFG, the larger its scheduling cost is and thus the more expensive the computation of its schedule will be. Similarly to granularity, the regularity ρ is defined as the ratio of the arithmetic work to the size of the DFG.

Definition 7 *The regularity $\rho(x)$ is defined by:*

$$\rho(x) = \frac{W_a(x)}{N_a(x) + N_t(x)}.$$

A PRAM algorithm (or equivalently its related DFGs) if said of polynomial regularity *iff:*

$$\rho(x) = |x|^{1+\epsilon} \quad \textit{with} \quad \epsilon > 0.$$

Notation. In the following, we will consider the execution of a given algorithm on a given p processors machine with an arbitrary input x of size n. Thus, all notations are implicitly related to x and n. For instance, N_a will denote $N_a(x)$, the number of task nodes in the macro data-flow graph related to the execution on an unbounded number of processors with x in input.

2.2.2 Off-line and on-line scheduling

The DFG corresponding to the execution may be partially determined at compile time by data flow analysis of the code of the algorithm, or may be discovered during the execution (depending on the value of computed data) and completely known only after the end of the execution. Depending on this knowledge of the DFG, the scheduling can be then computed off-line or on-line.

Static allocation of tasks to processors.

When the DFG corresponding to the execution can be analyzed at compile-time, it is possible to find a good schedule by hand, may be using static scheduling tools. The result of the scheduling is to assign each task of the DFG to one processor (or more if replication is required). On a given processor, tasks are sequentially ordered[2] in order to respect precedences; data dependencies between them are emulated by access to shared data in the local memory. When tasks are placed on different processors, data-dependencies (i.e. access to data and precedence relations) may be emulated in two different ways:

- By communication. The data corresponding to a write-read dependency has then to be explicitly sent from the writing task to the reading one. This operation corresponds to a physical global copy of the data; locally unreferenced data have to be deleted (local garbage-collection).
  An important point is that the completion of receiving instructions implicitly implements the precedence relation (synchronization).

- By shared-memory access. Communications that implement remote access are then implicit. However, the precedence relation between non local tasks has to be described using global synchronization tools.

As a result, before execution, each processor gets its own program. Usually, this program is the same for all the processors but is parameterized by the pid of the executing processor in order to implement different behaviors. PYRROS uses this approach and a specific scheduling algorithm which performs a clustering of tasks [26, 25].

Dynamic allocation of tasks to processors.

In computer algebra, elementary tasks are often of unknown cost. For instance, costs of arithmetic operations (on rationals, polynomials or matrices) are usually unknown at compile time since their are related to characteristics of the values computed at execution time (size of the data, degree of a polynomial, sparsity of a matrix). Depending on such values, parallelism (i.e. creation of a task) may be generated during the execution. In such a case, an on-line scheduling algorithm is used.

Most of on-line scheduling algorithms are based on the following greedy scheme called *list-scheduling* [4, 11]:

- When a processor creates a new task (`fork` instruction of the PRAM language ATH), it stores it in a list of tasks.

[2] Multi-threading may be used to describe a partial execution order.

Note that, there may exist *ready* tasks, i.e. whose precedence relations are satisfied, and non-ready tasks, i.e. whose one of the precedent tasks is not completed.

- When a processor becomes *idle* (i.e it has no ready task to execute), it gets a ready task in the list if any.

Algorithms vary depending on the way the list is managed and processors put and get tasks in it.

The program that implements the algorithm expresses a *functional parallelism*: tasks generally correspond to procedure or function calls. Non-completed tasks or data are called *future*. An important point concerns the management of data, parameters of the task: they can be systematically copied in a stack corresponding to the function call or passed by a reference to data in the shared memory. Precedence relations between tasks may correspond either to data dependencies or to task precedences.

Scheduling operations

The previous section does not specify which instructions a scheduling can perform, except classical computations and the possibility of executing a basic task – an elementary node in the DFG – on a processor. Migration instructions allow to suspend a task during its execution in order to map it, maybe later, on another processor [52, 4]:

- *migration restricted to restart:* when a task is moved to another processor, its execution restarts from its beginning;
- *migration:* when a task is migrated to another processor, its execution restarts from its last instruction performed.

A scheduling algorithm with *no-preemption* makes no use of those operations: it has no control on a task once it has assigned it to a processor, just getting information when the task is finished. Migration restricted to restart, denoted in [52] as *no-preemption with restarts*, is useful on machines whose processors are not identical.

2.2.3 Which scheduling algorithms in computer algebra ?

An important point is that on-line and off-line scheduling algorithms have theoretical foundations [29, 22, 35, 11, 30]. There exist provably good approximation algorithms for both with bounded competitive ratio. For both, specific algorithms are developed to increase performances for certain classes of graphs (for instance trees or SP11 graphs – fork-join –).

Of course, performances of off-line algorithms are better when the DFG is known and the machine fixed. However, since on-line algorithms make no hypothesis on the execution (or few for the determination of tasks precedences), they can be used for any class of applications and thus are of general interest.

Thus, both techniques are used in computer algebra. For instance, block-scattering matrix mapping, which can be considered as an hand-made off-line algorithm, leads to near-optimal performances for linear algebra problems like dense matrix multiplication or inversion (cf chapter 1) over a small finite field (e.g. GF(2)) on a distributed architecture with identical processors.

However, due to their generality and their close relation with functional parallelism [31, 51], on-line scheduling algorithms are of specific interest for a parallel computer algebra system. In the following we thus focus of those algorithms.

2.3 On-line scheduling algorithms

2.3.1 Foundations of on-line scheduling

Theoretical foundation of on-line scheduling algorithms is due to Graham [29]. The following theorem appears has an arbitrary grain version of Brent's principle presented in chapter 1. We recall its proof which is the basis of most of further results.

Theorem 9 *[29] If scheduling overhead (i.e. the cost of computing the schedule and managing the list of tasks) and communication costs are not considered, any list-scheduling algorithm has competitive ratio* $\left(2-\frac{1}{p}\right)$*, i.e.*

$$T_p \leq \left(2-\frac{1}{p}\right) T_p^\star$$

.

A list-scheduling algorithm is such that, at any time, at least one processor is executing a task. Then, if at a given time a processor is idle then there exists at least one processor which executes a task. Let t_{j_1} be one of the tasks completed at date T_p and let d_{j_1} be the date when execution of t_{j_1} has been started. Two cases arise:

1. either no processor was idle before d_{j_1}.
2. either there was at least one processor idle at a certain date before d_{j_1}. Let θ be the latest date before d_{j_1} when a processor was idle. At θ, t_{j_1} was not ready (else it would have been started on an idle processor). Thus, there exists a task t_{j_2} such that t_{j_2} was being executed at θ and $t_{j_2} \prec t_{j_1}$. Let d_{j_2} be the date when execution of t_{j_2} has been started.

Recursively applying this scheme until case 1 occurs, we build a sequence of tasks $t_{j_k} \prec \ldots \prec t_{j_2} \prec t_{j_1}$ such that, at any time where a processor is idle, there exist $1 \leq i \leq k$ such that t_{j_i} is being executed on one processor.

Similarly to chapter 1, let T be the minimal arithmetic time on an unbounded number of processors and W_a be the total number of operations. The total idle time is defined by $\#I = pT_p - W_a$. For $1 \leq i \leq k$, let l_i be the duration of task t_{j_i}. We have: $\#I \leq (p-1)\sum_{i=1}^{k} l_i$ which leads to:

$$pT_p \leq W_a + (p-1)\sum_{i=1}^{k} l_i.$$

Besides, since tasks t_{j_i}, $1 \leq i \leq k$ are on a critical path: $\sum_{j=1}^{k} l_i \leq T$. This leads to:

$$T_p \leq \frac{W_a}{p} + \left(1-\frac{1}{p}\right) T \tag{2.2}$$

We also have $T \leq T_p^\star$. Moreover, since W_a operations are to be executed in any schedule, $W_a \leq pT_p^\star$. Replacing in 2.2, we obtain: $T_p \leq \left(2 - \frac{1}{p}\right) T_p^\star$. □

As a corollary, we obtain the following constructive version of the simulation of a PRAM with an unbounded number of processors on one with p. Note that tasks is the DFG are of arbitrary durations; the only restriction which is respected in the DFG representation is that once a task is ready, it can be executed sequentially with no interruption due to synchronization.

Theorem 10 *Let $\mathcal{A}$ be an ATH PRAM program that run in (arithmetic) parallel time T and work W_a on a given PRAM with an unbounded number of processors. Then $\mathcal{A}$ can be executed by an on-line list scheduling to run in (arithmetic) parallel time T_p:*

$$Max\left\{\left\lceil\frac{W_a}{p}\right\rceil, T\right\} \leq T_p \leq \left\lceil\frac{W_a}{p} + \left(1 - \frac{1}{p}\right)T\right\rceil \tag{2.3}$$

The proof is direct from 2.2. □

Theorem 9 is stated in a restricted version [4]. In fact the bounds 2.2 holds even if the precedence relation $\prec$ considered by the list scheduling algorithm is weaker than the one $\prec'$ considered for defining the optimal schedule. The proof is direct since we will also have $t_{j_k} \prec' \ldots \prec' t_{j_2} \prec' t_{j_1}$. Clearly, the same remark holds if duration of tasks is increased.
This implies that neither adding precedence constraints such as synchronization barriers to obtain a well structured DFG nor inserting artificially null operations in order to have all tasks of the same length help any on-line algorithm.

Remark. This theorem generalizes Brent's principle (theorem 1 in chapter 1) to arbitrary DFGs, i.e. any ATH program where tasks are generated dynamically with arbitrary shared-data dependencies and are of unknown durations.

2.3.2 Lower bounds for competitive ratio

A natural question is then to determine if it is possible to have a better competitive ratio than $\left(2 - \frac{1}{p}\right)$, either on the same model or by considering larger classes of scheduling algorithms.
This problem has been studied in [52], in which the following proposition is proved.

Theorem 11 *[52] On the p-PRAM, the competitive ratio is lower bounded by $\left(2 - \frac{1}{p}\right)$ for any scheduling algorithm of the following classes:*

1. *Deterministic with no preemption,*
2. *Deterministic with migration;*

and is lower bounded by $\left(2 - \frac{1}{\sqrt{p}}\right)$ for any randomized scheduling with no preemption.

We only sketch the proof for the first case. The complete proof for this theorem is given in [52].

The adversary builds the following DFG instance G due to Graham [29]. G contains $1+p(p-1)$ independent tasks. One task α_1 is of length p, while other tasks β_k, $1 \leq k \leq p(p-1)$ are of length 1.

The optimal schedule is of length p. It executes the task α_1 on a given processor, and the $p(p-1)$ unit tasks β_k on the $p-1$ remaining processors.

The length of any schedule of G is equal to $p+t$, where t is the time when the task α_1 starts its execution. Since the tasks durations are unknown for the scheduling algorithm, the adversary strategy will thus consist in making t as large as possible.
The tasks that are processed first are then the $p(p-1)$ unit time tasks β_k, that are executed in $p-1$ time units with no idle time. Then, at time $t=p-1$, the task α_1 starts its execution. The length of the obtained schedule is then $2p-1$, which provides the desired lower bound. □.

As a consequence, neither preemption nor randomization can improve consequently performances compared to list-scheduling.

In order to increase the competitive ratio, it is then required to use additional informations on the DFG such as its shape or duration of its tasks.

For instance, we consider the case where all tasks are independent and sorted according to their durations; note that only the ordering is known but not durations. In this case, the on-line *LPT* list-scheduling algorithm that assigns the task of maximal duration when a processor becomes idle has competitive-ratio [29][12]:

$$\text{Min}\left\{\left(\frac{4}{3}-\frac{1}{3p}\right),\left(1+\frac{1}{N_a}\frac{p-1}{p}\right)\right\} \tag{2.4}$$

Note that if no information is given on the durations tasks, then the fact that they are independent is of no help to decrease the competitive ratio $\left(2-\frac{1}{p}\right)$ (cf the adversary considered in the proof of theorem 11).

Remark. List scheduling algorithms are involved as a basic level in on-line approximation algorithms used for other kind of machines such as [52, 30]:

- uniform machines: processors speeds are constant and differ each one from a constant unknown factor;
- non-uniform machines: there are no relation between processors speeds; the duration of a task varies depending on the processor which executes it.

In this case, at least migration restricted to restart is required in order to guarantee a competitive ratio [52].

2.3.3 Communications and scheduling overheads

Previous theorems do not take into account neither the cost of tasks allocation (i.e. scheduling overhead) neither communications required for access in shared memory.

Several authors have considered the theoretical influence of those overheads on list scheduling algorithms in order to provide provably optimal on-line scheduling algorithms. In [13], Cole and Vishkin give an algorithm to schedule n independent tasks optimally on a PRAM with $\frac{n}{\log n}$ processors; this algorithm is used to implement the first optimal algorithm for list-ranking [37, 39]. In [6], Blelloch, Gibbons and Matias study the scheduling of nested fine grain computations, implemented in the language NESL [5]. Blumofe and Leiserson give an optimal list-scheduling algorithm for

strict multi-threaded computations[3] [8, 7], based on randomized work-stealing; this algorithm is in the kernel of the Cilk language [38]. Any of those scheduling algorithms restricts to a shape of DFG and do not take into account contention problems.

In this section, we give a near optimal scheduling algorithm for any DFG shape but with restrictions on the arithmetic and communication costs. We prove that, if input size is large enough, efficient and coarse-granularity PRAM algorithms[4] can near-optimally be scheduled on a distributed architecture.

We assume that the target machine is a distributed architecture with p identical processors. In order to take into account communication costs and contention, we refer to the LogP model (cf section 2.1.1). The duration between the sending and the reception of a small message (i.e. one word) is bounded by $\sigma = 2g + 2o + L$.

Furthermore, we assume that a shared memory is simulated on the architecture with the help of hashing functions (see section 2.1.3); the delay occurring for any access in the shared memory is bounded by h. Note that h is related to the number of processors if no slackness is used.

Like in chapter 1, let C_d and W_c denote respectively the communication delay and work involved by the algorithm.

Theorem 12 *Let $\mathcal{A}$ be an ATH PRAM program that has parallel arithmetic cost (T, W_a), communication cost (C_d, W_c) and scheduling cost (N_d, N_a). Then $\mathcal{A}$ can be executed to run in parallel time T_p (including scheduling and communication overheads):*

$$T_p \leq \frac{W_a + hW_c}{p-1} + \left(1 - \frac{1}{p-1}\right)(T + C_d) + 4\sigma N_d N_a \tag{2.5}$$

The proof is based on an adaptation of the scheme used in theorem 9.
We consider here an implementation of a list scheduling on $p-1$ processors, indexed $p_1, \ldots p_{p-1}$. The last processor, p_0, handles the list of tasks and assigns tasks to other processors.

For the sake of simplicity, we restrict the proof to the case where any shared variable is written only once and then read only once; once read access have been completed, the space related to the shared data is garbaged. This corresponds to the case of an EREW program with single-assignment variables.

When a processor p_i completes the execution of a task, it sends a message to p_0 and waits for receiving a new ready task to perform from p_0.
When a processor p_i creates a new task (`fork` instruction) it asynchronously sends to p_0 a message of size bounded by N_d that defines all data dependencies of the new task (i.e. the shared data that it will read before its execution or write after its completion).

Processor p_0 manages a list Q of ready tasks and a list I of idle processors. For this purpose, it uses two arrays: one, A, stores the task nodes created and not completed; the other, B, the descriptors of the allocated shared variables. Any descriptor in B points to the task that requires the corresponding shared data in reading. Pointers from B to A are updated at task creation and task completion. When a task in A is no more pointed to by any element in B, it is put in Q. The cost of this arithmetic computation on p_0 is proportional to N_d but independent from p and σ: we neglect it compared to $(p-1)\sigma N_a N_d$.

[3] There is always a dependency between a thread and one of its ancestor and access to shared data are not considered.
[4] i.e. with polynomial speed-up, constant inefficiency and $W_c(n) = W_a^{\epsilon}(n)$ with $\epsilon < 1$ (cf chapter 1).

When p_0 receives a message of task completion, it first updates A and B, putting new ready tasks if any in Q. It puts the processor in I. Then, while there are ready tasks in Q and idle processors in I, it gets a task from Q and a processor from I, removes them from the lists and asynchronously sends a message assigning the task to the processor; the length of the message is at most N_d. For the whole execution, the computation time on p_0 needed for the management of those lists is proportional to N_a and independent from p, N_d and σ: we also neglect it compared to $(p-1)\sigma N_a N_d$.

Note that due to contention, a processor which is idle may wait at most $(p-1)\sigma N_d$ after p_0 has assigned a new task to it and before it receives it. Conversely, when a processor completes a task, processor P_0 receives the corresponding message at most $(p-1)\sigma N_d$ tops after.

Moreover, let l_i be the length of the task t_i, $1 \leq i \leq N_a$; let c_i be the number of – unit size – shared data handled by t_i (i.e. read or written during its execution). From the point of view of p_0, a processor p_i is said *idle* when it is in the list I. Thus, the processor executing t_i is considered as active (i.e. *not idle*) when it is not in the list I, i.e. from the moment p_0 has sent t_i to it and until p_0 receives the corresponding task completion message: this duration is bounded by $l_i + hc_i + 2N_d(p-1)\sigma$.

Let $\#I$ be the total idle time seen from p_0 on processors $p_1, \ldots, p_{p-1}$. Let T_p be the length of the schedule; we have:

$$(p-1)T_p \leq \#I + \sum_{i=1}^{N_a}(l_i + hc_i + 2N_d(p-1)\sigma) \tag{2.6}$$

We now follow the scheme of theorem 9. Let t_{j_1} be the last task completion message received by p_0 at date T_p and let d_{j_1} be the date when p_0 has assigned t_{j_1}. Two cases arise:

1. either no processor was idle for p_0 before d_{j_1}.
2. or there was at least one processor idle for p_0 at a certain date before d_{j_1}. Let θ be the latest date before d_{j_1} when a processor was idle. At θ, t_{j_1} was not ready (else it would have been assigned on an idle processor). Thus, there exists a task t_{j_2}, $t_{j_2} \prec t_{j_1}$, such that t_{j_2} has been assigned by p_0 before θ and whose completion message has been received by p_0 after θ. Let d_{j_2} be the date when p_0 has assigned t_{j_2}.

Recursively applying this scheme until case 1 occurs, we build a sequence of tasks $t_{j_k} \prec \ldots \prec t_{j_2} \prec t_{j_1}$ such that, at any time when a processor is idle, there exists $1 \leq i \leq k$ such that p_0 has assigned t_{j_i} to a processor and has not received the corresponding completion message yet. We thus have: $\#I \leq (p-2)\sum_{i=1}^{k}(l_{j_i} + hc_{j_i} + 2N_d(p-1)\sigma)$. Besides, since tasks t_{j_i}, $1 \leq i \leq k$ are on a critical path: $\sum_{j=1}^{k} l_{j_i} \leq T$ and $\sum_{j=1}^{k} c_{j_i} \leq C_d$ which leads to:

$$\#I \leq (p-2)(T + hC_d + 2N_a N_d(p-1)\sigma) \tag{2.7}$$

where T denotes the minimal arithmetic time on an unbounded number of processors.

Let $W_a = \sum_{i=1}^{N_a} l_i$ be the arithmetic work and $W_c = \sum_{i=1}^{N_a} c_i$ be the communication work. Replacing 2.7 in 2.6 leads to:

$$(p-1)T_p \leq (p-2)(T + hC_d) + W_a + hW_c + 4N_a N_d(p-1)\sigma$$

which concludes the proof. □

As a corollary we consider a coarse-granularity efficient PRAM algorithm with polynomial regularity and bounded degree. For the corresponding DFG, this implies that for an input x of large enough size:

- polynomial speed-up: pT (note that p is fixed) is neglected compared to W_a;
- polynomial granularity: since p and h are assumed constant, taking a sufficient large size instance leads to neglect phC_d and hW_c compared to W_a;
- polynomial regularity: $4N_a N_d \sigma$ (note that σ and N_d are bounded) is also neglected..

This leads to the following result.

Theorem 13 *Let $\mathcal{A}$ be an efficient ATH PRAM program that has polynomial granularity, polynomial regularity and and bounded degree ($N_d = O(1)$).*
Then, for any $\epsilon > 0$, execution time of $\mathcal{A}$ on a distributed architecture with p processors is asymptotically bounded by:

$$T_p(x) < (1+\epsilon)\frac{W_a}{p}.$$

This time includes communication and scheduling overheads.

To obtain an near-optimal scheduling algorithm, we use slackness; we consider the execution of the previous scheduling algorithm on a machine with q identical processors, $q \geq p$. Taking into account the above remarks, execution time including communication and scheduling overheads is bounded by:

$$T_q(x) < (1+\epsilon)\frac{W_a}{q-1}.$$

We can now emulate this machine on the one with p processors; corresponding execution time is

$$T_p(x) \leq \lceil \frac{q}{p} \rceil T_q(x) < (1+\epsilon)\left(1-\frac{1}{q}\right)\frac{W_a}{p}.$$

Choosing q enough larger than p and considering large size enough input data concludes the proof. □

Another way to obtain near-optimal simulation consists in using a distributed list-scheduling strategy. A classical example is *randomized work-stealing*: when a processor becomes idle, it selects uniformly at random a processor to steal a task. When a processor creates a task, it keeps it locally. Such a strategy is theoretically studied in [7]. Asymptotic bounds are given in the framework of strict multi-threaded computations. Other variants export tasks when exceeding a certain number of task creations.
Such list scheduling strategies are very popular in parallel functional languages such as Multilisp [32] or Prolog [17].

In the last section, we turn to an effective implementation of the ATH language which allows the building of the DFG and thus the effective use of the above provably optimal on-line scheduling algorithm.

2.3.4 Athapascan: a simulation of the ATH PRAM language

ATHAPASCAN [43] is a parallel procedural language, inspired by Jade [50], that allows the construction of the DFG of an application during the execution. It thus makes possible the use of provably optimal on-line scheduling algorithms. We give in this section an overview of the main features of the language.

Similar to the ATH language introduced in chapter 1, ATHAPASCAN supports CUMULATIVE-CRCW PRAM algorithms. The building of the DFG is implicit; the `fork` operation (called `new_task` in ATHAPASCAN) may take in argument an optional scheduling strategy, default being a distributed list-scheduling algorithm. Taking benefit of some knowledge on the graph, this allows to choose a well-suited scheduling algorithm such as block-scattering for dense matrix computations or DSC for DAG with known durations [25].

2.3.5 The ATHAPASCAN programming model

The ATHAPASCAN language is strict and para-functionnal. It is implemented by a C++ library; it uses inheritance and templates to provide a friendly and easy-to-use interface.

In ATHAPASCAN, parallelism is expressed by asynchronous procedure calls, which correspond to the building of *tasks*. A task describes the execution of a specific procedure (which is defined by formal parameters and a block of instructions) with effective parameters. Two parameter-passing modes are possible: the by value mode copies the effective parameter into the local memory of the task and the by reference mode shares the data among different tasks.

References to shared data are typed according to their access modes. Four modes are defined to access shared data: *read* (`a1_shared_r`), *write* (`a1_shared_w`), *read/write* (`a1_shared_r_w`) and *accumulation* (`a1_shared_cw`). The three first modes are standard and are used in other parallel languages [38, 50]. Accumulation is realized from the initial value of the object by incrementation; this incrementation is defined by a binary function `f` (default is the C++ operator +=) which is assumed to be *associative* and *commutative*.
Thus, ATHAPASCAN allows the implementation of CUMULATIVE-CRCW PRAM algorithms.

The semantics of ATHAPASCAN [5] is such that each reading of a shared datum gets the value of the last update (writing) in the sequential order of task executions (depth-first ordering). To make such an order easy to compute, ATHAPASCAN does not allow side effects on shared variables. In the current implementation of ATHAPASCAN, this semantics is implemented in the following way: a task becomes *executable* when all the effective parameters that it requires in read (or read/write) mode have been updated by the predecessor tasks (relative to the sequential order of task creations).

2.3.6 Execution model of ATHAPASCAN

The control of the execution is based on the building of a macro data flow graph. This DFG is represented by a direct acyclic hyper-graph, which is distributed among the processors. Vertices correspond to tasks and edges to data dependencies related to shared objects: hyper-edges are used to describe concurrent writings and concurrent readings on shared objects. This graph can

[5] ATHAPASCAN [43] allows other access to shared objects: postponed (suffix `p`) access allow the expression of a larger degree of parallelism and arrays of shared objects.

be labeled with information attributes (*arithmetic cost* for tasks and *data size* for shared object dependencies). This graph is used to implement both the semantics and the scheduling of tasks. Different scheduling algorithms (denoted as *schedulers*) are available and user-specific ones may be added. The role of the scheduler is restricted to informing the system where and when tasks have to be executed, taking into account some information available from the graph. This functionality makes possible the implementation of different classical provably good scheduling algorithms (list scheduling, ETF [11], DSC [26], work-stealing [38] for example).

The following rules define the way an execution is handled:

- The first executable task is the `a1_main()` function.
- During the execution of a task:
  - when a task is created (call to the `a1_new_task` directive), the new task is inserted into the graph;
  - when a task terminates, shared data that it accessed in write or read/write mode are updated. The task is then removed from the graph and the scheduler is informed of new ready tasks (i.e. all shared objects accessed in read or read/write mode are available).
- The scheduler analyzes the graph to make task mapping and starting decisions. The system performs the scheduling decision. When all shared data required by a task in read/read-write mode have been received at the affected node, the task is started on the processor it has been assigned.

2.3.7 An example of ATHAPASCAN program

The figure 2.3 presents an ATHAPASCAN source code for the triangular resolution of $AX = B$; the algorithm is presented in the abstract language ATH in chapter 1 (fig. 1.6). Note that in ATHAPASCAN, all acess modes (read, write or read/write) are explicitly given in order to ensure that the sequential order of execution can be determined directly from task creation (`a1_new_task` instruction).

2.4 Conclusion

In this chapter, the on-line scheduling of a parallel PRAM program on a distributed architecture with a bounded number of processors has been analyzed. List-scheduling strategies, frequently arising in parallel language implementations, have theoretical foundations. An optimal simulation of a PRAM program with polynomial speed-up, polynomial regularity and coarse-granularity is given; costs of communications are considered under the model LogP and a shared-memory is emulated using hash functions.

Due to its experimental good performances [57, 56], most of languages implementing dynamic parallelism use heuristics based on list-scheduling. They essentially differ on the shape of the DFG, depending on the programming model they implement. Thus, the performance of list scheduling

```
struct Update : public a1_task_elem {
  Update( int size ) {
    set_cost(size*size*size);
  }
  // Performs X += -1/A*Y
  void operator() ( a1_shared_cw<matrix<float> > X,
                    a1_shared_r<matrix<float> > A,
                    a1_shared_r<matrix<float> > Y) {
    X.cumul( - A.read().inverse() * Y.read() );
  }
}

struct FinalDivision : public a1_task_elem {
  FinalDivision( int size ) {
    set_cost(size*size*size);
  }
  // Performs X = 1/A*X
  void operator() ( a1_shared_rw<matrix<float> > X,
                    a1_shared_r<matrix<float> > A) {
    X.write( A.read().inverse() * X.read() );
  }
}

struct TriangularSolve : public a1_task {
  TriangularSolve( int nb_elem ) {
    set_cost(nb_elem*nb_elem/2);
  }
  // Performs triangular resolution A*X=B
  // A is coded such that A[n*i+j] ::= A[i][j]
  void operator() (int n,
                   a1_array_of_shared_rp<matrix<float> > A,
                   a1_array_of_shared_cw<<matrix<float> > X,
                   a1_array_of_shared_rp<matrix<float> > B) {
    for(int i=0; i<n; i++) {
      X[i].cumul( B[i].read() );
      a1_new_task( FinalDivision(), X[i], A[n*i+i] );
      for(int j=i+1; j<n; j++)
        a1_new_task(Update(), X[j], A[n*i+j], B[j]);
    }
  }
}
```

Figure 2.3: Triangular resolution of $AX = B$

may vary depending on this model. For instance, if synchronizations are authorized in the language (waiting for some future value for instance), the scheduling has to use migration; if not, no guarantee can be given on the competitive ratio.

We focus in this conclusion on languages that use a provably efficient on-line scheduling algorithm. HPF 2 introduced groups of independent tasks of unknown durations via function calls. A BSP [54] program execution consists in a sequence of super-steps, each set of independent tasks. All shared memory access performed at a step are effective at the next one. Dynamic load balancing is possible [54] but requires task migration in the considered implementation [28].

Functional languages have been using list-scheduling for a long time. For a survey on parallelism in functional languages, see [33], we just mention here some characteristic languages. Sisal[44] is a data-flow based language which defines a fine grain DFG; however, programming macro-tasks in order to obtain a coarse-granularity algorithm is not directly possible. NESL [5] provides a nested parallel model: graphs correspond to recursive n-ary sets of independent tasks with no data-dependencies but synchronization at the join point. Access are emulated on a virtual shared memory. Cilk [7, 38] is inspired from Multilisp and implements a model of strict functional computation in a C-like language. Tasks are mapped on the functions; all data are accessed in the stack. Functions can be migrated at a synchronization point, explicitly defined in the program. Migrations are reduced to a copy of the stack. ATHAPASCAN [19, 43] is inspired from Jade [50]; it is a C++ library that implements a programming model similar to the language ATH presented in the previous chapter. Data-dependencies are defined by access to a shared data. Tasks correspond to procedure calls; parameters can be passed by value or by reference to a shared-data. This last mode defines the precedence. When a task is ready, it can be executed till completion with no synchronization.

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Contents

1 Parallel efficient algorithms **3**
1.1 PRAM, DFG and cost analysis 4
1.2 Increasing granularity 13
1.3 Redundancy and cascading divide&conquer 18
1.4 Randomization to decrease time or preserve work. 23
1.5 Parallel time complexity and NC Classification 26
1.6 Conclusion 27

2 Programming models and scheduling **33**
2.1 Asynchronous distributed architectures 35
2.2 How to schedule a DFG 39
2.3 On-line scheduling algorithms 43
2.4 Conclusion 50