Approximation Algorithms for the Wafer to Wafer Integration Problem*

Trivikram Dokka[1], Marin Bougeret[2], Vincent Boudet[2], Rodolphe Giroudeau[2], and Frits C.R. Spieksma[1]

[1] ORSTAT, K.U.Leuven, Naamsestraat 69, B-3000 Leuven, Belgium
{trivikram.dokka,frits.spieksma}@econ.kuleuven.be
[2] LIRMM Montpellier, France

Abstract. Motivated by the yield optimization problem in semiconductor manufacturing, we model the wafer to wafer integration problem as a special multi-dimensional assignment problem (called WWI-m), and study it from an approximation point of view. We give approximation algorithms achieving an approximation factor of $\frac{3}{2}$ and $\frac{4}{3}$ for WWI-3, and we show that extensions of these algorithms to the case of arbitrary m do not give constant factor approximations. We argue that a special case of the yield optimization problem can be solved in polynomial time.

Keywords: wafer-to-wafer integration, approximation, computational complexity, efficient algorithm.

1 Introduction

Consider the following problem. Given are m sets V_i, $i = 1, ..., m$. Each set contains n p-dimensional vectors; each entry of each vector is a nonnegative integer. We define the *cost* of vector $u = (u_1, u_2, \ldots, u_p)$ as follows: $c(u) = \sum_{i=1}^{p} u_i$. Given a pair of vectors u, v, we can construct the vector $u \vee v$ by defining the operation $\vee$ as follows:

$$u \vee v = (\max(u_1, v_1), \max(u_2, v_2), \ldots, \max(u_p, v_p)).$$

Notice that $(u \vee v) \vee w = u \vee (v \vee w)$.

Consider now an m-tuple, ie, a set of m vectors $u^1, u^2, \ldots, u^m \in V_1 \times V_2 \times \ldots \times V_m$. The cost of an m-tuple equals $c(u^1 \vee u^2 \vee \ldots \vee u^m)$. Our problem, that we denote by WWI (see Section 1.1), is to find n disjoint m-tuples such that each vector is used exactly once, while total cost is minimum. In the figure below an instance with $m = 3, n = p = 2$ is depicted; notice that this instance has the property that each vector is a 0-1 vector; the value of an optimal solution to this instance equals 2 (this is achieved by joining the first vector of V_1, the second vector of V_2, and the first vector of V_3 into $(1, 0)$; the remaining three vectors form $(0, 1)$).

We were motivated to look at this optimization problem by an application in the semi-conductor industry, that we now proceed to describe.

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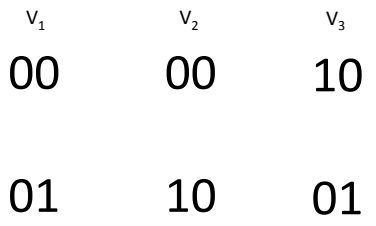

Fig. 1. WWI instance; $m = 3, n = p = 2$

1.1 The Application

Our understanding of the semi-conductor industry, and in particular the wafer-to-wafer production process is primarily based upon Reda et al. [6], Verbree et al. [9], Taouil and Hamdioui [8]. In the semi-conductor industry, Through Silicon Vias (TSV) based three-Dimensional Stacked Integrated Circuits (3D-SIC) is an emerging technology that provides large benefits: a smaller footprint, a higher interconnect density between stacked dies, higher performance, and lower power consumption due to shorter wires when compared to planar IC's. One of the key steps in the production of 3D-SIC's is stacking. There are three different ways of stacking: (1) wafer to wafer (2) die to wafer (3) die to die (see [6]). Of these three approaches, wafer to wafer stacking offers the highest manufacturing throughput coupled with other advantages. However, wafer to wafer stacking approach suffers from a drawback that it may have a low yield. The main motivation of this paper is to study this yield optimization problem in the wafer to wafer integration process.

The yield optimization problem in the semi-conductor industry can be informally described as follows: there are m lots of wafers called wafer lots, with each wafer lot consisting of n wafers. A wafer consists of a string of bad dies and good dies; in our context this translates to a '0' in case of a good die, and a '1' in case of a bad die (such a string corresponds to a vector in the description of WWI). The objective is to form n stacks (a stack corresponds to an m-tuple) by integrating one wafer from each lot (a set V_i) while optimizing the yield i.e., minimizing the total number of bad dies in the resulting stacks (which is equivalent to maximizing the total number of good dies in the resulting stacks). Integrating two wafers can be seen as superimposing the two corresponding strings; in this operation the position in the merged string is only 'good' when the two corresponding entries are good, otherwise it is 'bad'. Due to this reason we call the above problem the wafer to wafer integration (WWI) problem. We refer to it as WWI-m, where m is the number of wafer-lots.

Notice that the yield optimization problem described here is a special case of WWI, since instances of the yield optimization problem have 0-1 vectors (instead of vectors with arbitrary integral entries).

Dimensions of typical instances occurring in the semi-conductor industry have values of m, n, and p ranging as follows: $3 \leq m \leq 10, 25 \leq n \leq 75, 500 \leq p \leq 1000$ (see [6][9]).

1.2 Goal and Related Work

Our main intention in this paper is to formulate the WWI-m as a combinatorial optimization problem and study it from an approximation point of view. Usually, the yield optimization problem is formulated as a maximization problem, however, we feel that studying the minimization problem is especially relevant from approximation point of view. Indeed, owing to the fact that in the yield optimization instances, the number of bad dies in each wafer is typically much less than the number of good dies, it make sense to be able to approximate the (smaller) minimization optimum instead of the (larger) maximization optimum.

There is increasing attention for the yield optimization in the literature. One example is the contribution [6]. In [6] the problem is formulated as an multi-index assignment problem; further, computational performance with straightforward heuristics is reported. Some recent work on this problem is also reported in [8] [9]. As we will show, WWI can be seen as a multi-index assignment problem where the costs have a certain structure. Three dimensional assignment problems with so-called decomposable costs have been studied in Crama and Spieksma [4] and Burkard et al. [3]. Multi-dimensional assignment extensions of this cost structure appear in Bandelt et al. [1]. An survey on multi-dimensional assignment problems can be found in [7] and chapter 10 of [2].

1.3 Our Results

Our results can be summarized as follows:

- We present an IP-formulation that is an alternative to the traditional formulation given in [6]. This alternative formulation contains fewer variables, and may be more suited from a computational perspective (see Section 2).
- We prove that the yield optimization problem is NP-hard (see Section 3).
- We give two simple approximation algorithms for WWI-3, one with a $\frac{3}{2}$ performance guarantee, and one with a $\frac{4}{3}$ performance guarantee (see Section 4.1). We also show that natural extensions of these algorithms to the case of arbitrary m fail to provide a constant-factor guarantee (see Section 4.2).
- We show that, in case of a fixed m and a fixed p, the yield optimization problem is solvable in polynomial time (see Section 5).

2 Problem Formulation

In subsection 2.1 we give a straightforward formulation of the yield optimization problem in wafer to wafer integration as a m-dimensional axial assignment problem, see also [6]. Section 2.2 presents an alternative IP-formulation that may be more suited from a computational perspective.

2.1 IP Formulation

We set $K = V_1 \times V_2 \times \ldots \times V_m$, ie, K corresponds to the set of m-tuples. Next, for each $a \in K$, there is a binary variable x_a indicating whether m-tuple a is selected ($x_a = 1$) or not ($x_a = 0$). The formulation is now as follows (see also [6])

$$\min \quad \sum_{a \in K} c(a) \cdot x_a \tag{1}$$

$$\sum_{a:\ u \in a} x_a = 1 \ \text{ for each } \ u \in \cup_{i=1}^{m} V_i, \tag{2}$$

$$x_a \in \{0,1\} \ \text{ for each } a \in K. \tag{3}$$

Observe that constraints (2) ensure that each vector u is in an m-tuple.

2.2 Alternative IP Formulation

In this section we give an IP formulation that is different from the classical formulation and contains fewer variables.

In this formulation, we model the problem by treating V_1 as the *hub*. Each vector in $\cup_{i=2}^{m} V_i$ is assigned to a vector in V_1; this decision is modelled by a binary variable as follows. There is a variable $z_{u,v}$, where $u \in V_1$ and $v \in \cup_{i=2}^{m} V_i$, such that:

$$\begin{aligned} z_{u,v} &= 1 \text{ if vectors } u \text{ and } v \text{ are contained in the same } m\text{-tuple,} \\ &= 0 \text{ otherwise.} \end{aligned}$$

In addition, we introduce variables $y_{u,\ell}$ as follows.

$$y_{u,\ell} = \text{the value in the } \ell\text{-th position of the } m\text{-tuple containing vector } u \in V_1.$$

$$\min \sum_{u \in V_1} \sum_{\ell=1}^{p} y_{u,\ell} \tag{4}$$

$$\sum_{u \in V_1} z_{u,v} = 1 \quad \text{for each } \ v \in \cup_{i=2}^{m} V_i, \tag{5}$$

$$\sum_{v \in V_i} z_{u,v} = 1 \quad \text{for each } \ u \in V_1, \text{ for each } i = 2, \ldots, m, \tag{6}$$

$$y_{u,\ell} \geq \max(u_\ell, v_\ell) \cdot z_{u,v} \quad \text{for each } u \in V_1, \text{for each } v \in \cup_{i=2}^{m} V_i, 1 \leq \ell \leq p, \tag{7}$$

$$z_{u,v} \in \{0,1\} \quad \text{for each } u \in V_1, \text{ for each } v \in \cup_{i=2}^{m} V_i. \tag{8}$$

Here, $u_\ell(v_\ell)$ denotes the ℓ^{th} entry in the vector $u(v)$. Observe that this alternative formulation has very few variables ($O(mn^2 + np)$) when compared to the number of variables in classical assignment formulation ($O(n^m)$). Even for reasonably small instances it will be difficult to solve the resulting problem with IP solvers using the classical formulation, whereas we might be able to solve them using (4)-(8).

3 The Complexity of WWI

In this section we describe a reduction from MAX-3DM to WWI. Recall that for a given pairwise disjoint sets X,Y,Z, and a set of ordered triples $T \subseteq X \times Y \times Z$, a *matching* in T is a subset of $M \subseteq T$ in which no two ordered triples in M agree in any coordinate. The goal of the MAXIMUM 3-DMENSIONAL MATCHING problem (shortly, MAX-3DM) is to find a matching in T of maximum cardinality.

Kann [5] showed that 3-bounded MAX-3DM is APX-complete.

Reduction. Consider an arbitrary instance I of MAX-3DM with three sets $X = \{x_1, ..., x_q\}$, $Y = \{y_1, ..., y_q\}$, and $Z = \{z_1, ..., z_q\}$, and a subset $T \subseteq X \times Y \times Z$. Let the number of triples be denoted by $|T|$.

Starting from the instance I of MAX-3DM, we now build a corresponding instance I' of WWI-3 by specifying V_i $(i = 1, 2, 3)$, as follows:

- for each element in $x_i \in X$ there is a vector $v_{1i} \in V_1$
- for each element in $y_j \in Y$ there is a vector $v_{2j} \in V_2$
- for each element in $z_k \in Z$ there is a vector $v_{3k} \in V_3$
- each vector has length $|T|$ i.e., $p = |T|$; in fact, for each triple $e = (x_i, y_j, z_k) \in T$, there is a position in each vector corresponding to that triple. The three vectors v_{1i}, v_{2j}, and v_{3k} corresponding to triple (x_i, y_j, z_k), have a '0' in that position, all other vectors have a '1' in that position.

This completes the description of WWI-3 instance.

It is easy to see that a solution to an instance of MAX-3DM with value k corresponds to a solution to the corresponding instance of WWI-3 with cost $pq - k$. Thus we can state the following theorem:

Theorem 1. *WWI-3 is NP-hard, even for the special case of* $0-1$ *vectors (i.e., yield optimization).*

Notice that, when the yield optimization problem would have been formulated as a maximization problem, straightforward reductions from MAX-kDM imply that a constant factor approximation for the maximization version of WWI-m would imply P=NP.

4 Approximation Algorithms for WWI-m

In this section we first prove that a straightforward algorithm (called heuristic H) for WWI-3 is a $\frac{3}{2}$ approximation algorithm. We show how a simple modification of this heuristic allows us to improve the worst-case ratio to $\frac{4}{3}$. Finally, we show that a natural extension of heuristic H to WWI-m can perform arbitrarily bad.

4.1 The Case $m = 3$

Algorithm 1. Heuristic H

1. Solve an assignment problem between V_1 and V_2, based on costs $c(u \vee v)$, $u \in V_1$,$v \in V_2$. Call the resulting matching M.
2. Solve an assignment problem between M and V_3 based on costs $c((u \vee v) \vee w)$, $u \vee v \in M$, $w \in V_3$.

Theorem 2. *Heuristic H is a $\frac{3}{2}$-approximation algorithm for WWI-3. This bound is tight.*

Proof. We first introduce some notation. Let OPT denote the value of an optimal solution, and let $cost(H)$ refer to the value of the solution found by H. Let $c(V_i)$ equal total cost of the vectors in V_i, ie, $c(V_i) = \sum_{u \in V_i} c(u)$, for $i = 1, 2, 3$. Let c_{12}^{OPT} denote the value of a partial optimal solution restricted to $V_1 \times V_2$, ie, when we remove from the optimal solution the vectors from V_3; the total weight that remains equals c_{12}^{OPT}. Recall that M refers to matching found by H in the first step, and let c_{12}^{H} be the value of the partial solution obtained after Step 1 of the heuristic H.

Let us call x (y) the amount with which the value of a partial optimal (heuristic) solution increases when vectors from V_3 are matched optimally to the optimal (heuristic) pairs from $V_1 \times V_2$. Thus, by definition:

$$x = OPT - c_{12}^{OPT} \tag{9}$$

Clearly, the following inequality is valid:

$$c(V_3) \leq OPT \tag{10}$$

(9) and (10) imply

$$c(V_3) - x \leq c_{12}^{OPT}. \tag{11}$$

Consider a set U consisting of n p-dimensional vectors with total cost $c(U) = \sum_{u \in U} c(u)$. In addition, consider a set V, also consisting of n p-dimensional vectors. Let us now assign the vectors from V to the vectors of U using as a cost $c(u \vee v)$ for each $(u, v) \in U \times V$. Let the value of the resulting optimal solution be denoted by $c(U \times V)$. We say that an amount equal to $c(V) - (c(U \times V) - c(U))$ from V is *covered* by U (or equivalently, we say that U is *able to cover an amount of* $c(V) + c(U) - c(U \times V)$ from V).

Consider now the partial heuristic solution found after Step 1, ie, consider M.

Lemma 1. *There exists a feasible assignment of the vectors in V_3 to the pairs from M such that at least the amount $\frac{1}{2}(c(V_3) - x)$ from V_3 is covered by M.*

Argument: To argue that the lemma is true, consider the partial optimal solution restricted to $V_1 \times V_2$. Apparently, these n vectors are able to cover an amount of $c(V_3) - x$ from V_3 when assigning the strings from V_3 to these vectors (since

$OPT = c_{12}^{OPT} + x$). However, each vector in $V_1 \times V_2$ consists of p numbers, each one arising from either V_1 or V_2. Thus, we can partition the amount covered $c(V_3) - x$ into two disjoint parts: one part covered by numbers from vectors in V_1, one part covered by numbers from vectors in V_2. It follows that if one considers the following two assignments: one where the vectors from V_3 are assigned to the vectors from V_1 as in the optimal solution, and one where the vectors from V_3 are assigned to the vectors from V_2 as in the optimal solution, that at least one of these solutions will cover $\frac{1}{2}(c(V_3) - x)$. This proves the lemma.

We can now derive

$$\begin{aligned}
\text{cost}(H) &= c_{12}^{H} + y \\
&\leq c_{12}^{H} + c(V_3) - (\frac{1}{2}c(V_3) - \frac{1}{2}x) \\
&= c_{12}^{H} + \frac{1}{2}c(V_3) + \frac{1}{2}x \\
&\leq c_{12}^{OPT} + \frac{1}{2}c(V_3) + \frac{1}{2}x \\
&\leq c_{12}^{OPT} + \frac{1}{2}[c_{12}^{OPT} + x] + \frac{1}{2}x \\
&\leq \frac{3}{2}c_{12}^{OPT} + \frac{3}{2}x = \frac{3}{2}OPT.
\end{aligned}$$

The first inequality follows from Lemma 1, the second inequality follows from the fact that the heuristic, in Step 1, computes an optimum assignment between sets V_1 and V_2 whose costs cannot exceed c_{12}^{OPT}, the third inequality follows from (11) and the final inequality follows from the definition of x. Tightness follows from the instance depicted in Figure 1: observe that, for this instance, OPT = 2, whereas heuristic H might find a solution with value 3. □

A minor modification of heuristic H (denoted by H_{heavy}) allows us to improve the worst-case ratio without actually increasing the computational effort. Indeed, let us slightly modify H by ensuring that in Step 1 the *heaviest* set V_i is present, ie, we ensure that the set V_i for which $c(V_i)$ is maximal, is assigned to some $V_j, j \neq i$ in the first step.

Theorem 3. *Heuristic H_{heavy} is a $\frac{4}{3}$-approximation algorithm for WWI-3. This bound is tight.*

Algorithm 2. Heuristic H_{heavy}

0. Let $j = \arg\max_{i=1,2,3} c(V_i)$.
1. Solve an assignment problem between V_j and some $V_i, i \neq j$, based on costs $c(u \vee v)$, $u \in V_j, v \in V_i$. Call the resulting matching M.
2. Solve an assignment problem between M and the remaining set $V_k, k \neq j, k \neq i$ based on costs $c((u \vee v) \vee w)$, $u \vee v \in M$, $w \in V_k$.

Proof. Let us assume, without loss of generality, that set V_1 is the heaviest set. Thus, we have $c(V_1) \geq c(V_2)$ as well as $c(V_1) \geq c(V_3)$. Even more, let us assume (again wlog) that in Step 1 of H_{heavy} sets V_1 and V_2 are assigned to each other. We distinguish three cases.

Case 1: $0 \leq c(V_1) \leq \frac{1}{3}\text{OPT}$.
This case is trivial since any feasible solution is in fact optimal: $\text{cost}(H_{heavy}) \leq c(V_1) + c(V_2) + c(V_3) \leq 3 \cdot \frac{1}{3}\text{OPT} = \text{OPT}$.

Case 2: $\frac{1}{3}\text{OPT} < c(V_1) \leq \frac{2}{3}\text{OPT}$.
This case is similar to the analysis in Theorem 2. We derive:

$$\begin{aligned}\text{cost}(H_{heavy}) &= c_{12}^{H_{heavy}} + y \\ &\leq c_{12}^{OPT} + c(V_3) - (\frac{1}{2}c(V_3) + \frac{1}{2}x) \\ &= c_{12}^{OPT} + \frac{1}{2}c(V_3) + \frac{1}{2}x \\ &\leq \text{OPT} + \frac{1}{2}c(V_3) \leq \frac{4}{3}OPT.\end{aligned}$$

The last inequality follows from the assumption in this particular case, and the fact that $c(V_3) \leq c(V_1)$.

Case 2: $\frac{2}{3}\text{OPT} < c(V_1) \leq \text{OPT}$.
We denote by Q the weight from V_3 that is covered by V_1 when we solve an assignment problem between V_1 and V_3. The following is true:

$$c(V_1) + c(V_3) - Q \leq \text{OPT}. \tag{12}$$

We now derive:

$$\begin{aligned}\text{cost}(H_{heavy}) = c_{12}^{H_{heavy}} + y \ &\leq c_{12}^{H_{heavy}} + c(V_3) - Q \\ &\leq c_{12}^{OPT} + c(V_3) - Q \\ &\leq c_{12}^{OPT} + \text{OPT} - c(V_1) \\ &\leq \text{OPT} + \frac{1}{3}\text{OPT} = \frac{4}{3}OPT.\end{aligned}$$

The first inequality follows from Step 2 of H_{heavy} (the weight added to $c_{12}^{H_{heavy}}$ will not exceed the 'uncovered' weight from V_3 when we assign the vectors from V_3 to V_1), the second from Step 1 of H_{heavy}, the third inequality follows from (12), and the last inequality follows from the assumption in this particular case.

Tightness follows from the instance depicted in Figure 2: observe that, for this instance, OPT = 6, whereas heuristic H_{heavy} might find a solution with value 8. □

V_1	V_2	V_3
100000	000010	001000
010000	000001	000100
001000	100000	000010
000100	010000	000001
000000	000000	000000
000000	000000	000000

Fig. 2. H_{heavy}: an instance where OPT = 6 and cost(H_{heavy}) = 8

An obvious improvement to heuristic H and H_{heavy} would consist of a heuristic that runs H for all possible pairs in the first step, add the remaining set in the last step, and then choosing the best of the three feasible solutions found. Interestingly, this heuristic (which involves solving 6 assignment problems) does not have a lower worst case ratio than H_{heavy} (which only solves two assignment problems). This also follows from the example depicted in Figure 2.

Notice that heuristic H, in contrast to H_{heavy} can be seen as an online algorithm for a natural, online variant of WWI-3. Indeed, consider the setting where the sets V_1, V_2, and V_3 arrive sequentially over time, and that, before the arrival of a next set, the just arrived set V_i must be assigned to the partial tuples. Results given above imply directly:

Corollary 1. *Heuristic H is a $\frac{3}{2}$ competitive algorithm.*

Clearly, in this framework, H_{heavy} is not an online algorithm.

4.2 The Case of Arbitrary m

A natural extension of heuristic H to the case of arbitrary m is as follows. We iteratively assign set V_i to the existing partial tuples from $V_1 \times V_2 \times \ldots \times V_{i-1}$. Let us call the resulting heuristic H_{seq}. The performance of H_{seq} can be arbitrarily bad as can be seen from the description of the following instances. To understand these instances, it can be helpful to see each vector as a circle with p positions; in such a circle, the 1s, as well as the 0s, will appear consecutively. Let $v_{i,j}$ denote the j-th vector from V_i. Formally, the instances are described as follows:

Choose m such that there exists a value of p with $m = p(p-1)+1$ (thus, in these instances, the length of a vector increases with m), and set $n = p$.

- for each $k \in \{1, \ldots, p-1\}$, there are 1s in position $i-(k-1)p$ to position $i-(k-1)p+k-1$ (modulo p) in vector $v_{i,1}$, for each $i \in \{(k-1)p+1, kp\}$.
- There is a 1 in each position of the vector $v_{(p(p-1)+1,1}$.
- Each other vector is an all-zero vector.

Notice that the cost of an optimal solution equals p, whereas H_{seq} may find a solution with cost $m = p(p-1)+1$.

Corollary 2. *The worst case ratio of H_{seq} is at least $O(\sqrt{m})$.*

An instance with $p = 3$ is depicted in Figure 3.

V_1	V_2	V_3	V_4	V_5	V_6	V_7
100	000	000	110	000	000	111
000	010	000	000	011	000	000
000	000	001	000	000	101	000

Fig. 3. H_{seq}: an instance where OPT = 3 and cost(H_{seq}) = 7

Another natural heuristic to consider is the so-called Multiple Hub-Heuristic (see [1]), which can be informally described as in Algorithm (3):

The performance of the multiple hub heuristic MH can be arbitrarily bad. Indeed, consider the following instance. The length of each vector equals 2, ie, $p = 2$, and consider some even value for the number of sets m. let $n = \frac{m}{2} + 1$. The first vector of each of the sets V_i, $i = 1, 2, \ldots, \frac{m}{2}$ is specified as follows: For $i = 1, 2, \ldots, n$, put $v_{i,1} = (1\ 0)$; for $i = \frac{m}{2} + 1, \ldots, m$ put $v_{i1} = (0\ 1)$. All other vectors in the instance are equal to (0 0). It can be seen that $OPT = 2$ whereas cost(MH) $= \frac{m}{2} + 1$.

Corollary 3. *The worst case ratio of MH is $O(m)$.*

Notice that this performance is in contrast with the performance of the multiple hub-heuristic for other variants of decomposable minimum cost m-dimensional assignment problems, see [1].

Algorithm 3. Multiple-Hub-Heuristic MH

for h = 1 to m **do**
  for i = 1 to m **do**
    1. Solve an assignment problem between V_h and V_i, $i \neq h$, based on costs $c(u \vee v)$, $u \in V_h$,$v \in V_i$. Call the resulting matching M_{hi}.
  end for
  Combine all M_{hi}, to construct M_h.
end for
Output the min-cost solution of all M_h.

5 Fixed p

In this section we consider the yield optimization problem, i.e., we consider instances that feature 0-1 vectors only. We will argue that instances of the yield optimization problem with a fixed p can be solved in polynomial time (for each fixed m).

Consider a solution of the yield optimization problem. It consists of n 0-1 vectors. Thus, we can classify these n 0-1 vectors as belonging to at most 2^p different types (each type corresponding to a distinct 0-1 vector of length p). We use the symbol t to index these types.

We say that a vector from type t is *compatible* with a vector from type s if the vector of type t has a '1' in each of the positions where the vector of type s has a '1'. We write type t is *compatible* with a vector from type s as $t \succ s$. Further, given an instance of the yield optimization problem, we let k_s^i denote the number of 0-1 vectors of type s in set V_i, $s = 1, \ldots, 2^p$, $i = 1, \ldots, m$.

We construct the following formulation that features variables x_t:

$$x_t = \text{number of 0-1 vectors of type } t \text{ in the final solution, } t = 1, \ldots, 2^p.$$

We also need "transportation" type variables; for each $i = 1, \ldots, m, s, t = 1, \ldots, 2^p$:

$$z_{s,t}^i = \text{number of 0-1 vectors of type } s \text{ from set } V_i \text{ assigned to class } t.$$

The formulation (with parameter c_t referring to the number of '1's in a vector from type t):

$$\min \sum_{t=1}^{2^p} c_t x_t \tag{13}$$

$$\sum_{s:\ t \succ s} z_{s,t}^i = x_t \quad \text{for each } t = 1, \ldots, 2^p, i = 1, \ldots, m, \tag{14}$$

$$\sum_{t:\ t \succ s} z_{s,t}^i = k_s^i \quad \text{for each } s = 1, \ldots, 2^p, i = 1, \ldots, m, \tag{15}$$

$$x_t \quad \text{integer} \qquad \text{for each } t = 1, \ldots, 2^p, \tag{16}$$

The objective function (13) minimizes the total cost. Constraints (14)-(15) are the familiar transportation constraints. Notice further that integrality of x_t implies integrality of $z_{s,t}^i$.

Observe that this formulation involves $O(2^p)$ binary variables, $O(m2^{2p})$ continuous variables, and $O(m2^p)$ constraints.

Lemma 2. *Formulation (13)-(16) is correct.*

Proof. See Appendix.

When we fix p and m the above formulation has a fixed number of variables and constraints. Thus we can use Lenstra's algorithm to solve this IP in polynomial time. This implies:

Corollary 4. *For each fixed p, and for each fixed m, the yield maximization problem can be solved in polynomial time.*

It is true, however, that for the values of p encountered in practice, the above sketched approach is not practical.

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Appendix

Proof of Lemma 2

Proof. Consider a feasible solution to the yield optimization problem. This solution prescribes for each type of vector in each set V_i how many of these vectors are assigned to a vector of type t. This determines the $z^i_{s,t}$ values; clearly, these values will satisfy constraints (14)-(16), since our solution is valid. Vice versa, consider $z^i_{s,t}$ values that satisfy (14)-(16). One can construct a feasible solution to WWI-m as follows: (1) Create a set X of n vectors with x_t vectors of type t. (2) Solve an assignment problem between X and V_i, for each $i = 1, \ldots, m$, based upon a bipartite graph $G = (L \cup R, E)$ involving a node in L for each vector in X a node in R for each vector in V_i, and two nodes are connected if the vector in X is compatible with the vector in V_i. This assignment problem will have a feasible solution since $z^i_{s,t}, x_t$ satisfy (14)-(16). (3) Construct m-tuples of vectors by matching m vectors one from each V_i together in an m-tuple if they all are matched to same vector in X in (2). This corresponds directly to a feasible solution. □