

Thesis

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Introduction

This thesis explores the curious mathematics behind a theorem covered in a typical course in algorithms and data structures, namely the Max-flow/Min-cut theorem. The aim of the thesis is to demonstrate the intimate relationship between problems in combinatorial optimization, and a field called linear programming. In order to do this, we develop useful tools along the way. We first review concepts in elementary graph theory (bipartite graphs and matchings), and introduce the topic of linear programming. We then look at a problem called the vertex problem, which we will discover is fundamentally related to the matching problem on bipartite graphs. We discover this relationship using duality theory, an important aspect of linear programming. As a consequence of our duality theorems we are show that the max-flow/min-cut relationship in flow networks is just a special case of linear programming duality. After this, we turn our attention to the main algorithm in this thesis: the Hungarian algorithm. This algorithm for solving weighted bipartite matchings was developed in the 50s by Harold Kuhn. This algorithm is significant in that it used duality theory as its fundamental design principle; in many ways it predicted the use of linear programs to solve combinatorial optimization problems. Finally, we take a step back in our final chapter and look at the general “primal-dual method” inspired by the Hungarian algorithm. This method uses general linear programming principles to design algorithms for network design problems. The main result here is that, quite magically, we are able to turn *weighted* problems into equivalent unweighted versions, which are often easier to solve. We finish by describing a class of auction algorithms

which use this primal-dual methodology.

Chapter 1

Bipartite graphs, matchings, and linear programming

In this chapter we review bipartite graphs and matchings, and discuss their relevance to this thesis. We will also begin our discussion of linear programming in a casual way to give the reader a taste for what is coming.

1.1 Bipartite graphs and matchings

Throughout this thesis we will be interested in a specific subclass of graphs known as bipartite graphs. Unless otherwise noted, our algorithms will assume a bipartite structure.

Definition 1.1. A *bipartite graph* is a graph whose vertices can be partitioned into two sets U and V such that all edges connect a vertex $u \in U$ to a vertex $v \in V$. We will denote this graph $G = (U, V, E)$, where E is the edge set $(U \times V)$.

In Figure 1 we have a bipartite graph with vertex partition given by $U = \{A, B, C, D\}$ and $V = \{E, F, G\}$. All edges in this graph are between a node $u \in U$ and a node $v \in V$. The graph on the right is also bipartite. It may take a little more time to convince yourself that you can partition the vertices into disjoint U and V in a way that maintains the

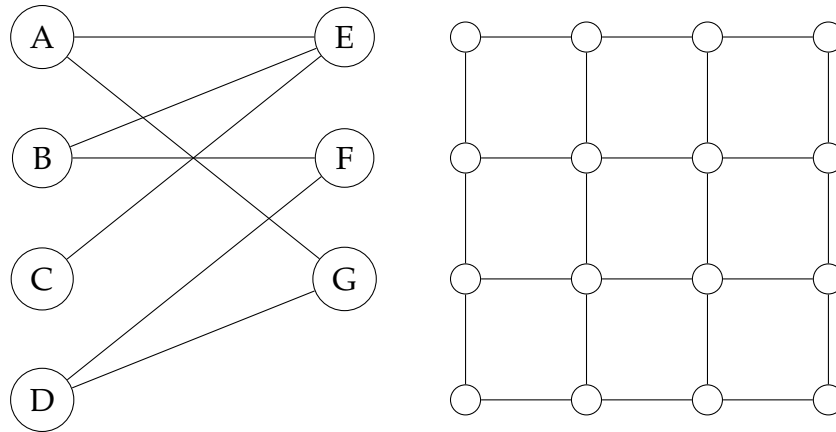


Figure 1.1: Examples of bipartite graphs

bipartite property. Try it! Now that we know what we are working with, let's introduce a problem that we'd like to solve on these graphs.

Definition 1.2. Let $G = (U, V, E)$ be a bipartite graph. A subset $M \subset E$ is a *matching* if no two edges in M are incident to the same vertex. We call a matching *perfect* if all vertices are an endpoint of an edge in M .

We say that a vertex $w \in U \cup V$ is *matched* with respect to M if it is an endpoint of some edge in M . The following figure shows some examples of different matchings on one of the graphs from Figure 1, where the bold edges are the edges in the matching. There many different valid matchings on the graph in Figure 2. Oftentimes, we want to

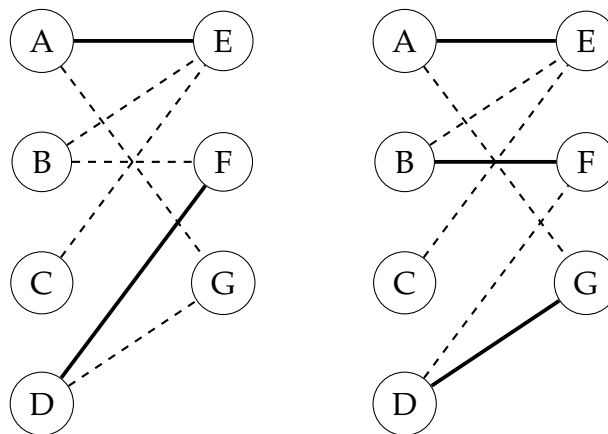


Figure 1.2: Examples of matchings on a bipartite graph

find the largest matching on a graph. This is classically motivated by economic examples, where we have a set of bidders, and a set of goods, and the edges between them denote a bidder i 's willingness to pay for good j . For a money-hungry auctioneer, the goal here would be to find the “largest” matching on the graph, i.e. the one that maximizes profit of the auction. We will discuss this interpretation in more depth later on in the thesis, but it naturally leads to the following definition.

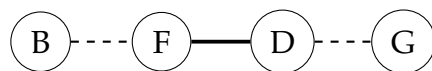
Definition 1.3. A *maximal matching* on G is a matching M such that if any other edge not in M is added to M , it is no longer a valid matching. Alternatively put, M is maximal if there is no matching M' such that $M \subset M'$.

Both matchings in Figure 2 are maximal matchings; in each case there are no edges that we can add to M and have that M is still a matching. However, notice that these matchings have different sizes, even though both are maximal on G . This leads to the following definition.

Definition 1.4. A matching M on a graph G is said to be a *maximum matching* if for all other matchings M' on G , $|M'| \leq |M|$.

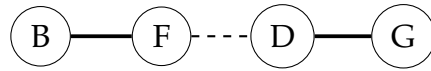
In our example, the matching on the right given by $M = \{(A, E), (B, F), (D, G)\}$ is a maximum matching (convince yourself). In general there may be many unique maximum matchings on a graph.

In this section we are interested in general methods for finding maximum matchings on bipartite graphs. One of the fundamental approaches is to look at certain subgraphs called alternating paths. Before we define what these are, let's look at a motivating example. Suppose we have the matching on the left in Figure 2. So $M = \{(A, E), (D, F)\}$. Consider the following sequence of vertices in the graph: Call this sequence p . Let's perform an



operation that we will denote $M \oplus p$, which operates like XOR: add to M each edge in

p that isn't in M , and remove from M each edge in p that is in M . This gives us the following segment, where (B, F) and (D, G) are now in M , but (D, F) is not: First, we



must check that the new M is still a valid matching; we do this by noticing that B and G were originally unmatched, so it's okay for one of their incident edges to be added. Also, notice that the size of our matching has grown by 1! In fact, this new matching is exactly the matching given by the graph on the right in Figure 2. This is a general technique in finding maximum matchings. We want to look for these paths that start and end at unmatched vertices, and whose edges are alternately matched and unmatched. If we can find one of these paths, we will be able to increase the size of matching. We define this formally now.

Definition 1.5. Let G be a graph and M some matching on G . An *alternating path* is a sequence of vertices and edges that begins with an unmatched vertex, and whose edges alternate between being in M and not in M .

Definition 1.6. An *augmenting path* is an alternating path that starts and ends on unmatched vertices. When we augment M by an augmenting path p , we use the notation $M \oplus p$.

This motivates a general method for finding a maximum matching on a bipartite graph: just keep looking for augmenting paths, and augment the current matching by that augmenting path. Of course, we need to prove that this in fact gives us a maximum matching. The following theorem says exactly that.

Theorem 1.7. (Berge, 1957) *A matching M on G is a maximum matching if and only if G contains no augmenting paths with respect to M .*

This gives us the following framework for finding maximum matchings in bipartite graphs.

ALG 1(G)

```
1   $M := \emptyset$ 
2  while there exists an augmenting path  $p$ 
3       $M := M \oplus p$ 
4  return  $M$ 
```

Note that we have yet to describe the details of this algorithm. Before we do so, we are going to take a step back a bit and look at the maximum matching problem from a slightly different perspective. In doing so, we will develop a language for talking about this problem that will serve us throughout the rest of this thesis. At first, the approach will appear purely pedagogic, but hopefully the reader will understand the significance of it by the end of the thesis.

1.2 Linear programming

The development of combinatorial optimization has been deeply intertwined with the discipline of linear programming. In its most basic form, linear programs are given by some linear objective function that you want to optimize, along with some linear constraint equations. For a more detailed treatment on linear programming, we will refer you to (CITE LP SOURCES). For the purposes of this thesis, we will treat linear programming more casually, only requiring a few key results. Moreover, we will not be discussing methods of actually solving linear programs, for which there are at least a couple of well known but complicated algorithms.

In the general linear-programming problem, our goal is to optimize some linear function that is constrained by a set of linear inequalities. These problems are ubiquitous in applied math and computer science, as they model a system in which something needs to be optimized according to competing resources. We can express a general *maximization*

linear program as

$$\text{maximize} \quad \sum_{j=1}^n c_j x_j \quad (1.1)$$

$$\text{subject to} \quad \sum_{j=1}^n a_{ij} x_j \leq b_j, \quad i = 1, \dots, m \quad (1.2)$$

$$x_j \geq 0, \quad j = 1, \dots, n. \quad (1.3)$$

We call the function in (1) our *objective function*, and the linear inequalities (2) and (3) our constraints. Similarly, a *minimization* linear program takes the form

$$\text{minimize} \quad \sum_{j=1}^n c_j x_j \quad (1.4)$$

$$\text{subject to} \quad \sum_{j=1}^n a_{ij} x_j \geq b_j, \quad i = 1, \dots, m \quad (1.5)$$

$$x_j \geq 0, \quad j = 1, \dots, n. \quad (1.6)$$

Let's see an example of a linear program:

$$\begin{aligned} \text{maximize} \quad & x_1 + x_2 \\ \text{subject to} \quad & 4x_1 - x_2 \leq 8 \\ & 2x_1 + x_2 \leq 10 \\ & 5x_1 - 2x_2 \geq -2 \\ & x_1, x_2 \geq 0. \end{aligned}$$

For a simple linear program such as this, one may use basic substitution/elimination methods to get a solution; in this case it turns out that $x_1 = 2$ and $x_2 = 6$ is an optimal solution. Note that our constraint equations specify a solution space in \mathbb{R}^2 . In general, one can imagine n -dimensional solution spaces in which more complicated algorithms are re-

quired for finding solutions. For those interested, the two main algorithms for this are the Simplex method and the Ellipsoid method; detailed treatment on these algorithms can be found [CITE]. These algorithms are beyond the scope of this thesis, but it is important to note that they exist. Many problems which on the face may not appear to be optimization problems turn out to be easily rephrased as linear programs. Our goal in the next section will be to describe two problems in terms of their linear programs. One of these problems we've already looked at, which is the maximum matching problem. The other will be a problem called the minimum vertex cover problem. These two problems share a very fundamental relationship, which we will discover via their linear programs. We will also revisit a problem that algorithms students are familiar with, which is the max-flow/min-cut theorem. However, we will approach it from a different perspective, using linear programs, which will hopefully demonstrate the cool, deep math that is behind the relationship between flows and cuts.

Chapter 2

Vertex cover, duality theory, and the max-flow/min-cut relationship

This chapter aims to describe fundamental relationships between combinatorial optimization problems and linear programming. More specifically, we will define a new problem – the vertex cover problem – and show how this problem is intimately related to the matching problem. The way we show this relationship is through duality theory. We also use this to shed light on the max-flow/min-cut theorem.

2.1 The vertex cover problem

We begin this section by defining a new problem. Again, this is a problem on graphs in general, but we will be restricting our attention to bipartite graphs. This is a good idea for many reasons, but the most pertinent reason is that this problem is NP-complete in the general case.

Definition 2.1. Let $G = (U, V, E)$ be a bipartite graph. A subset $C \subset U \cup V$ is said to be a *vertex cover* if for each $(u, v) \in E$ we have that at least one of $u, v \in C$. C is a *minimum vertex cover* if for any other cover C' , $|C| \leq |C'|$.

Using what we've already learned, we can specify at least one relation between matchings and vertex covers: namely, the set of all vertices of all edges in any maximal matching on a graph forms a vertex cover. Figure 1 shows some examples of vertex covers on the graph we looked at in the previous chapter.

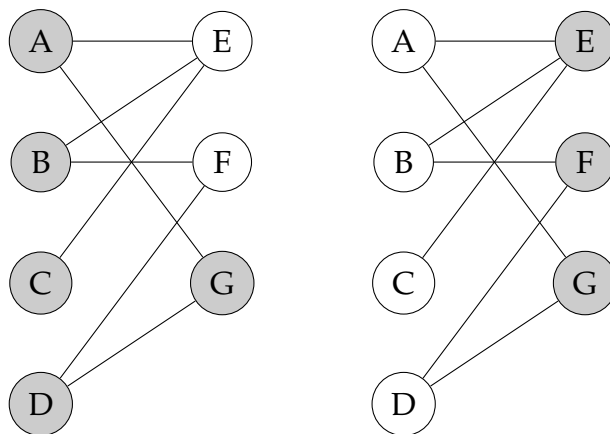


Figure 2.1: Examples of vertex covers

You can convince yourself that the cover on the right is a minimum cover. This brings us to an important theorem. We first prove a lemma.

Lemma 2.2. Let $G = (U, V, E)$ be a bipartite graph. Let M be a matching on G and C a cover on G such that $|M| = |C|$. Then M is a maximum matching and C is a minimum covering.

Proof. Let M' be a maximum matching on G and C' a minimum covering on G . For each $(u, v) \in M'$, C' must include either u or v , which tells us that $|M'| \leq |C'|$. Then we have

$$|M| \leq |M'| \leq |C'| \leq |C|.$$

Thus, if $|M| = |C|$ we have equalities above, which means that the size of a maximum matching is equal to the size of a minimum covering. \square

Theorem 2.3. (König-Egervary) For any bipartite graph G , if M is a maximum matching on G and C is a minimum vertex cover on G , then $|M| = |C|$.

Before we prove this, we define a simple term: an M -alternating path is an alternating path with respect to a matching M .

Proof. Let $G = (U, V, E)$ be a bipartite graph, and let M be a maximum matching on G . Furthermore, define

$$A := \{s \in S \mid s \text{ unmatched}\} \quad (2.1)$$

$$B := \{\text{all vertices connected to nodes in } A \text{ by } M\text{-alternating paths}\}. \quad (2.2)$$

Let $L = B \cap U$ and $R = B \cap V$. Then we have the following:

1. Every node in R is saturated.
2. $N(L) = R$,

where $N(L)$ denotes the set of all vertices connected to elements of L (the “neighbors” of L). The first claim comes from the fact that, if M is a maximum matching, then our alternating paths starting at nodes in A must have length ≥ 2 , and must have even length (otherwise we would have an augmenting path, which contradicts our assumption that M is a maximum matching). The second comes from that fact that every node in $N(L)$ is connected to vertices in A by an alternating path.

Now, define $K := (U \setminus L) \cup R$. Every edge in G must have one of its endpoints in K . If not, there there would be an edge with one end in L and one end in $V \setminus R$, which contradicts $N(L) = R$. So K is a covering of G . Moreover, $|K| = |M|$, since for each edge in M we’ve included one of its endpoints in K (the vertices we’ve chosen are those in $N(L)$ and those in $U \setminus L$). Thus, by the previous lemma, K is a minimum covering. \square

All of this tells us that there is a deep relationship between maximum matchings and minimum vertex covers on bipartite graphs. Given a solution to one, we can turn it into a solution to the other. This is what we seek to accomplish next.

2.2 Duality theory

Definition 2.4. Let

$$\begin{aligned} &\text{maximize } \mathbf{c}^T \mathbf{x} \\ &\text{subject to } A\mathbf{x} \leq \mathbf{b} \\ &\mathbf{x} \geq 0 \end{aligned}$$

be our linear program, which we will call the *primal* linear program. Then we define the *dual* of this linear program to be the linear program

$$\begin{aligned} &\text{minimize } \mathbf{b}^T \mathbf{y} \\ &\text{subject to } A^T \mathbf{y} \leq \mathbf{c} \\ &\mathbf{y} \geq 0 \end{aligned}$$

The first thing to note is that the dual of the dual is the primal. Let us introduce some notation. First, let us denote the primal maximization problem by the letter Γ , and the dual minimization problem by the letter Ω . For a given linear program, we denote an optimal solution by **OPT**.

Theorem 2.5. (*Weak duality*) *If the primal linear program (in maximization form) and the dual (in minimization form) are both feasible, then*

$$\mathbf{OPT}(\Gamma) \leq \mathbf{OPT}(\Omega).$$

What's surprising is the following theorem.

Theorem 2.6. (*Strong duality*) *Given two linear programs Γ and Ω that are duals of each other,*

if one is feasible and bounded, then so is the other. Additionally,

$$\mathbf{OPT}(\Gamma) = \mathbf{OPT}(\Omega).$$

Our goal now is to use this duality theory to better understand the relationships between our combinatorial optimization problems.

2.2.1 Maximum-matching duality

Let's first turn maximum matching into a linear program. Our goal is to maximize the number of edges in our matching. Our constraint is that no edge is incident to more than one edge in the matching. So for each edge (u, v) , we will need a corresponding x_{uv} . Our objective function is then pretty simple: maximize the number of x_{uv} . Now we need to figure out our constraint equations. For a fixed node $u \in U$, the number of edges in the matching incident to u is given by $\sum_{v \in V} x_{uv}$. So we want that this is ≤ 1 . Similarly, for any node v , we want $\sum_{u \in U} x_{uv} \leq 1$. This gives us the following linear program.

$$\text{maximize } \sum_{u,v} x_{uv} \tag{2.3}$$

$$\text{subject to } \sum_v x_{uv} \leq 1, \quad \forall u \in U, \tag{2.4}$$

$$\sum_u x_{uv} \leq 1, \quad \forall v \in V, \tag{2.5}$$

$$x_{uv} \in \{0, 1\}. \tag{2.6}$$

What we've given here is an *integer* linear program, since we've restricted our x variables to be integers. In general, solving integer linear programs is NP-hard. However, in this case it is well known that this linear program attains integer solution at extreme of the polyhedron solution space, so we can drop the integrality requirements and just say $x_{uv} \geq 0$.

Now let's try and construct the dual of this linear program. We will need a variable y_u for each vertex u . Similarly, we need a variable y_v for each vertex v . Our objective will be to minimize over the sum of these y_u, y_v . Since our constraint in the primal is the constant vector 1, our only constraint will be that $y_u + y_v \geq 1$. This gives us the dual linear program

$$\text{minimize} \quad \sum_{u,v} (y_u + y_v) \quad (2.7)$$

$$\text{subject to} \quad y_u + y_v \geq 1 \quad \forall u, v, \quad (2.8)$$

$$y_u, y_v \geq 0. \quad (2.9)$$

This dual problem tells us that each edge must be “covered” by at least one of its incident vertices. This is exactly the vertex cover problem! So for unweighted bipartite graphs, the linear programs for maximum matchings and minimum vertex covers are duals of each other. We will use this insight to construct our algorithms for solving the maximum matching problem.

There is another version of this problem, called the *maximum weight matching*. In this version, we are given a bipartite graph with non-negative edge weights w_{uv} for all edges (u, v) . Instead of trying to maximize the number of edges in the matching, the goal is to find a matching M such that $\sum_{(u,v) \in M} w_{uv}$, or the weight of the matching, is greater than the weight of any other matching. We can easily encode this problem by making a slight

modification to our linear program from before. The primal is given by

$$\text{maximize} \quad \sum_{u,v} w_{uv} x_{uv} \quad (2.10)$$

$$\text{subject to} \quad \sum_v x_{uv} \leq 1, \quad \forall u \in U, \quad (2.11)$$

$$\sum_u x_{uv} \leq 1, \quad \forall v \in V, \quad (2.12)$$

$$x_{uv} \geq 0. \quad (2.13)$$

Then the dual is

$$\text{minimize} \quad \sum_u y_u + \sum_v y_v \quad (2.14)$$

$$\text{subject to} \quad y_u + y_v \geq w_{uv} \quad \forall u, v, \quad (2.15)$$

$$y_u, y_v \geq 0. \quad (2.16)$$

The dual is a sort of weighted vertex cover. One way to think about it is that each edge has a “cost” given by w_{uv} , and each of its endpoints has to pool “money” in order to pay at least that cost. So instead of edges being covered or not covered, there’s a certain “amount” that they have to be covered.

2.2.2 Revisiting a familiar problem - max-flow/min-cut

Here we demonstrate that a common problem discussed in algorithms courses, and a quite amazing theorem, is really a special case of what we’ve been discussing. Recall that a flow network is defined as follows.

Definition 2.7. A *flow network* $G = (V, E)$ is a directed graph in which each edge $(u, v) \in E$ has nonnegative *capacity* $c_{uv} \geq 0$. Furthermore, there are two vertices, a source s and a sink t . We assume that every $v \in V$ lies in some path $s \rightarrow \cdots \rightarrow v \rightarrow \cdots \rightarrow t$.

Definition 2.8. Let $G = (V, E)$ be a flow network. A *flow* in G is a function $f : V \times V \rightarrow \mathbb{R}$ that satisfies the following:

- Capacity constraint: For all $u, v \in V$, $0 \leq f(u, v) \leq c_{uv}$.
- Conservation: For all $u \in V \setminus \{s, t\}$, we have

$$\sum_{v \in V} f(v, u) = \sum_{v \in V} f(u, v).$$

This says that for all vertices v except our source and sink, the flow out of v is equal to the flow into v .

We define the value of the flow $val(f) = \sum_{v \in V} f(s, v) - \sum_{v \in V} f(v, s)$, which is just the flow out of our source minus the flow into our source.

The problem as demonstrated in a typical algorithms course is to find a maximum flow from s to t . This is done using a method developed by Ford and Fulkerson which does what *ALG1* does, but in a residual graph (essentially the subgraph where the flows $f(u, v) < c_{uv}$, we refer the reader to [CITE CLRS] for more detail).

Now, we recall the definition of a cut in a flow network.

Definition 2.9. A *cut* (S, T) of a flow network $G = (V, E)$ is a partition of V into sets S and $T = V \setminus S$ such that $s \in S$ and $t \in T$. Given a flow f , the *net flow* $f(S, T)$ across the cut (S, T) is defined as

$$f(S, T) = \sum_{u \in S} \sum_{v \in T} f(u, v) - \sum_{u \in S} \sum_{v \in T} f(v, u).$$

Finally, the *capacity* of the cut is $c(S, T) = \sum_{u \in S} \sum_{v \in T} c(u, v)$. A minimum cut has capacity less than or equal to all other cuts in the network/

One of the coolest and most surprising theorems in an algorithms course is that, given a flow network G , the value of the maximum flow ($val(f)$) is equal to the capacity of the minimum $s - t$ cut of G .

Our goal now is to build up to this same theorem using the tools developed in this chapter. What follows is due to [CITE VAZIRANI]. We will first give a linear program for the maximum flow problem. To make things simpler, let's introduce an arc of infinite capacity from the sink t to the source s ; this converts this to a circulation, with the objective to maximize the flow $f(t, s)$. This allows us to enforce flow conservation at s and t as well, which makes the corresponding linear program simpler. The linear program is as follows.

$$\begin{aligned}
& \text{maximize} && f(t, s) \\
& \text{subject to} && f(u, v) \leq c_{uv}, && (u, v) \in E \\
& && \sum_{v: (v, u) \in E} f(v, u) \leq \sum_{v: (u, v) \in E} f(u, v), && u \in V \\
& && f(u, v) \geq 0 && (u, v) \in E.
\end{aligned}$$

It is not immediately obvious why the second set of inequalities implies flow conservation; all it seems to say is that for each u , the total flow into u is at most the total flow out of u . However, note that if this holds for all u , we in fact have equality of incoming and outgoing flow, since a deficit of flow at some u implies a flow surplus at some v . So this does in fact give us conservation of flow. Now we want to find the dual of this program. Our sense (hopefully) is that the dual will somehow relate to minimum cuts, given the foreshadowing of the previous section. Let's see! We introduce variables d_{uv} and p_u for each type of inequality in the primal.

$$\begin{aligned}
& \text{minimize} && \sum_{(u, v) \in E} c_{uv} d_{uv} \\
& \text{subject to} && d_{uv} - p_u + p_v \geq 0, && (u, v) \in E, \\
& && p_s - p_t \geq 1, \\
& && d_{uv} \geq 0, && (u, v) \in E.
\end{aligned}$$

It is known that extreme point solutions to these linear programs takes on values 0 or 1 at each coordinate. Let consider an optimal dual solution $(\mathbf{d}^*, \mathbf{p}^*)$. First, in order to satisfy $p_s^* - p_t^* \geq 1$ with 0,1 values, it must be the case that $p_s^* = 1$ and $p_t^* = 0$. This motivates an $s - t$ cut (S, T) with S consisting of nodes with value 1, and T the nodes with value zero. For an edge (u, v) such that $u \in S$ and $v \in T$, we have that $p_u^* = 1$ and $p_v^* = 0$, so by the first constraint $d_{uv}^* = 1$. This means that for any edge (u, v) in the cut, the corresponding $d_{uv}^* = 1$. Note that any other d_{uv}^* where (u, v) is not in the cut, it's value can be 0 without violating the constraints (we want them to be 0 since we are minimizing our objective function). Thus the objective function's value is equal to the capacity of this (S, T) cut, which must be a minimum cut. Strong duality tells us that this corresponds to the value of $f(t, s)$ in the primal linear program. To make this clear, we give an example of a flow network.

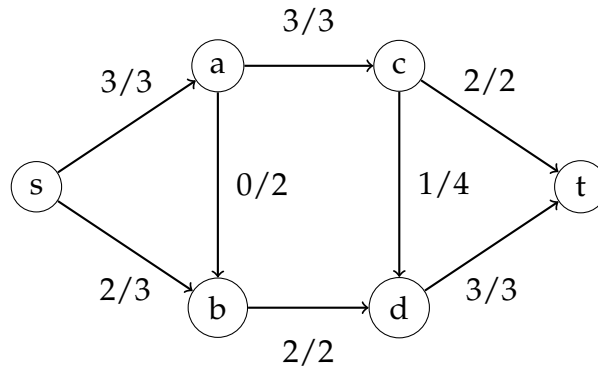


Figure 2.2: Flow network with maximum flow of 5

In Figure 2.2 we have a flow network that has a maximum flow from s to t of 5. We should expect that, given our discussion of the dual, we should be able to find a cut (S, T) of value 5 such that S consists of nodes with value 1, and T consists of nodes with value 0. We know that we must set $p_s = 1$ and $p_t = 0$. This gives us the following dual solution, where grey nodes are in S and white nodes are in T .

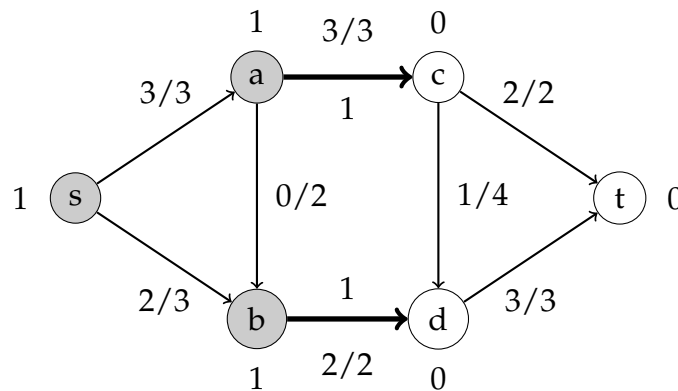


Figure 2.3: Flow network with corresponding dual solution

In Figure 2.3, the (S, T) cut corresponds to the sets $S = \{s, a, b\}$ and $T = \{c, d, t\}$; the edges in the cut are the bold edges $a \rightarrow c$ and $b \rightarrow d$. So the capacity of the cut is 5, which is what we'd expect. In order to make the figure readable, we leave out our dual labeling on edges that are not in the cut (their label is 0 anyways). The label on edges in the cut is 1, which we leave in. Note that our dual labeling of edges and vertices in this flow network satisfies our dual constraints.

Thus, we have given an alternative formulation of the max-flow/min-cut relationship using linear programs. In this next chapter we present the main algorithm of this thesis, the Hungarian algorithm, which is interesting in itself, and also as a tool to motivate general primal-dual algorithms.

Chapter 3

The Hungarian algorithm

In this chapter we present a comprehensive look at the Hungarian algorithm. The algorithm relies on some interesting techniques, which we spell out carefully. Ultimately, we show how this algorithm solves its associated primal and dual linear programs – maximum weight matching and minimum weight vertex cover, respectively.

3.1 Preliminaries - alternating trees, labelings, and the Kuhn-Munkres Theorem

In this section we use the motivation of the linear programs to develop an algorithm for simultaneously solving the maximum weight matching and minimum weight vertex cover problems. Our algorithm works by taking every exposed node on the left, and from each such node building a collection of alternating paths. We define this collection formally.

Definition 3.1. Let $G = (U, V, E)$ be a bipartite graph, and M a matching on G . An *alternating tree* with respect to M is a tree which satisfies two conditions:

- the tree contains exactly one node $u \in U$. We call u the *root* of the tree.

- all paths between the root and any other node in the tree are alternating paths.

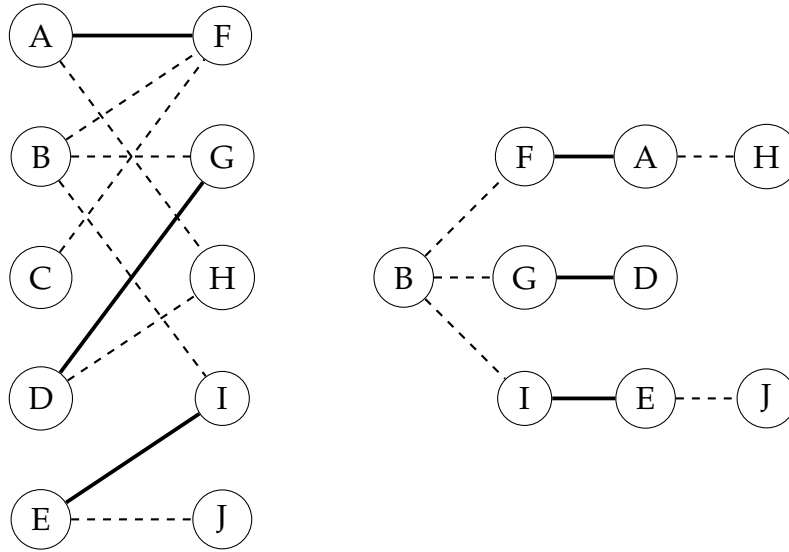


Figure 3.1: Bipartite graph with matching, corresponding alternating tree rooted at vertex B.

Let's remind ourselves what the primal-dual linear programs motivate. We want to minimize $\sum w_{uv}$, and maximize $\sum (y_u + y_v)$. Moreover, we want optimal solutions such that $\sum w_{uv} = \sum (y_u + y_v)$. For our primal, we are keeping track of edge weights. For the dual, we will be keeping track of a “labeling” on vertices, given by the y_u and y_v values. We define what a labeling is now.

Definition 3.2. A *vertex labeling* on a weighted bipartite graph $G = (U, V, E)$ is a function $l : U \cup V \rightarrow \mathbb{N}$. We call the labeling *feasible* if for all $u \in U$ and $v \in V$, $l(u) + l(v) \geq w_{uv}$.

This labeling corresponds to our dual variables; i.e. a feasible labeling is a feasible dual solution. It will be helpful us to look at a certain subset of our graph where the labeling is exact ($l(u) + l(v) - w_{uv} = 0$).

Definition 3.3. The *equality subgraph* of $G = (U, V, E)$ is the graph $G_l = (U, V, E_l)$, where

$$E_l = \{(u, v) : l(u) + l(v) = w_{uv}\}.$$

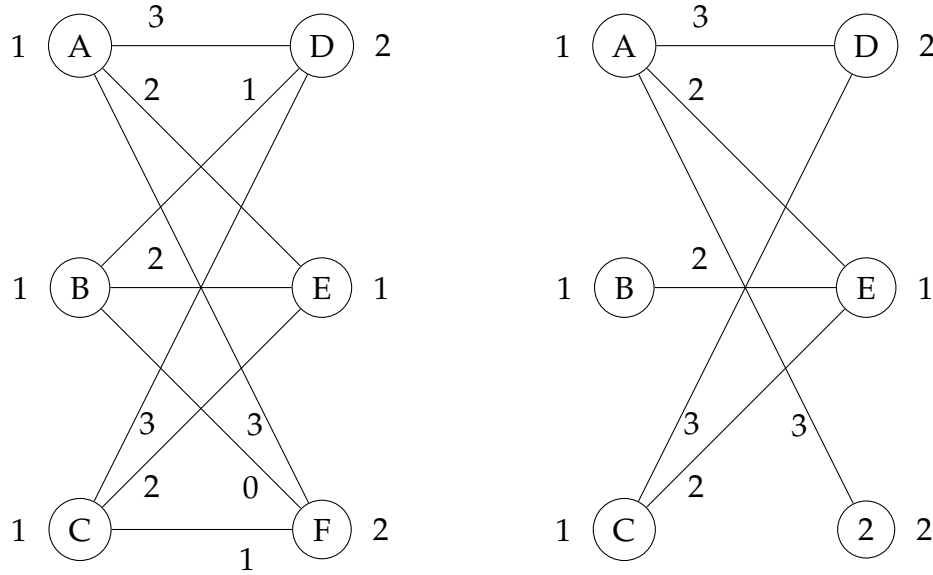


Figure 3.2: A weighted bipartite graph and its corresponding equality subgraph.

In Figure 2 we show a bipartite graph along with its corresponding equality graph.

Theorem 3.4. (Kuhn-Munkres) *If l is a feasible labeling and M is a perfect matching in G_l then M is a max-weight matching.*

Proof. Let M' be a perfect matching in G . Since every $u \in U \cup V$ is matched by exactly one edge in M' , then

$$\sum_{(u,v) \in M'} w_{uv} \leq \sum_{(u,v) \in M'} (l(u) + l(v)) = \sum_{w \in U \cup V} l(w).$$

This says that the sum of our label values is an upper bound on the weight of any perfect matching. Now suppose that M is a perfect matching in G_l . Then

$$\sum_{(u,v) \in M} w_{uv} = \sum_{w \in U \cup V} l(w).$$

So $\sum_{(u,v) \in M'} w_{uv} \leq \sum_{(u,v) \in M} w_{uv}$, meaning M must be maximum weight. \square

Lemma 3.5. Let $S \subseteq U$ and $T = N_l(S) \neq V$. Set

$$\alpha_l = \min_{u \in S, v \notin T} \{l(u) + l(v) - w_{uv}\}$$

and

$$l'(w) = \begin{cases} l(w) - \alpha_l & \text{if } w \in S \\ l(w) + \alpha_l & \text{if } w \in T \\ l(w) & \text{otherwise.} \end{cases}$$

Then l' is a feasible labeling, and

1. If $(u, v) \in E_l$ for $u \in S$ and $v \in T$ then $(u, v) \in E_{l'}$.
2. If $(u, v) \in E_l$ for $u \notin S$, and $v \notin T$, then $(u, v) \in E_{l'}$.
3. There is some edge $(u, v) \in E_{l'}$ for $u \in S, v \notin T$.

Proof. First, we show that l' is feasible. For $u \in U, v \in V$, there are four possibilities:

- if $u \in S$ and $v \in T$ then $l'(u) + l'(v) = l(u) + l(v)$.
- if $u \notin S$ and $v \notin T$, $l'(u) = l(u)$ and $l'(v) = l(v)$, so $l'(u) + l'(v) = l(u) + l(v)$.
- if $u \in S$ and $v \notin T$, $l'(u) = l(u) - \alpha$ and $l'(v) = l(v)$. We know $\alpha = \min_{u \in S, v \notin T} \{l(u) + l(v) - w_{uv}\}$, which means $\alpha \leq l(u) + l(v) - w_{uv}$, and thus $l'(u) + l'(v) \geq w_{uv}$.
- if $u \notin S$ and $v \in T$, $l'(u) = l(u)$ and $l'(v) = l(v) + \alpha$, which is clearly feasible.

(1) and (2) follow from above. To see (3), note that there is some (u, v) with $u \in S, v \in T$ such that $\alpha = l(u) + l(v) - w_{uv}$, so when we take $l'(u) = l(u) - \alpha$ and $l'(v) = l(v)$, we

get

$$\begin{aligned}
 l'(u) + l'(v) - w_{uv} &= l(u) - \alpha + l(v) - w_{uv} \\
 &= l(u) + l(v) - w_{uv} - \alpha \\
 &= \alpha - \alpha \\
 &= 0.
 \end{aligned}$$

This is exactly what it means for an edge (u, v) to be in E_l . □

We now look at the Hungarian method for finding maximum-weight matchings on bipartite graphs. This method was originally developed by Kuhn and Munkres, who named it in honor of the Hungarian mathematicians Kőnig and Egervary.

3.2 The Hungarian Algorithm

We now look at the Hungarian method for finding maximum-weight matchings on bipartite graphs. This method was originally developed by Kuhn and Munkres, who named it in honor of the Hungarian mathematicians Kőnig and Egervary.

The Hungarian Method

1. Choose initial feasible labeling l and matching M in G_l .
2. If M is perfect in G_l , we are done. Otherwise, pick exposed vertex $u \in U$. Set $S = \{u\}$, $T = \emptyset$.
3. If $N_l(S) = T$, update labels as in lemma (this forces $N_l(S) \neq T$).
4. If $N_l(S) \neq T$, pick $v \in N_l(S) \setminus T$
 - If v is exposed, $p = u \rightarrow v$ is an augmenting path. Set $M := M \oplus p$. Go to 2.

- If v is matched to some w , expand our alternating tree. $S := S \cup \{w\}$, $T := T \cup \{v\}$. Go to 3.

The correctness of the algorithm follows from the lemma and theorem. We now provide an example of this algorithm. Our initial matching is $M = \{(B, E), (C, D)\}$ (see Figure 3).

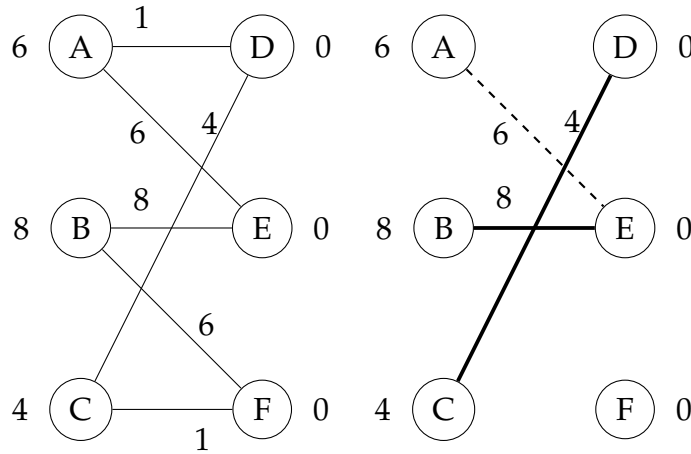


Figure 3.3: Bipartite graph (left) and corresponding equality graph (right) with initial matching

Note that the current state of the graph is primal-dual feasible. Our algorithm chooses an exposed vertex in U , say A . So we have $S = \{A\}$ and $T = \emptyset$. We have that $N_l(S) \neq T$, so we find $E \in N_l(S) \setminus T$. E is matched, so we grow our alternating tree as follows: $S := S \cup \{B\} = \{A, B\}$, $T := T \cup \{E\} = \{E\}$. At this point $N_l(S) = T$, so we adjust our dual variables. Calculate $\alpha = \min_{u \in S, v \notin T} \{l(u) + l(v) - w_{uv}\} = 2$ from edge (B, F) . Our new equality graph is shown in Figure 4.

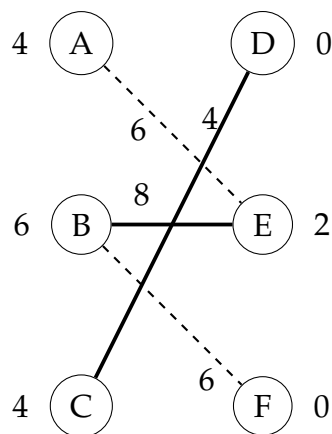


Figure 3.4: Second equality graph.

Now, $S = \{A, B\}$ is the same, but $N_l(S) = \{E, F\}$ has changed. $T = \{E\}$, so $N_l(S) \neq T$. So we choose $F \in N_l(S) \setminus T$. F is unmatched, meaning it is an endpoint of an augmenting path. In particular, $p = A, E, B, F$ is an augmenting path. Thus we improve our matching with $M := M \oplus p = \{(A, E), (B, F), (C, D)\}$. Our equality graph with the new matching is given in Figure 5.

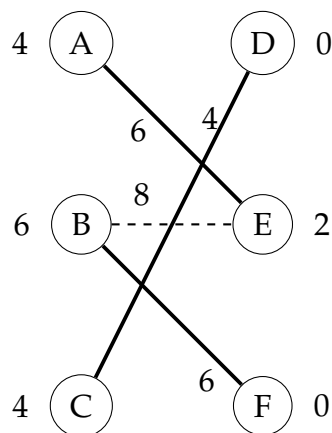


Figure 3.5: Equality graph after augmenting.

This is a perfect matching on the equality graph, so this matching must be a maximum weight matching on the graph. We can check that the values of the primal and dual solutions agree. The sum of weights in the matching is $6 + 6 + 4 = 16$, and the sum of the values of our dual variables is $4 + 6 + 4 + 2 = 16$.

Note that if we want to just find a maximum cardinality matching on a bipartite graph, we can just give all edges weight 1 and run this algorithm.

This algorithm was one of the first primal-dual algorithms developed, and it anticipated many later variations on the same theme. It displays the surprising connection between combinatorial optimization and linear programming, which we explore in the next section.

Chapter 4

General primal-dual method and auction algorithms

In this chapter we will take a step back and look at primal-dual algorithms in more generality. The goal will be to describe a method of solving a set of primal-dual algorithms for *network design problems*. Again, we will be restricting our attention to bipartite graphs. In network design problems we are given a graph $G = (U, V, E)$ and a cost c_{uv} for each edge $(u, v) \in E$, and the goal is to find a minimum/maximum-cost subset $E' \subset E$ that satisfies some criteria. Our maximum-cost matching problem is an example of this. There are other common examples that we will explore later on, but for now it suffices to just think of these problems as choosing subsets of our graph according to some stipulations. Throughout, we will be looking at undirected graphs.

4.1 The Classical Primal-Dual Method

We begin by looking at what's known as the “classical” primal-dual method, which is concerned with linear programs for polynomial-time solvable optimization problems. This will allow us to build up a framework for a more general primal-dual method that we

can use for approximation algorithms - i.e. for problems that are known to be *NP*-hard.

** (Not sure if this is within the scope of the thesis – discuss with Jim)

Let's consider the linear program

$$\text{minimize } \mathbf{c}^T \mathbf{x} \quad (4.1)$$

$$\text{subject to } A\mathbf{x} \geq \mathbf{b} \quad (4.2)$$

$$\mathbf{x} \geq 0 \quad (4.3)$$

and its dual

$$\text{maximize } \mathbf{b}^T \mathbf{y} \quad (4.4)$$

$$\text{subject to } A^T \mathbf{y} \leq \mathbf{c} \quad (4.5)$$

$$\mathbf{y} \geq 0. \quad (4.6)$$

We first define a concept that we will use throughout the rest of this thesis.

Definition 4.1. (Complementary slackness) Given two linear programs in the form above, the *primal complementary slackness conditions* are the conditions which, given primal solution \mathbf{x} , are necessary for a dual solution \mathbf{y} :

$$x_j > 0 \implies A^j \mathbf{y} = c_j,$$

where A^j is the j th column of A . Similarly, the *dual complementary slackness conditions* are the conditions which, given dual solution \mathbf{y} , are necessary for a primal solution \mathbf{x} :

$$y_i > 0 \implies A_i \mathbf{x} = b_i,$$

where A_i is the i th row of A . Together, these conditions give us necessary and sufficient conditions for solving the primal-dual system, which we will prove. The (maximization)

primal slackness variables are given by $\mathbf{s} = \mathbf{b} - A\mathbf{x}$. The dual slackness variables are given by $\mathbf{t} = A^T\mathbf{y} - \mathbf{c}$.

Theorem 4.2. *Let \mathbf{x} be a primal feasible solution, and \mathbf{y} a dual feasible solution. Let \mathbf{s} and \mathbf{t} be the corresponding slackness variables. Then \mathbf{x} and \mathbf{y} are optimal solutions if and only if the following two conditions hold:*

$$x_j t_j = 0 \quad \forall j \quad (4.7)$$

$$y_i s_i = 0 \quad \forall i. \quad (4.8)$$

Proof. Let $u_i = y_i s_i$ and $v_j = x_j t_j$, and $\mathbf{u} = \sum_i u_i$, $\mathbf{v} = \sum_j v_j$. Then $\mathbf{u} = 0$ and $\mathbf{v} = 0$ if and only if (7) and (8) hold. Also,

$$\begin{aligned} \mathbf{u} + \mathbf{v} &= \sum y_i s_i + \sum x_j t_j \\ &= \sum y_i (b_i - A_i x_i) + \sum x_j (A_j^T y_j - c_j) \\ &= \sum b_i y_i - \sum c_j x_j, \end{aligned}$$

so we get that $c^T \mathbf{x} = b^T \mathbf{y}$ if and only if $u + v = 0$, which proves the statement. \square

The general “tug-of-war” between the primal and dual suggests an economic interpretation of slackness conditions. We can think of our primal (maximization) problem as concerned with profit given some constraints on resources, i.e. a resource allocation problem. The dual can be interpreted as a valuation of the resources – it tells us the availability of a resource, and its price. So if we have optimal \mathbf{x} and \mathbf{y} , we can interpret slackness as follows: if there is slack in a constrained primal resource i ($s_u > 0$), then additional units of that resource must have no value ($y_u = 0$); if there is slack in the dual price constraint ($t_v > 0$) there must be a shortage of that resource ($x_v = 0$).

We now give an example of complementary slackness in action. Let’s look back to our maximum weight matching problem. Recall the primal linear program for maximum-

weight matching:

$$\text{maximize} \quad \sum_{u,v} c_{uv} x_{uv} \quad (4.9)$$

$$\text{subject to} \quad \sum_v x_{uv} \leq 1, \quad \forall u \in U, \quad (4.10)$$

$$\sum_u x_{uv} \leq 1, \quad \forall v \in V, \quad (4.11)$$

$$x_{uv} \geq 0. \quad (4.12)$$

and its dual

$$\text{minimize} \quad \sum_u y_u + \sum_v y_v \quad (4.13)$$

$$\text{subject to} \quad y_u + y_v \geq c_{uv} \quad \forall u \in U, v \in V, \quad (4.14)$$

$$y_u, y_v \geq 0. \quad (4.15)$$

The format here is a little different, since our primal is a maximization problem and the dual is a minimization, but it's easy enough to reverse the roles. It's easy to see our corresponding primal complementary slackness conditions are

$$x_{uv} > 0 \implies y_u + y_v = c_{uv}. \quad (4.16)$$

The dual complementary slackness conditions are

$$y_u > 0 \implies \sum_v x_{uv} = 1, \quad (4.17)$$

$$y_v > 0 \implies \sum_u x_{uv} = 1. \quad (4.18)$$

In general, the slackness conditions guide us in our algorithm – they tell us how, given a solution to one of the problems, we should augment the solution to the other. For

example, the algorithm we presented for maximum-weight matching/minimum vertex cover initializes with a solution to both the primal and dual that satisfies conditions (8) and (10); the algorithm then at each step works to decrease the number of conditions in (9) that are unsatisfied, while maintaining satisfiability of (8) and (10). This method is not unique to the Hungarian algorithm. In fact, the Hungarian algorithm paved the way for this general method, which we describe presently.

Looking back at the original linear programs at the beginning of this chapter, suppose we have a dual feasible solution \mathbf{y} . We can then state the problem of finding a feasible primal solution \mathbf{x} that obeys our complementary slackness conditions as another *restricted* linear program. Define the sets $A = \{j \mid A^j \mathbf{y} = c_j\}$ and $B = \{i \mid y_i = 0\}$. So A tells us which dual constraints (5) are tight, given the solution \mathbf{y} , and B tells us which y_i are 0. What we want to do is give a linear program to find a solution \mathbf{x} that minimizes the “violation” of the complementary slackness conditions and the primal constraints, and to do so we will index our variables by these sets. We will have slack variables s_i which will describe the difference between $A_i \mathbf{x}$ and b_i for $i \notin A$. We do this because we want to look at all $y_i > 0$ where we do not have that $A_i \mathbf{x} = b_i$. So part of our objective function will be to minimize the sum of these s_i . We also want to minimize the sum over variables x_j where $j \notin A$. This is because we want to see if there are any x_j such that $A^j \mathbf{y} \neq c_j$. So we give the following restricted primal linear program:

$$\text{minimize } \sum_{i \notin B} s_i + \sum_{j \notin A} x_j \quad (4.19)$$

$$\text{subject to } A_i \mathbf{x} \geq b_i \quad i \in B, \quad (4.20)$$

$$A_i \mathbf{x} - b_i = s_i \quad i \notin B, \quad (4.21)$$

$$\mathbf{x} \geq 0, \quad (4.22)$$

$$\mathbf{s} \geq 0. \quad (4.23)$$

Observe that if this restricted primal has a feasible solution (\mathbf{x}, \mathbf{s}) such that the objective

function is 0, then \mathbf{x} is a feasible primal solution that satisfies the complementary slackness conditions for the dual solution \mathbf{y} . This means that \mathbf{x} and \mathbf{y} are optimal primal and dual solutions. If, however, the optimal solution to this restricted primal has value greater than 0, more work is required. We can consider the dual of the restricted primal:

$$\text{maximize } \mathbf{b}^T \mathbf{w} \quad (4.24)$$

$$\text{subject to } A^j \mathbf{w} \leq 0 \quad j \in A, \quad (4.25)$$

$$A^j \mathbf{w} \leq 1 \quad j \notin A, \quad (4.26)$$

$$w_i' \geq -1 \quad i \notin B, \quad (4.27)$$

$$w_i' \geq 0 \quad i \in B. \quad (4.28)$$

What we want here is to improve our dual solution. By assumption, the optimal solution to this linear program's primal is greater than 0, so we know that this dual has a solution \mathbf{w} such that $\mathbf{b}^T \mathbf{w} > 0$. What we want is the existence of some $\epsilon > 0$ such that $\mathbf{y}' = \mathbf{y} + \epsilon \mathbf{w}$ is a feasible dual solution. In particular, a solution of this form will be an improvement on our original solution \mathbf{y} . We can calculate bounds on ϵ as follows. The two conditions we must satisfy in order to maintain dual feasibility are that $y_i' \geq 0$ and $A_j^T y' \leq c_j$. This means that we need

$$y_i + \epsilon w_i \geq 0 \quad (4.29)$$

$$A_j^T \mathbf{y} + A_j^T \epsilon \mathbf{w} \leq c_j. \quad (4.30)$$

Let's consider the first one. When $w_i > 0$, we are fine; we need to be careful when $w_i < 0$ since this could potentially violate the inequality. Solving in this way, we get a first bound on ϵ :

$$\epsilon \leq \min_{i \in B: w_i < 0} (-y_i / w_i).$$

Now let's address the second inequality. When $A_j^T \mathbf{w} \leq 0$, we are definitely okay. We need

to be careful about violating the constraint when $A^T w > 0$. Thus, we can calculate a second bound on ϵ :

$$\epsilon \leq \min_{j \in A: A_j^T w > 0} \frac{c_j - A_j^T y}{A_j^T w}.$$

If we choose the lower of these two ϵ values, we obtain a new feasible dual solution that has greater objective value. We can then work by reiterating the procedure, with the hope that we find an optimal primal solution.

It's not immediately clear why reducing our original linear programs to a series of linear programs is helpful. However, note that the vector \mathbf{c} has totally disappeared in the restricted primal and its dual. Recall that in the original linear program, \mathbf{c} gave us the edge-costs on our graph. So this method reworks our original weighted problem into unweighted parts, which are easier to solve. Oftentimes, it is the case that we can interpret these unweighted problems as purely combinatorial problems, which means that instead of actually solving the problem with linear programming, we can solve it by combinatorial methods. Using a combinatorial algorithm to find a solution \mathbf{x} that obeys the complementary slackness conditions, or to find an improved dual solution \mathbf{y} , is oftentimes more efficient.

4.2 Primal-dual method for weighted matchings

Let us now look at an example of this method. We will look at a weighted matching problem, as in the previous chapter, but this time we will look at *minimizing* the cost of the matching, instead of maximizing. We do this mainly because it illustrates something important about the underlying structure of these matching problems. It will be easy to see how the same method can be used for the case in which we want a maximum matching. So the primal linear program for a minimum weight perfect matching on a

bipartite graph is given as follows.

$$\text{minimize} \quad \sum_{u,v} c_{uv} x_{uv} \quad (4.31)$$

$$\text{subject to} \quad \sum_v x_{uv} \geq 1, \quad \forall u \in U, \quad (4.32)$$

$$\sum_u x_{uv} \geq 1, \quad \forall v \in V, \quad (4.33)$$

$$x_{uv} \geq 0. \quad (4.34)$$

Its dual is

$$\text{maximize} \quad \sum_u y_u + \sum_v y_v \quad (4.35)$$

$$\text{subject to} \quad y_u + y_v \leq c_{uv} \quad \forall u \in U, v \in V, \quad (4.36)$$

$$y_u, y_v \geq 0. \quad (4.37)$$

We need to start with a dual feasible solution, and try to find a primal solution that minimizes the violation of the constraints and slackness conditions. We can start with the trivial dual solution of $y_u, y_v = 0$ for all u, v . Let's now think about our primal complementary slackness. The set A is given by $\{(u, v) \in E : y_u + y_v = c_{uv}\}$. We know that these are the edges we want to include in our matching, and since we know our linear program has integer solutions at extreme points of the polyhedron, let's specify that $x_{uv} = 0$ for $(u, v) \notin A$. Now, our other slackness variables s_u, s_v look like

$$\sum_{v:(u,v) \in E} x_{uv} - s_u = 1$$

$$\sum_{u:(u,v) \in E} x_{uv} - s_v = 1.$$

So we want to minimize over the sum of s_u and s_v . Note that at this point, our set B consists of all vertices. So our restricted primal linear program is

$$\text{minimize } \sum_{u \in U} s_u + \sum_{v \in V} s_v \quad (4.38)$$

$$\text{subject to } \sum_v x_{uv} - s_u = 1 \quad \forall u, \quad (4.39)$$

$$\sum_u x_{uv} - s_v = 1 \quad \forall v, \quad (4.40)$$

$$x_{uv} = 0 \quad (u, v) \notin A, \quad (4.41)$$

$$x_{uv} \geq 0 \quad (u, v) \in A, \quad (4.42)$$

$$\mathbf{s} \geq 0. \quad (4.43)$$

Let's first observe that all components of this restricted primal take on values 0 or 1, as in the original primal. Moreover, note that we have turned a weighted problem into an unweighted combinatorial problem. We've specified that we are not including any edge $(u, v) \notin A$ in our matching, and we are trying to include as many $(u, v) \in A$ as possible in our matching by minimizing the slackness variables. Note that the graph $G' = (U, V, J)$ is exactly the equality subgraph as defined in the previous chapter! In our Hungarian algorithm we repeatedly sought to find maximum cardinality matchings within this subgraph, which is exactly what this restricted primal is having us do. This tells us that the problems of maximum weight matching and minimum weight matching only differ in the labeling we are specifying. The underlying procedure for solving both of the problems is essentially the same. So if we find a perfect matching in G' , we will have found an \mathbf{x} that obeys the complementary slackness conditions, i.e. $\sum_u s_u + \sum_v s_v = 0$. Moreover, this implies that the dual solution $\sum_u y_u + \sum_v y_v$ must be optimal as well. Now, if the solution $\sum_u s_u + \sum_v s_v > 0$, we do not have an optimal \mathbf{x} , so we need to adjust

our dual. We look at this now, in the dual linear program of the restricted primal.

$$\text{maximize } \sum_{u \in U} w_u + \sum_{v \in V} w_v \quad (4.44)$$

$$\text{subject to } w_u + w_v \leq 0 \quad \forall (u, v) \in A, \quad (4.45)$$

$$w_u + w_v \leq 1 \quad \forall (u, v) \notin A, \quad (4.46)$$

$$w_u, w_v \geq -1 \quad u, v \notin B, \quad (4.47)$$

$$w_u, w_v \geq 0 \quad u, v \in B, \quad (4.48)$$

$$\mathbf{w} \geq 0. \quad (4.49)$$

We now want to find an ϵ such that the solution $z = \sum_u y_u + \sum_v y_v + \epsilon(\sum_u w_u + \sum_v w_v)$ is (1) feasible and (2) an improvement of the dual objective. First of all, we know that since the restricted primal has solution ≥ 0 , the solution to this dual will also be ≥ 0 . So we just need to worry about the condition

$$y_u + y_v + \epsilon(w_u + w_v) \leq c_{uv}.$$

So we get that we at least need that $\epsilon \leq \min_{(u,v) \notin A: w_u + w_v > 0} \frac{c_{uv} - y_u - y_v}{w_u + w_v}$. We can refine this by noting that since $0 < w_u + w_v \leq 1$ for $(u, v) \notin A$, we have $\epsilon = \min_{(u,v) \notin A} (c_{uv} - y_u - y_v)$. Note that the negative of this is exactly the quantity we modify our labeling by in the Hungarian algorithm in the previous chapter. Thus we've found an ϵ that maintains dual feasibility, and increases the objective function. We can use this solution and revisit the restricted primal in order to look for an improved feasible primal solution.

4.3 Auction algorithms

We now look at a cool application of weighted bipartite matching. Let's imagine our assignment problem is an auction. We will motivate this through an economic lens. Sup-

pose that a good g has a price p_g , and a buyer b must pay p_g to receive this item. Suppose that the amount b is willing to pay for g , or b 's valuation of g , is c_{bg} . Then the net value of this item to b is $c_{bg} - p_g$. Let's also adopt the notation that $A(b)$ is the set of all items bidder b is interested in (willing to pay for). Assuming that each bidder b is a rational agent acting in their own best interest, each b would want to be assigned a good g_b that maximizes their net value, i.e. for all b with good g_b , we have that

$$c_{bg_b} - p_{g_b} = \max_g \{c_{bg} - p_g\}. \quad (4.50)$$

This would give us an economic system in equilibrium; no bidder would have an incentive to seek/trade for a different good. The reason this is of interest to us is that an equilibrium assignment maximizes total profit (i.e. it solves the primal linear program for maximum weight matchings), while the corresponding set of prices solves the associated dual problem.

Our goal is to maximize the total amount earned in the auction – i.e. to maximize $\sum_{(b,g)} c_{bg}$, under the constraint that no bidder gets more than one good, and no good is purchased by more than one bidder.

It will be useful for us to consider an alternate, but equivalent dual problem, which is suggested by our complementary slackness conditions. In the auction setting, it does not make sense to have a labeling on the vertices in B , the set of buyers, whereas in the usual dual we have variables on both sets of vertices that partition our graph. The usual dual would look like

$$\text{minimize } \sum_b p_b + \sum_g p_g \quad (4.51)$$

$$\text{subject to } p_b + p_g \geq c_{bg} \quad (4.52)$$

$$p_b, p_g \geq 0 \quad u, v \in B. \quad (4.53)$$

In the auction setting, only values of p_g make sense; “prices” on the bidders is meaningless (unless our auction is two-way, which we are not considering). Note that $\sum_b p_b$ is minimized when we set each p_b to the smallest value allowed by the constraints (4.52), i.e. $p_b := \min_g \{c_{bg} - p_g\}$. This gives us the following equivalent dual objective function

$$\text{minimize } \sum_b \min_g \{c_{bg} - p_g\} + \sum_g p_g, \quad (4.54)$$

which is unconstrained. Note that this is an exact translation of our complementary slackness conditions (4.50). Again, this illustrates the importance of complementary slackness in the creation of algorithms for combinatorial optimization.

4.3.1 Naive algorithm

We will first proceed in describing a naive auction algorithm, which although clever, has cases in which it loops infinitely. Nonetheless, we will use it as motivation for our final auction algorithm.

Throughout the algorithm, we will keep track of two sets of bidders – those that are currently assigned a good g , and those that are not. At a general step in our algorithm, we will want to consider a bidder b that is currently not the owner of any good, and find a good g_b such that $c_{bg_b} - p_{g_b} > \max_{g \in A(b)} \{c_{bg} - p_g\}$. We will then need to do two things: (1), replace the current owner b' of g_b (put them back into the set of unassigned bidders) with b ; (2) bump the price of g_b by some amount δ_b . Intuitively, we want to increase p_{g_b} by the largest amount while still maintaining the fact that g_b offers maximal value to b . We can easily accomplish this by setting $\delta_b = v_b - w_b$, where $v_b = \max_{g \in A(b)} \{c_{bg} - p_g\} = p_{g_b}$ and $w_b = \max_{g \in A(b), g \neq g_b} \{c_{bg} - p_g\}$. This simply says that we are bumping the price p_b by the difference in the profit of the best object and the profit of the second best object.

This is a simple algorithm that has one major problem: there are cases in which it does

not terminate. The issue is that the bidding increment δ_b is zero when more than one object offers maximum value to a bidder b . Consequently, a situation may arise in which a number of bidders contend for a smaller number of equally desirable objects, creating an infinite loop. An example of this is shown in Figure 4.1. In Figure 4.1 the initial prices on

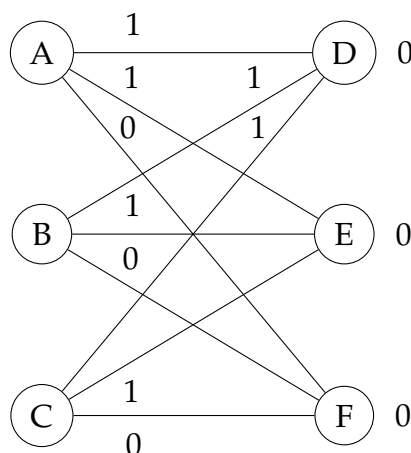


Figure 4.1: Example where the naive algorithm does not terminate.

all goods is zero. All bidders desire objects D and E since they offer a benefit of $1 - 0 = 1$, while object F offers a benefit of zero to all bidders. Say we first assign bidder A to object D . Then the algorithm will cycle as bidders B and C compete for item E without its price being raised.

The issue with our naive algorithm can be remedied by adding a small perturbation term to our complementary slackness, which will break the cycles that cause problems like in 4.1.

4.3.2 The auction algorithm

In real auctions each bid for an object must raise the object's price by some minimum positive increment ϵ . This will ensure that no matter what, whenever someone is assigned an item, the price raises by some non-zero amount. We will say that an assignment and a

price vector p satisfy ϵ -complementary slackness if

$$c_{b_g} - p_{g_b} \geq \max_{g \in A(b)} \{c_{bg} - p_g\} - \epsilon. \quad (4.55)$$

The new auction algorithm will be the same as the naive one, except in the way we adjust prices; now $\delta_b = v_b - w_b + \epsilon$. We give the pseudocode below. We will use a data structure called a queue (\mathcal{Q}) to keep track of unassigned bidders. A queue works exactly how one might think: new elements are placed on the “bottom”, and members are removed from the “top”.

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1  For each  $g \in G$ , set  $p_g := 0$ 
2  Queue  $\mathcal{Q} := B$ .
3  while  $\mathcal{Q} \neq \emptyset$ 
4       $b = \mathcal{Q}.deque()$ 
5      Find  $g_b \in A(b)$  that maximizes  $c_{bg} - p_g$ 
6      Let  $b'$  be the bidder currently assigned to  $g_b$  (null if none)
7      if  $c_{bg_b} - p_{g_b} \geq 0$ 
8           $\mathcal{Q}.enqueue(b')$ 
9          Assign  $g_b$  to  $b$ 
10      $p_g = p_g + \delta$ 
11 return Assignment and price vector

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Figure 4.2: Auction algorithm

Now, because we are relaxing our complementary slackness constraint, we should not expect to get optimal solutions. In fact, optimality of our solution relies entirely on how we choose our ϵ . It turns out that whatever ϵ is, the solution by a feasible assignment and price vector that satisfy ϵ -complementary slackness will be within $n\epsilon$ of optimal. We state this formally.

Theorem 4.3. *A feasible assignment that satisfies ϵ -complementary slackness along a price vector \mathbf{p} is within $n\epsilon$ of being optimal. Furthermore, the price vector is within $n\epsilon$ of being an optimal dual solution.*

Proof. Let Γ be an optimal assignment, and Ω the corresponding dual cost. So

$$\Gamma = \max_{g_b} \sum_b c_{bg_b} \quad \text{where } g_b \neq g_{b'} \text{ if } b \neq b' \quad (4.56)$$

$$\Omega = \min_{p_g} \left\{ \sum_b \max_{g \in A(b)} \{c_{bg} - p_g\} + \sum_g p_g \right\} \quad (4.57)$$

Firstly, weak duality tells us that $\Gamma \leq \Omega$. Since our assignment and prices satisfy ϵ -complementary slackness, we know that

$$\max_{g \in A(b)} \{c_{bg} - p_g\} - \epsilon \leq c_{bg_b} - p_{g_b}.$$

This means that we have

$$\Omega \leq \sum_b \left(\max_{g \in A(b)} \{c_{bg} - p_g\} + p_{g_b} \right) \leq \sum_b c_{bg_b} + n\epsilon \leq \Gamma + n\epsilon.$$

This shows that both our primal and dual objective functions are within $n\epsilon$ of optimality. □

Note that in the case where all c_{bg} are integers (which is usually the case in auctions), and we choose $\epsilon < 1/n$, then our algorithm outputs the optimal solution.

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